Fall 2014

Techniques for improving the scalability of data center networks

Advait Dixit
Purdue University

Graduate School ETD Form 9
(Revised 01/14)

PURDUE UNIVERSITY
GRADUATE SCHOOL
Thesis/Dissertation Acceptance

This is to certify that the thesis/dissertation prepared

By Advait Abhay Dixit

Entitled
Techniques for Improving the Scalability of Data Center Networks

For the degree of Doctor of Philosophy

Is approved by the final examining committee:

Ramana Rao Kompella	
Y. Charlie Hu	
Patrick Eugster	
Sonia Fahmy	

To the best of my knowledge and as understood by the student in the *Thesis/Dissertation Agreement. Publication Delay, and Certification/Disclaimer (Graduate School Form 32)*, this thesis/dissertation adheres to the provisions of Purdue University's "Policy on Integrity in Research" and the use of copyrighted material.

Approved by Major Professor(s): Y. Charlie Hu
Ramana Rao Kompella

Approved by: Sunil Prabhakar, William J. Gorman 11/04/2014

Head of the Department Graduate Program Date

TECHNIQUES FOR IMPROVING THE SCALABILITY OF DATA CENTER NETWORKS

A Dissertation

Submitted to the Faculty

of

Purdue University

by

Advait Abhay Dixit

In Partial Fulfillment of the

Requirements for the Degree

of

Doctor of Philosophy

December 2014

Purdue University

West Lafayette, Indiana

To my parents.

ACKNOWLEDGMENTS

I would like to thank my advisors Professor Ramana Rao Kompella and Professor Y. Charlie Hu for their guidance and support throughout my PhD. They helped me get started with my PhD, patiently guided my research and kept me motivated in difficult times. I will forever be indebted to them.

I would like express my gratitude to Dr. Fang Hao, Dr. Sarit Mukherjee and Dr. T. V. Lakshman, all from Bell Labs, for introducing me to software-defined networking. The project that started during my internship there evolved into ElastiCon which is incorporated in this dissertation.

Professor Patrick Eugster and Dr. Kirill Kogan have helped me immensely during my last year at Purdue University. They helped guide my research in a new direction and provided the basic idea behind composing SDN controllers, which I have included in this dissertation.

I am grateful to Dr. Nandita Dukkipati who mentored me during my internship at Google. The experience of working in a real data center environment has been of great help. I also want to thank Nipun Arora, my mentor during my internship at NEC Labs. The internship helped me understand the complexities involved in a deploying a software-defined network.

I also want to thank Dr. Rick Kenell and Dr. Jeff Turkstra for helping me get started when I first arrived at Purdue. My thanks also go to labmates Dr. Myungjin Lee, Dr. Pawan Prakash and Hitesh Khandelwal. The technical discussions and light-hearted conversations in the corridors of Lawson building made my time in the lab more productive and enjoyable.

I owe my greatest gratitude to my parents, Rita and Abhay Dixit, and my sister and brother-in-law, Ruhi and Harsha Joshi, for encouraging me to return to academia for a PhD and supporting me for the entire journey. Last, but not the least, I thank my wife Praveena Kunaparaju. Your love and companionship have given me strength to face the challenges of graduate student life.

TABLE OF CONTENTS

LIST OF TABLES

LIST OF FIGURES

ABBREVIATIONS

ACK	Acknowledgment
BGP	Border Gateway Protocol
CPU	Central Processing Unit
DCTCP	Data Center Transmission Control Protocol
DHT	Distributed Hash Table
DRM	Distributed Resource Management
DSACK	Duplicate Selective Acknowledgment
ECMP	Equal Cost Multipath Protocol
ECN	Explicit Congestion Notification
IP	Internet Protocol
ISP	Internet Service Provider
MP-TCP	Multi Path Transmission Control Protocol
NAT	Network Address Translation
NetFPGA	Network Field Programmable Gateway Array
OVS	Open vSwitch
RED	Random Early Discard
RTT	Round Trip Time
RPS	Random Packet Spraying
SDN	Software-defined Networking
TCAM	Ternary Content Addressable Memory
TCP	Transmission Control Protocol
ToR	Top-of-rack
VLB	Valiant Load Balancing
VM	Virtual Machine

ABSTRACT

Dixit, Advait Abhay Ph.D., Purdue University, December 2014. Techniques for Improving the Scalability of Data Center Networks. Major Professors: Ramana Rao Kompella and Y. Charlie Hu.

Data centers require highly scalable data and control planes for ensuring good performance of distributed applications. Along the data plane, network throughput and latency directly impact application performance metrics. This has led researchers to propose high bisection bandwidth network topologies based on multi-rooted trees for data center networks. However, such topologies require efficient traffic splitting algorithms to fully utilize all available bandwidth. Along the control plane, the centralized controller for software-defined networks presents new scalability challenges. The logically centralized controller needs to scale according to network demands. Also, since all services are implemented in the centralized controller, it should allow easy integration of different types of network services.

In this dissertation, we propose techniques to address scalability challenges along the data and control planes of data center networks.

Along the data plane, we propose a fine-grained traffic splitting technique for data center networks organized as multi-rooted trees. Splitting individual flows can provide better load balance but is not preferred because of potential packet reordering that conventional wisdom suggests may negatively interact with TCP congestion control. We demonstrate that, due to symmetry of the network topology, TCP is able to tolerate the induced packet reordering and maintain a single estimate of RTT.

Along the control plane, we design a scalable distributed SDN control plane architecture. We propose algorithms to evenly distribute the load among the controller nodes of the control plane. The algorithms evenly distribute the load by dynamically

configuring the switch to controller node mapping and adding/removing controller nodes in response to changing traffic patterns.

Each SDN controller platform may have different performance characteristics. In such cases, it may be desirable to run different services on different controllers to match the controller performance characteristics with service requirements. To address this problem, we propose an architecture, FlowBricks, that allows network operators to compose an SDN control plane with services running on top of heterogeneous controller platforms.

1 INTRODUCTION

Distributed applications such as three-tier web applications and distributed big data applications (*e.g.,* Hadoop) running in large data centers support a bulk of the web and business services. Due to the distributed nature of these applications, the data center network characteristics directly impact application performance metrics such as query processing rate and completion time. This has led to several research initiates to improve the performance of data center networks. In the data plane, researchers have proposed topologies with full bisection bandwidth for data center networks based on multi-rooted trees [1, 2]. These topologies enable all end hosts can communicate with each other simultaneous at line rate without any bottlenecks at core links. At the control plane, SDN paradigm has gained popularity due to ease of management and faster convergence. However, a centralized SDN controller cannot manage large data center networks. So, researchers have proposed physically distributed SDN controller architectures that can handle the demands of large data centers. Data center network operators prefer to introduce new service through the SDN controller rather than middleboxes thus, adding to the complexity of designing an SDN controller. To address the growing number of network services and scalability challenges, researchers have proposed flexible modular open source SDN controller architectures which enable dynamic introduction and configuration of new services.

However, the unique characteristics of data center networks present new challenges. Recent experiments for characterizing data center traffic have found significant spatial and temporal variation in traffic volumes [1, 3, 4], which means that the data center network design cannot pre-assume a given traffic matrix and optimize the routing and forwarding for it. Recent trends therefore favor network fabric designs based on multi-rooted tree topologies with full bi-section bandwidth (or with low oversubscription ratios such as 4:1) such as the fat-tree topologies [2]. In such topologies, traditional single-path routing is inadequate since the full bi-section bandwidth guarantee assumes that all paths that exist between a pair of servers can be fully utilized. Thus, equal-cost multipath (ECMP) has been used as the *de facto* routing algorithm in these data centers. However, because not all flows are identical in their size (or their duration), this simple scheme is not sufficient to prevent the occurrence of hot-spots in the network. Several solutions (*e.g.,*, Hedera [5], Mahout [6]) focus on addressing this hot-spot problem by tracking and separating long-lived (elephant) flows along link-disjoint paths. However, it is fundamentally not always feasible to pack flows of different size/duration across a fixed number of paths in a perfectly balanced manner. A recently proposed solution called MP-TCP [7] departs from the basic assumption that a flow needs to be sent along one path, by splitting each flow into multiple sub-flows and leveraging ECMP to send them along multiple paths. Since MP-TCP requires significant end-host protocol stack changes, it is not always feasible in all environments, especially in public cloud platforms where individual tenants control the OS and the network stack. Further, it has high signaling and connection establishment complexity for short flows, which typically dominate the data center environment [3, 4].

Along the control plane, a few recent papers have explored architectures for building distributed SDN controllers [8–10]. While these have focused on building the components necessary to implement a distributed SDN controller, one key limitation of these systems is that the mapping between a switch and a controller is *statically configured*, making it difficult for the control plane to adapt to traffic load variations. Real networks (e.g., data center networks, enterprise networks) exhibit significant variations in both temporal and spatial traffic characteristics. First, along the temporal dimension, it is generally well-known that traffic conditions can depend on the time of day (e.g., less traffic during night), but there are variations even in shorter time scales (e.g., minutes to hours) depending on the applications running in the network. Second, there are often spatial traffic variations; depending on where applications are generating flows, some switches observe a larger number of flows compared to other portions of the network. Now, if the switch to controller mapping is static, a controller may become overloaded if the switches mapped to this controller suddenly observe a large number of flows, while other controllers remain underutilized. Furthermore, the load may shift across controllers over time, depending on the temporal and spatial variations in traffic conditions. Hence static mapping can result in sub-optimal performance. One way to improve performance is to over-provision controllers for an expected peak load, but this approach is clearly inefficient due to its high cost and energy consumption, especially considering load variations can be up to two orders of magnitude.

However, each SDN controller architecture will have its own performance characteristics which are best suited for certain applications. Some controllers may be suitable for high throughput while others may have low response times. In such cases, it may be desirable to run different services on different controllers to match the controller performance characteristics with service requirements. With the growing number and complexity of network services, all service implementations may not be available for a SDN controller platform. This, along with the incompatibility between SDN controller, motivates the need for a framework that can easily integrate services implemented on different SDN controller platforms.

In this dissertation, we propose three techniques to improve the scalability of data and control planes in data center networks. Along the data plane, we address scalability with growing network bandwidth demand. Along the control plane, we address scalability in two ways. We allow the controller to scale with changing control plane processing and traffic demands. We also enable the controller to scale with growing number of network services. One key design principle that we adopted in our solutions is that they should work with existing network protocols as far as possible. For example, a large majority of the traffic in data centers uses TCP [11]. So, it is important to improve data center traffic without requiring any changes to TCP. Similarly, OpenFlow has become one of the prominent standards for SDN-based control planes in data centers (*e.g.,* Google [12]). We tried to adhere to the OpenFlow standard as much as possible for maximum impact.

In the first part of the dissertation, we propose random packet spraying (RPS) as an effective traffic splitting technique for data center networks that have multi-rooted tree topologies. We key observation is that the duplicate-acknowledgment threshold and packet reordering detection schemes built into TCP are sufficient to make TCP robust to any packet reordering that may be introduced by RPS. Using a data center testbed with RPS implemented on NetFPGA switches, we show that RPS performs better than ECMP and similar to MP-TCP (for long-lived flows). We study the adverse effects of link failures on RPS and propose an approach based on Random Early Discard (RED [13]) to mitigate these adverse effects.

In the second part of this dissertation, we propose algorithms to dynamically scale the computing resources and throughput of a distributed SDN controller in response to control plane traffic demands. To achieve this, we propose a seamless switch migration algorithm, an algorithm to redistribute network load evenly among controller nodes and an algorithm to add or remove controller nodes.

Finally, we propose a framework for combining network services implemented on different SDN controller platforms. This is done without modifying the controllers themselves and relying entirely on the standardized southbound API.

1.1 Thesis Statement

This dissertation proposed techniques to improve the performance of the control and data plane in data center networks. We achieve this using new techniques that are based on existing network protocols.

The thesis of this dissertation is as follows: *We can improve the performance of data plane and control plane in modern data center networks using practical easy-to-deploy techniques.*

1.2 Contributions

This dissertation makes three major contributions towards improving data center network performance:

- Along the data plane, we propose random packet spraying as a technique that can significantly improve the latency and throughput of data center networks that have symmetric multi-rooted tree topologies. For dealing with failures that destroy the symmetry of the network topology, we propose SRED, a combination of RED and drop-tail queue management algorithms that reduces the negative impact of RED on network throughput.
- Along the control plane, we propose an OpenFlow-compliant switch migration algorithm that can seamlessly handover control of a switch from one controller node to another of a distributed SDN control plane. Using this algorithm as a building block, we built ElastiCon, a distributed SDN controller that can add or remove controller nodes in response to network traffic demands and evenly distributes the load among controller nodes.
- We designed and prototyped FlowBricks, a framework that allows network operators to combine best-in-class network services that may be running on different SDN control planes. FlowBricks is designed to operate in a way that is transparent to the controllers and does not require additional standardization.

1.3 Dissertation Organization

This dissertation contains five chapters. In Chapter 3, we show that RPS is an effective traffic splitting technique for data centers networks with symmetric multi-rooted tree topologies. Chapter 4 describes the design and experimental evaluation of ElastiCon, a scalable distributed SDN controller. In Chapter 5, we present FlowBricks, a framework for composing a control plane from services running on heterogeneous SDN controllers. Finally, we present our conclusions and potential directions for future work in Chapter 6.

2 BACKGROUND

Data centers are the core of the internet computing infrastructure. Their sizes range from a few hundred server owned by small and mid-sized corporations to over 100,000 servers operated by big firms and governments. These data centers may be used to run web-services or run big data applications. Data centers can benefit enormously by the economies of scale. This has two consequences. First, large corporations have consolidated their data centers into a few large facilities around the globe. Second, small firms find it more economical to rent computing and storage resources in large data centers rather than operate their own data centers. The scale of these data centers means that any performance and utilization improvements achieved here translate to large financial gains for the data center operators. This has spurred researchers to explore various avenues for improving all aspects of data centers including storage [14], network [12] and processing at end hosts [15]. In this dissertation, we focus on improving the scalability of networks that connect the host in a data center.

2.1 Data Center Network Performance

Data center network throughput and latency are important performance metrics since they have been shown to directly affect application performance [16]. Researchers have explored various directions for improving these metrics in data centers. New data center network topologies and switch architectures [17] try to address this problem at the physical layer. Such efforts have focussed on increasing bisection bandwidth while reducing costs by using commodity components. Since they use commodity hardware, traffic needs to be split across several low bandwidth links to utilize all the available bandwidth. [2] proposes a fat-tree topology which uses the entropy in the IP address bits to spread traffic across all available paths. VL2 [1] also uses a multi-rooted tree topology but has higher bandwidth 10Gbps links at the core and 1Gbps links at the edge switches. It uses virtual IP address and a scheme called valiant load balancing (VLB) to split traffic. Bcube [18] proposes a server centric architecture. Server, in addition to performing computation, act as relay nodes for each other.

Most data center network topologies have multiple paths between end hosts and require a traffic splitting technique to fully utilize all paths. The most commonly used technique is ECMP which does not make any assumptions about the underlying topologies. In ECMP, flows (as identified by the TCP 5-tuple) between a given pair of servers are routed through one of the paths using hashing; therefore, two flows between the same hosts may take different paths, and ECMP does not affect TCP congestion control. However, because not all flows are identical in their size (or their duration), this simple scheme is not sufficient to prevent the occurrence of hot-spots in the network. In a recent study [3], the authors find that 90% of the traffic volume is actually contained in 10% of flows (heavy-hitters); if two heavy-hitter flows are hashed to the same path, they can experience significant performance dip. Several solutions (*e.g.,*, Hedera [5], Mahout [6]) focus on addressing this hot-spot problem by tracking and separating long-lived (elephant) flows among link-disjoint paths. However, it is fundamentally not always feasible to pack flows of different size/duration across a fixed number of paths in a perfectly balanced manner. A recently proposed solution called MP-TCP [7] departs from the basic assumption that a flow needs to be sent along one path, by splitting each flow into multiple sub-flows and leveraging ECMP to send them along multiple paths. Since MP-TCP requires significant end-host protocol stack changes, it is not always feasible in all environments, especially in public cloud platforms where individual tenants control the OS and the network stack. Further, it has high signaling and connection establishment complexity for short flows, which typically dominate the data center environment [3, 4].

A large majority of network traffic in data centers uses TCP [11]. This has led researchers to investigate the performance of TCP in data center environments and propose improvements. The TCP incast problem was commonly observed in data center networks with MapReduce [19] or distributed storage workloads. ICTCP [20], a variant of TCP, tries to solve the incast problem by proactively adjusting the receive window before packet drops occur. To reduce queuing latency in the network, DCTCP [21] proposes using ECN in the network to provide multi-bit feedback to end hosts. D^2TCP [22] also uses ECN bits for congestion avoidance but uses deadlines to efficiently allocate bandwidth in a distributed manner. These TCP enhancements need to ensure that they can co-exists with existing TCP variants. But, they have limited utility to data center operators because network stacks on the end host are controlled by tenants in public data centers.

2.2 SDN and Data Center Networks

The benefits of softwared-defined networking (SDN) have led data center operators to adopt the SDN paradigm for managing their networks [12]. SDN moves the control plane logic out of the switches to a centralized entity called a controller. It uses a standardized protocol to configure the data plane in the switches. While OpenFlow [23] is currently the preeminent standardized protocol and switch specification for SDNs, researchers have proposed new switch architectures [24] that provide features not currently supported in OpenFlow. For example, [25] proposes allowing end hosts to embed a small list of instructions in a packet. These instructions are executed at every router along the path of the packet. This allows end hosts to query and change network state which can be used for a wide range of purposes.

The centralized SDN control plane provides many benefits to data center operators. It allows easy management of the network through a centralized controller interface. Researchers have proposed innovative ways to leverage the global view of the centralized controller to improve the manageability of SDNs. NetSight [26] introduces the idea of "postcards" the contain complete information about a packet header and switch forwarding state at a particular hop of the packet. By correlating information from postcards collected from different packets at each hop, the centralized server can infer a variety of problems in the network. VeriFlow [27] allows operators to verify network invariants in real time and across updates to the forwarding state in the network.

SDNs let data center operators introduce new services (such as NAT, traffic monitoring) with just a software upgrade of the controller instead of deploying and maintaining service-specific middleboxes. Data center operators complete control over the implementation of network services without relying on switch vendors. This has drastically reduced costs but increased the complexity of developing new network services for the centralized controller. New programming languages such as Pyretic [28] aim to simplify the development of new services by abstracting away switch hardware and protocol-specific details to common layer. To address backward-compatibility of SDNs with existing middleboxes, researchers have proposed techniques to enforce policies to route traffic through middleboxes [29].

3 RANDOM PACKET SPRAYING

In this chapter, we study the feasibility of an intuitive and simple multipathing scheme called *random packet spraying* (RPS), in which packets of every flow are randomly assigned to one of the available shortest paths to the destination. RPS requires no changes to end hosts, and is practical to implement in modern switches. In fact, many commodity switches today (*e.g.,* Cisco [30]) already implement a more sophisticated per-destination round-robin packet spraying technique.

RPS approach, however, can potentially result in reordering of packets that belong to a flow—a problem that is known to negatively interact with TCP congestion control[1], at least in the wide-area networks [31]. Specifically, packets in a given flow that traverse multiple paths with potentially different latencies may arrive at the receiver out of order, *i.e.,* later-sent packets may be received ahead of earlier-sent ones. Since TCP can not distinguish reordered packets from lost packets, it will trigger congestion avoidance by cutting down its congestion window leading to suboptimal performance. Because of the potential packet reordering and its implication on TCP, networking researchers as well as practitioners have cautiously kept packet spraying out of consideration for data center networks.

In this chapter, we make two *key observations* that together suggest RPS is unlikely to be affected by packet reordering and hence a promising multipathing scheme for data center networks. First, we observe that modern data center networks based on multirooted tree topologies tend to be *symmetrical*, which essentially causes links along multiple paths between a source-destination to be grouped into *equivalence classes.* As a result, paths between a source-destination pair are likely to exhibit similar queue build-up, keeping latencies roughly equal. In addition, data center net-

[1]This is the reason that the feature though supported in commodity switches is not turned on by default

works are often engineered to provide low latencies to service latency sensitive traffic anyway. Solutions such as DCTCP [21] and HULL [32] provide even lower latencies at the (slight) expense of throughput. Low end-to-end latencies help RPS since the worst case latency differential between two paths is also going to be small.

Second, standard TCP originally designed for the wide area Internet already has a built-in mechanism to tolerate mild packet reordering. In particular, TCP does not perform fast retransmit unless 3 duplicate ACKs (DUPACKs) arrive for the same packet. Newer implementations of TCP in the Linux kernel are even more robust to packet reordering. They use timestamps and DSACK options to detect spurious fast retransmissions. If a spurious fast retransmission is detected, TCP reverts the reduction in congestion window size. Also, the TCP duplicate ACK threshold is dynamically adjusted. Hence, even if some occasional reordering happens in the data center network under RPS, the reordering may only mildly affect TCP performance.

In this chapter, we conduct an empirical study to validate these observations and study the overall performance under RPS multipath routing using a *real testbed* comprising of hardware NetFPGA-based switches organized in 4-ary fat-tree topology. Our experiments indicate that our observations typically hold true in practice and as a result, RPS achieves much better network-wide TCP throughput than ECMP.

While our experiments above show that RPS works well in symmetric topologies, production data centers are prone to link failures which may disturb the overall symmetry of the network. Such asymmetry in the topology can potentially lead to unequal load on links leading to sub-optimal throughput of RPS. However, no prior studies have quantified the impact of failures on the performance of RPS in data center networks. Thus, in second part of this chapter, we conduct detailed empirical analysis of RPS under failure conditions. We observe that if RPS alone is used, it can lead to significantly lower throughput in failure scenarios. We observe however that if the queue lengths are kept sufficiently small using simple active queue management scheme such as Random Early Discard (RED), the performance of RPS can be much better, almost comparable to complex solutions such as MP-TCP.

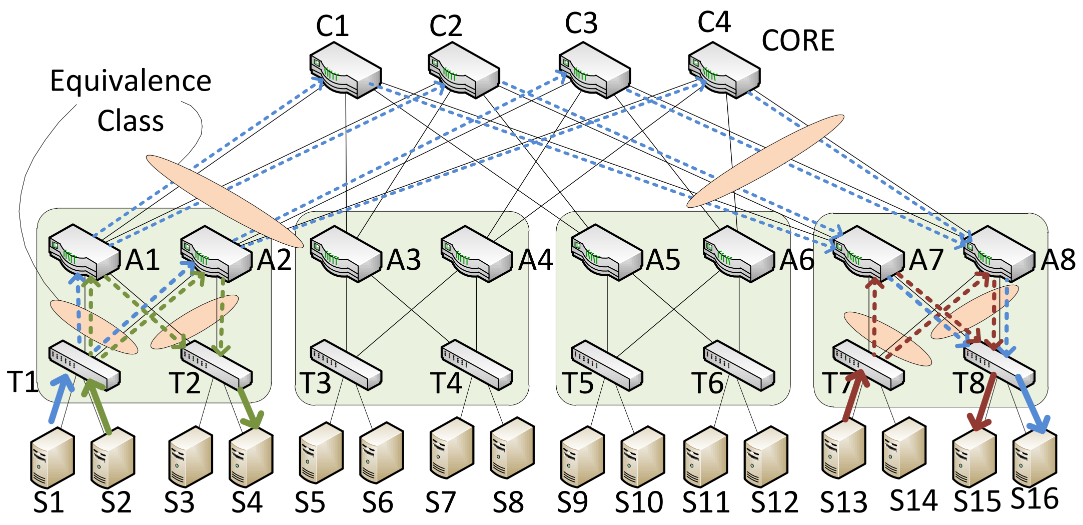


Figure 3.1.: Fat-tree topology with equivalence classes and imbalance with ECMP.

Contributions. In summary, the main contributions of the chapter include the following. (1) We conduct a first of its kind empirical study to debunk the myth that random packet spraying is inherently harmful to TCP, in the context of designing an effective multipathing scheme for data center networks. (2) Using a data center testbed with real RPS implementation over NetFPGA switches, we conduct detailed study on the reasons why RPS performs better than existing schemes such as ECMP and similar to MP-TCP (with long-lived flows). (3) We also study the adverse effect of link failures on the performance of RPS. Exploiting the key insight that smaller queues result in better performance even under failures, we propose an approach based on RED to mitigate these adverse effects.

3.1 Random Packet Spraying (RPS)

In this section, we start with an overview of RPS followed by theoretical analysis on why we expect RPS to perform well in data center networks.

3.1.1 RPS Overview

The basic idea of RPS is simple: Like ECMP, RPS uses all the equal-cost shortest paths between every source and destination pair. However, instead of hashing the flow key of a packet to determine the next hop for forwarding as in ECMP, RPS randomly spreads all packets that belong to each flow equally along different shortest paths. For example, in Figure 3.1, we show a flow from $S1 - S16$ that traverses the paths $S1 \rightarrow T1 \rightarrow \{A1, A2\} \rightarrow \{C1, C2, C3, C4\} \rightarrow \{A7, A8\} \rightarrow T8 \rightarrow S16$ to reach the destination. Thus, if the flow consists of 100 packets, roughly 25 packets will be routed through each of the four paths via core routers $C1 - C4$.

As shown before in literature [31], packet spraying can lead to severe packet reordering in the wide-area—the packets of a flow which take different paths may have orders of magnitude differences in latencies since there is no guarantees that the paths will be of equal lengths or have similar congestion. Even in data center environments, where latencies are low and uniform, RPS will potentially introduce packet reordering. TCP performs poorly in the presence of packet reordering. When the TCP sender receives three duplicate acknowledgments (DupACK), it assumes that a segment has been lost and reduces its congestion window size, which results in a drop in throughput. TCP maintains an estimate of round-trip (RTT) times. If paths have hugely varying latencies, TCP's RTT estimate will also be meaningless, which can lead to spurious retransmissions and timeouts. In fact, this concern of *potential* packet reordering is why none of the existing data centers use or existing proposals advocate the use of simple packet spraying schemes.

We make three key observations that indicate that packet spraying techniques like RPS are unlikely to result in significant packet reordering, and consequently should not affect TCP's performance in data center networks that employ multi-rooted tree topologies such as the fat-tree. Specifically:

Observation 1. In a multirooted tree topology like a fat-tree shown in Figure 3.1, links can be grouped together into *equivalence classes*. All links within each equivalence class have equal amount of load if all flows in the networks use RPS. Thus, even though each flow is routed along several paths, each of these paths is similarly loaded. So, the latency differential between these paths is expected to be quite small, and the amount of induced reordering due to packet spraying is likely to be small. (We analyze this in more detail next.)

Observation 2. TCP congestion control is robust to small amount of packet reordering in the network anyway. Given that TCP was designed for the wide area network, where some amount of reordering can happen due to failures and other events. The sender typically waits for 3 duplicate ACKs to infer that a loss event has occurred after which it performs fast retransmit and cuts its window in half. Besides this, the TCP implementation in newer Linux kernels detects spurious fast retransmission using the DSACK and timestamp options of TCP to rollback any erroneous reductions in the congestion window [33]. TCP also proactively avoids spurious fast retransmissions in the future by increasing the DupACK threshold [34].

Observation 3. Even if packet spraying using RPS induces slightly more fast retransmits compared to say a flow based technique like ECMP, the extra loss in throughput, due to the sender reducing its congestion window by half every time a fast retransmit event occurs, can be a small penalty compared to the better usage of the total aggregate available bandwidth across all paths. Thus, RPS' overall performance will be likely better than that of ECMP.

Further, data center operators are increasingly more concerned about end-to-end latencies. Thus, future data center designs are likely to ensure low and uniform latencies, using mechanisms such as HULL [32], DCTCP [21], DeTail [35]. If latencies across all paths are low and uniform, TCP end-host can maintain a single estimate of RTT for all paths.

In spite of low latencies in data center networks and improvements to TCP, researchers have focused mainly on load balancing schemes which avoid packet reordering. No measurement studies have been conducted to study the impact of these improvements on packet spraying in data center networks. Our analysis shows that

TCP is able to perform well with packet spraying in a data center environment, as long as packets are sprayed over equal length paths and queue lengths are kept almost equal along all paths. We hope that this result will encourage more research in simple packet spraying techniques for data centers.

3.1.2 Analysis

We formalize the concept of equivalence classes stated in Observation 1 above, which gives a key reason why different from in the Internet, significant packet reordering is unlikely to happen when RPS is running in data center networks which typically employ multi-rooted tree typologies such as fat trees.

When RPS is used to route packets between a source and a destination via all equal-cost paths, an *equivalence class* comprises all outgoing links from the switches at the same hop along all the equal-cost paths. For simplicity, we exclude links to/from end hosts (leaves in the tree) in the discussion. In a depth-h K-ary fat tree (each switch has K ports), each flow goes through $2h$-hop equal-cost paths and passes through $2h$ equivalance classes of links. Note different source-destination pairs can share some equivalance classes. Together, there are $2h$ *types* of equivalance classes in a depth-h fat tree. In particular, there are 4 equivalance classes in the depth-2 fat tree in Figure 3.1:

Type 1: A Type 1 class consists of the links from a ToR switch, ToR_i, to the $\frac{K}{2}$ aggregate switches Agg_j within the same pod.[2]

Type 2: A Type 2 class is the mirror image of a Type 1 class, and consists of the links from the $\frac{K}{2}$ aggregate switches Agg_j within a pod, to a ToR switch, ToR_i.

Under RPS, for an X-packet flow, the expected number of packets that will be routed through each of the $\frac{K}{2}$ links in a Type 1 or Type 2 equivalence class is $\frac{2X}{K}$.

Type 3: A Type 3 class consists of the links from all the aggregate switches Agg_i within a pod, to all $\frac{K^2}{4}$ core switches, C_n.

[2]The set of switches {T1,T2,A1,A2}, {T3,T4,A3,A4}, etc. in Figure 3.1, are referred to as pods in the fat-tree.

Type 4: A Type 4 class is the mirror imagine of a Type 3 class, and consists of the links from all the core switches C_n, to all aggregate switches, Agg_j, within a pod.

Under RPS, for an X-packet flow, the expected number of packets that will be routed through each of the $\frac{K^2}{4}$ links in a Type 3 or Type 4 equivalence class is $\frac{4X}{K^2}$.

Example. Consider the two paths between S1 and S5 in Figure 3.1. There are four equal-cost paths between them. The first hops of all paths form to the Type 1 equivalence class ($T1 \rightarrow A1$, $T1 \rightarrow A2$), the second hops of all paths belong to the Type 3 equivalence class ($A1 \rightarrow C1$, $A1 \rightarrow C2$, $A2 \rightarrow C3$, $A2 \rightarrow C4$), and so on. This hop-by-hop equivalence holds for paths between all hosts in the fat tree even if they are in different pods, in the same pod, or under the same ToR switch.

The equal spread of packets of each flow among the links in its hop-by-hop equivalence ensures that, given any set of flows, load and hence the queue lengths (measured in number of packets) among the links in each eqivalent class stays the same. This in turn implies that for a given flow, its packets traversing different paths will encounter the same queuing delay, and hence the same end-to-end delay. Thus, the receiver will observe only a few reordered packets due to small differences in queue lengths introduced by (1) difference in packet sizes; (2) flow sizes are not always in multiples of the number of paths; and (3) timing issues. However, these issues are expected to cause only a small queue length differential which results in a small amount of reordering within the network. We experimentally confirm this in Section 3.2.

3.2 Evaluating RPS

In this section, we evaluate RPS using a real hardware testbed. We first discuss the testbed configuration and our implementation of RPS and ECMP. We then provide comparisons of RPS with and ECMP and MP-TCP. Finally, we empirically confirm the three observations made in the previous section that explain the good performance of RPS in our testbed.

3.2.1 Testbed Configuration

Our testbed has 36 servers connected in a 4-ary ($k = 4$) fat-tree [2] topology (as shown in Figure 3.1). All the servers are equipped with 4GB RAM, Intel Xeon 2.40GHz quad-core processors running Centos 5.5 and two 1Gbps Ethernet ports. We have 20 NetFPGA boards, each deployed on a server, and interconnected in a fat-tree topology via 1 Gbps Ethernet links. Rest of the 16 servers form the endhosts connected to this network. A fat tree has an oversubscription ratio of 1:1. Removing two of the four core switches would have resulted in an oversubscription ratio of 2:1 but it would have reduced path diversity; there would be just two paths between hosts in different paths, which can bias our results significantly. Other oversubscription ratios (4:1, 8:1) would not be possible even. To overcome this, we emulate oversubscription of approximately 4:1 (and 8:1) by rate-limiting the core links to 230Mbps (and 115Mbps). The seemingly arbitrary choice of 230Mbps (instead of 250Mbps) stems from the limitations of the NetFPGA rate limiter, which allows only a few discrete values to choose from.

Implementation of RPS and ECMP

We implemented RPS and ECMP on NetFPGA switches by modifying the code base already provided by NetFPGA. For a packet arriving at the switch, we generate a random number (using the library provided by NetFPGA) to determine the output port (among all eligible output ports) to which the packet is forwarded. Implementing this is quite simple; we needed only about 100 lines of verilog code to implement this technique. RPS is a purely switch-based solution and does not require any help or modification at the end hosts.

Implementation of MP-TCP

To enable MP-TCP, we deployed the publicly released Linux kernel for MP-TCP [36] at the end hosts. This kernel still has a few performance and stability problems. For instance, we observed kernel panics sometimes when MP-TCP was handling many short-sized flows simultaneously. This prevented us from running experiments involving many short flows with the MP-TCP kernel. For long flows, we observed more stable results for MP-TCP. MP-TCP has also been noted to have a sub-standard performance with short flows because the control overhead of setting up and terminating many subflows becomes significant. For the above reasons, we present MP-TCP results for long flows in this dissertation. Since ECMP performs well with short and long flows, we compare RPS with ECMP in experiments involving both short and long flows.

3.2.2 TCP Throughput under Packet Spraying

We first measure the throughput obtained by long lived TCP flows in a random permutation matrix (similar to [7]). In such a setup, each host in the topology is either a sender or a receiver of exactly one TCP flow. All senders are randomly paired with receivers. A netperf client running at the sender sends a TCP flow to its receiver for the duration of the experiment. We measure the average throughput as a percentage of the ideal throughput and also compare performance of TCP flows under different schemes.

Figure 3.2 clearly depicts the gain in throughput experienced by TCP flows under a packet spraying technique (RPS). Even under different degrees of oversubscription, the throughput obtained under RPS is higher than those measured under MP-TCP or ECMP-like techniques. The low average throughput in case of ECMP-based forwarding can be attributed to the fact that two or more TCP flows may be forwarded over the same core link which becomes a bottleneck. For the entire flow duration of the flow, that link remains the hot spot in the network while leaving other links underutilized. Due to static allocation of paths in ECMP, if some of the flows are unlucky and are routed through a congested link, then they suffer permanently for the entire duration resulting in poor throughput.

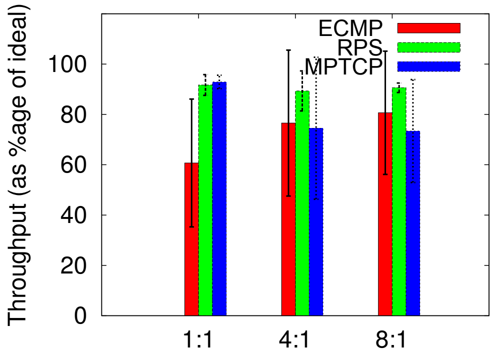


Figure 3.2.: Throughput for permutation matrix

Under RPS, average throughput achieved is about 90% of the ideal bandwidth in all 3 cases with different oversubscription ratios. Figure 3.2 also demonstrates that the variance in throughput obtained by different TCP flows is small. MP-TCP also achieves about 90% in case of a non-oversubscribed topology (subscription factor 1:1). This is consistent with results reported in [7] for a similar experimental setup. In case of oversubscribed topology though, the average throughput achieved by MP-TCP flows seems to suffer and it decreases from 90% to about 75%. This poor performance may be an artifact of MP-TCP itself or the released implementation of MP-TCP; unfortunately, there is no easy way for us to know precisely at the moment.

To study the effect of path diversity on RPS, we repeated the above experiment in simulation using fat trees with ($k =$) 6 and 8 pods. The number of paths between end hosts is 9 and 16 respectively ($k^2/4$). Intuitively, when the number of paths increases, the probability of packet reordering in packet spraying increases. However, we observed that the drop is not substantial showing that our analysis in Section 3.1.2 still largely holds.

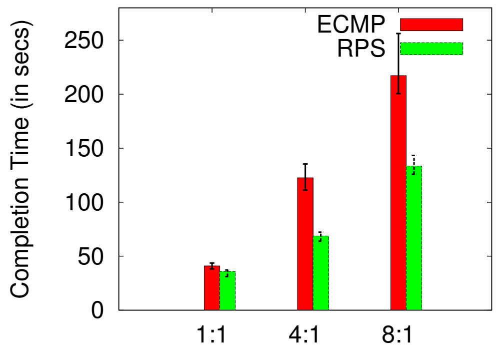
3.2.3 Data Transfer Time

We repeat the experiment performed in [7] (but with mixed short and long flows) to study how much time TCP takes to transfer the same amount of data under different schemes. This experiment shows the ability of the underlying mechanism to consume bandwidth more efficiently to transfer the same amount of data. In this experimental setup, each end host executes two clients which have to transfer 2GB of total data, which is divided into many flows with flow sizes drawn from the real data-center flow size distribution reported in [1]. A client sends these flows in sequence to randomly chosen destinations. The client forks a new netperf client for each flow. All clients begin simultaneously. We plot the median, and first/third quartiles of the completion time of all clients in Figure 3.3(a).

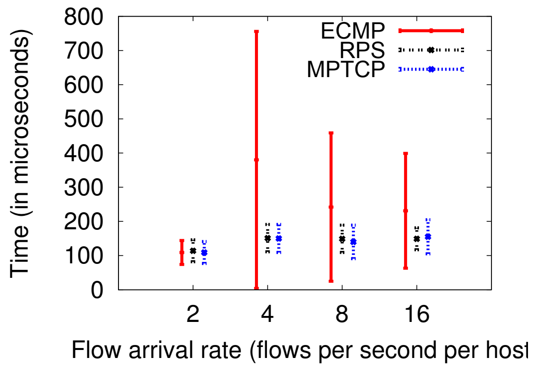
We observe that TCP flows are able to complete faster under RPS as compared to ECMP. (We cannot do this experiment with MP-TCP as it is unstable when there are large number of concurrent connections.) With 1:1 oversubscription, we observe that ECMP and RPS perform equally well. This is because in such a topology and flows being setup between random pairs of hosts, the edge links are more likely to be the bottleneck than the core of the network. So, TCP does not benefit from a better traffic splitting technique. In case of 4:1 or 8:1 oversubscribed networks, the packet spraying technique helps TCP flows to utilize the available capacity in a much more efficient manner in spite of the reordering. Hence, the time to transfer the same amount of total data is 25% smaller in case of RPS than ECMP.

3.2.4 Packet Latencies

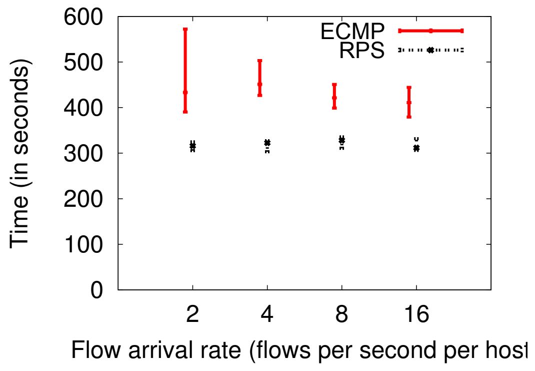
Packet latency is another important metric for flows in data center networks. Recent works like [21, 32] have focused on reducing packet latencies in the network so that applications can satisfy SLAs (service level agreements). To study the effect of packet spraying on packet latencies we ran background traffic between 14 (out of 16) end hosts in our testbed. The flow sizes for background traffic were drawn from



(a) Completion Time



(b) Ping RTT



(c) Hadoop Shuffle Time

Figure 3.3.: Performance of RPS with different traffic patterns.

the distribution in [1]. The flow arrival rate followed an exponential distribution and variable mean. We sent 200 back-to-back ping packets between the two hosts that did not carry background traffic. The two hosts which do not carry background traffic exchange ping packets. For ECMP, MP-TCP and RPS, a ping packet randomly takes one of the 4 paths between the end hosts. Buffers at the two end hosts are always empty because they do not transmit or receive any of the background traffic. So, ping packets experience similar latencies at end hosts. Since the ping packets are sent back-to-back, we can assume that packets taking the same path also observe very similar latencies. So, the variation in latencies between packets is almost entirely due to variation in latencies between different paths in the network. Figure 3.3(b) shows the mean RTT for the 200 back-to-back ping packets and the errorbars show the mean deviation reported by ping. We observed that the latency varied widely with ECMP indicating that different paths between the two hosts different loads. Packets experience similar mean latencies with RPS and MP-TCP, but experience higher variance with MP-TCP.

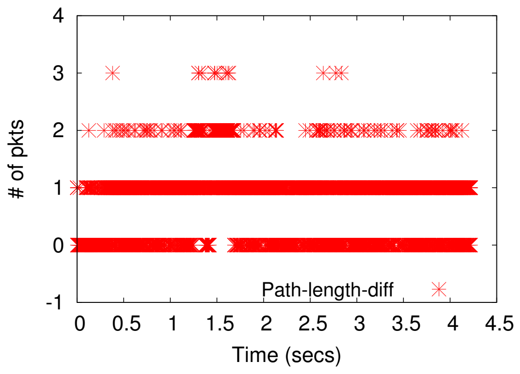
3.2.5 Effect on MapReduce

In order to quantify the impact of packet spraying on applications, we run Hadoop Sort application on 4 of the 16 end hosts in our testbed (other 12 hosts have background traffic between them as before). To emulate a network constrained cloud application, we reduce the bandwidth of each link to 115Mbps but kept the oversubscription ratio at 1:1.

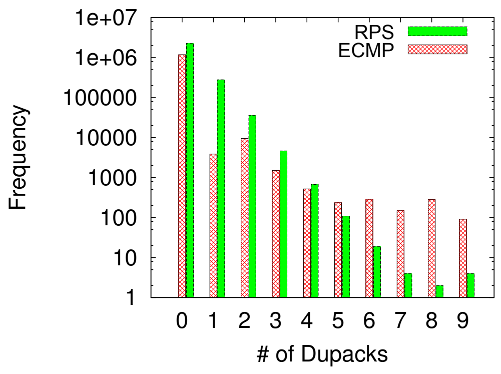
Figure 3.3(c) shows the time taken for the shuffle phase of Hadoop sorting 4GB of data averaged over 3 runs. On the x-axis, we vary the intensity of background traffic (that is, flow arrival rate of background traffic). We observe a 20% to 30% reduction in shuffle time with RPS. Also, the variance in completion time is much smaller with RPS than ECMP. Since a fat-tree is provides full bisection bandwidth, end hosts running Hadoop can communicate with each other at full line rate even when background traffic intensity increases. So, increasing background traffic intensity does not affect shuffle phase completion time. RPS completes the sort phase quicker than ECMP because it is able to utilize the available bisection bandwidth more efficiently than RPS. We were not able to perform this experiment with MP-TCP due to the stability issues with MP-TCP implementation.

3.2.6 Analysis of Packet Spraying

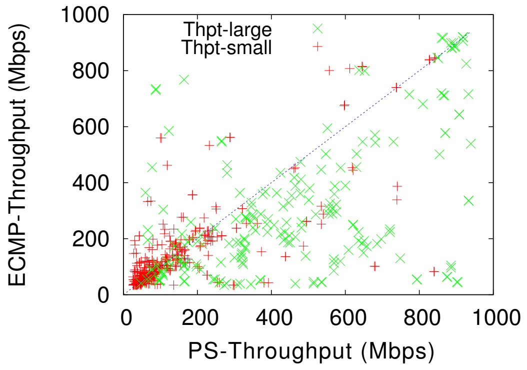
Now, we conduct experiments to validate our analysis and our understanding of why TCP performs well under packet spraying in data center networks. Specifically, we empirically validate our key observations made in Section 3.1 using experiments conducted on our testbed. In these experiments, each end host starts new flows with



(a) Q-length diff. across a src-dest pair



(b) # of consecutive DupAcks



(c) Throughput comparison with ECMP

Figure 3.4.: Microscopic analysis to validate our understanding of RPS performance.

start times based on a Poisson distribution with 2 arrivals per second. As before, flow sizes are drawn from the distribution reported in [1]. The traffic matrix was executed two times, once each for ECMP and RPS. Both times, the random number seeds were initialized to the same value to ensure that flows of the same size were started at the same time in both runs, allowing us to make a flow-by-flow comparison between RPS and ECMP.

Queue length differential

In Section 3.1.2, we argued that packet reordering will be limited because all paths between a pair of hosts have similar latencies. Latencies are largely determined by queue lengths that a packet encounters at every hop. We polled the queue lengths at every hop along a path between a pair of hosts in our testbed. We were able to poll about 1000 times per second; this is the maximum rate allowed by our NetFPGA platform. We used NTP to synchronize timestamps at all switches. By summing the queue lengths of all hops, we determined the path queue-length, that is, the total queue-length that a packet would encounter if it were forwarded along that path. We did this for all paths between a source destination pair and plotted the instantaneous difference between the highest and the lowest path queue-lengths in Figure 3.4(a). In a perfectly balanced network, this path queue-length differential will always be zero. However, in Figure 3.4(a), the queue-length differential is less than or equal to one 93% of the time. Flows between the pair of end hosts under observation may experience some reordering when the path queue-length differential increases to two or three, but that is relatively infrequent (less than 7%).

DupACKs

We now measure the number of dupACKs that an end host will receive. We log the number of dupACKs received at the sender and plot Figure 3.4(b). The x-axis shows the number of dupACKs that the sender received for a particular TCP segment. For both ECMP and RPS, the sender received no dupACKs for almost one million TCP segments. RPS received exactly one dupACK almost 200,000 TCP segments, and exactly two dupACKs for 30,000 and so on. We see that the frequency of k dupACKs reduces exponentially with increasing k.

TCP does not enter fast-retransmission until it sees greater number of dupACKs than the dupAckThreshold (default is 3). So, the first three bars in the figure will not lead to a drop in throughput. Since the number of dupACKs reduce exponentially, we observe that fewer that 55,000 (about 2%) of the transmitted TCP sequence numbers cross the three dupACK threshold. Surprisingly, we find a similar order of magnitude dupACKs in ECMP. However, dupACKs in ECMP are entirely due to packet losses and therefore cause a drop in throughput. But, dupACKs in RPS are due to a combination of reordered and lost packets. While it is difficult to ascertain the exact number of reordered and lost packets, note that dupACKs due to reordered packets are handled well as stock Linux TCP implementation has adaptive dupACKThreshold to reduce spurious reductions in TCP congestion window due to reordered packets. In any case, RPS should perform, if not better, no worse than ECMP because of dupACKs. But, flows have higher available bandwidth in the case of RPS than ECMP due to the availability of combined bandwidth across all sub-paths, which increases the performance of RPS compared with ECMP.

Effect on throughput of individual flows

We compare the performance of large and small flows in RPS and ECMP. It is generally expected that small flows should obtain good throughput under ECMP; it is the large flows that usually suffer. In Figure 3.4(c), we plot the throughput observed by the same flows under ECMP and RPS. The x-coordinate of a point corresponds to a flow's throughput under RPS while the y-coordinate is its throughput with ECMP. Points below the diagonal line indicate higher throughput with RPS. The green crosses represent large flows (greater than 1MB) while the red pluses are for short flows. From the graph, we can clearly see that large flows benefit most with RPS, while small flows perform equally well with both ECMP and RPS. Although only 10% of the flows in the distribution are large, the difference in throughput is significant enough to affect applications. Also, the number of bytes belonging to large flows is a larger fraction of the overall network utilization than the number of large flows. We also conducted a similar experiment as in Section 3.2.3, and observed a significant reduction in data transfer time. This experiment has an important implication: RPS cannot benefit much by treating large and small flows differently. Some earlier proposals [5] work only on large flows and let ECMP handle smaller flows. Since RPS handles small flows just as well as ECMP, we can apply RPS to all flows in the network. This avoids the additional complexity of trying to identify large flows in the network.

3.3 Handling Asymmetry

So far, we have investigated the behavior of packet spraying in symmetric multirooted tree topologies. But in the real world, a data center network may not be symmetric at all times. The data center may have an asymmetric topology to start with. Even in networks with symmetric topologies, asymmetries may arise due to various reasons. For instance, a failure condition (link/switch failure or link degradation) can result in an asymmetric topology. Under the above scenarios, different paths between a pair of end hosts in the network may see different levels of congestion.

In Section 3.2.6, we showed that the queue length differential is low in a symmetric network topology. However, the queue length differential can be significant in an assymmetric network topology due to the absence of equivalence classes. Below, we first show how flows suffer from this queue length differential. We then show how we can force queue lengths to be almost equal using existing techniques that prevent queues from growing large.

3.3.1 Problem Illustration

In this section, we use a very simple setup to demonstrate how asymmetries can impact RPS. We consider two scenarios (failures and mixture of routing strategies) which we believe are common in data centers.

Two Flow Experiment: We describe our experimental results obtained using the testbed with an oversubscription ratio of 4:1. Results for the 1:1 and 8:1 are similar and are hence skipped for brevity. To illustrate the problem, we consider only 2 flows in the network: flow F1 from S1 to S10 and and flow F2 from S15 to S9 as shown in Figure 3.5. For this experiment, ignore the shaded box indicating data transfer. Both flows use RPS and last for the entire duration of the experiment. As expected, F1

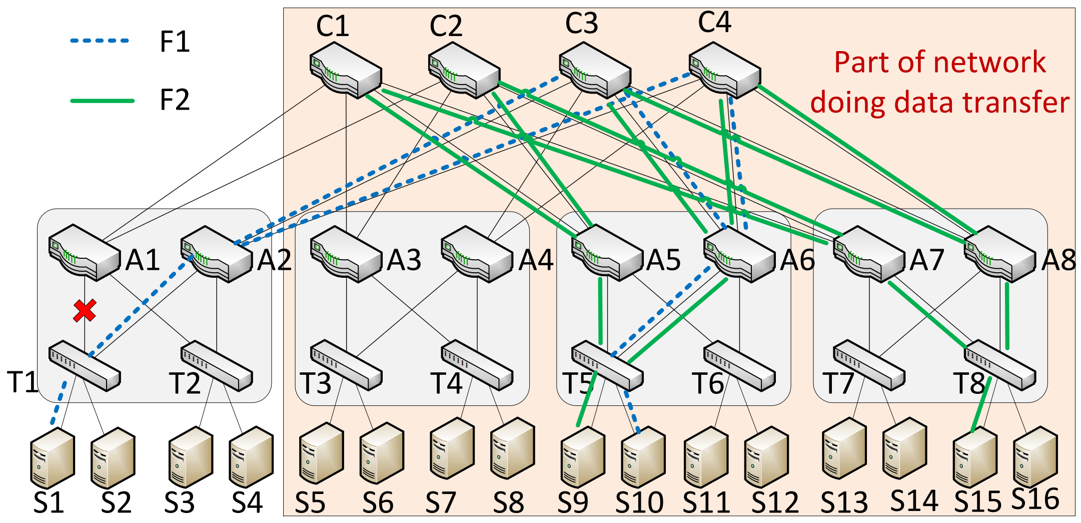


Figure 3.5.: Experimental setup in case of failure

and F2 observe a throughput of 407Mbps. Now, we fail the link between T1 and A1 which lies along the path to F1's destination. So, while RPS is able to evenly spray F2's packets over all four paths to S9, F1's packets are sprayed over the remaining two paths to S10. This ensures F2's packets will see higher queue lengths along paths shared with F1 and lower queue lengths along other paths. As a result, F2's throughput drops from 407Mbps to 155Mbps.

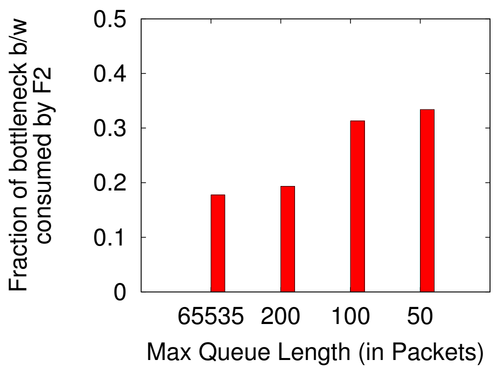


Figure 3.6.: Queue length with RPS in two flow experiment.

3.3.2 Key Observation

We first demonstrate that the drop in throughput is indeed due to difference in queue lengths and reducing it alleviates the problem. To do so, in the two-flow experiment setup described above, we statically limit the buffer size at the output buffer of all ports in the network. When the output buffer is unrestricted, the queue length differential reaches more that 300 packets. But, limiting the output buffer limits the queue length differential too. However, it also causes a drop in throughput due to lower link utilization. To find out throughput loss due to queue length differential, we observe the fraction of the bandwidth that F2 receives at the bottleneck link. A bottleneck link is always one of the links shared by both flows. So, F2 should always get 50% of the bottleneck bandwidth. However, as Figure 3.6 shows, when the queue length is unrestricted (corresponding to the bar at 65535), F2 consumes just 17% of the bandwidth. Reducing buffer sizes reduces the queue length difference between the paths that carry F2's packets. As a result, it is able to sustain a higher throughput, reaching 35% when buffer sizes are restricted to 50 packets. This demonstrates that the higher the queue length differential, the lower the throughput.

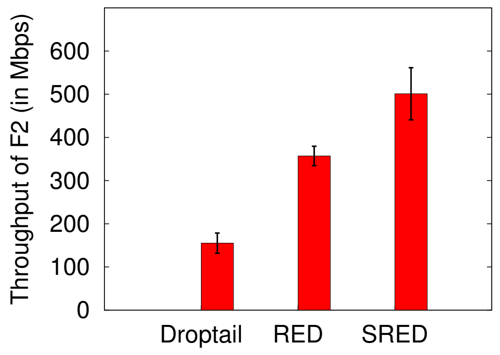
3.3.3 A Practical Solution: Keeping Queue Lengths Equal

The queue length differential due to assymetries is the main factor that impacts RPS performance, but we observe that the extent of this impact is very much dependent on the maximum size a queue is allowed to grow. Modern data center operators try to keep queue lengths to the minimum to keep end-to-end latencies low and predictable, which helps RPS. Standard active queue management techniques like RED and numerous newer solutions [21, 32, 37] can be used to achieve this. Later in this section, we show how RED improves the performance of RPS in the presence of link failures. We were not able to test out RPS with newer solutions since they are still emerging.

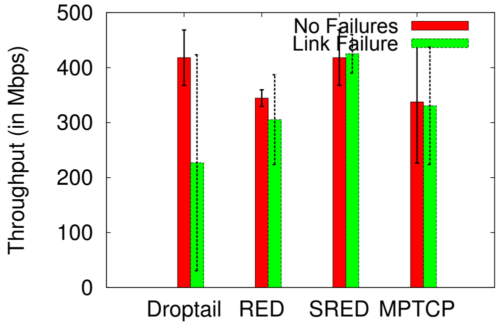
RED probabilistically drops packets as soon as the queue length crosses some threshold. Setting the threshold too low results in being aggressive in dropping packets, and hence, queue length is kept relatively low. However, the total throughput and utilization is reduced as well. We also introduce a new variant of RED that limits the ill-effects of RED while still reducing queue length differentials. We call this variant Selective-RED (SRED). SRED *selectively* enables RED only for flows that induce a queue length differential. These are flows which do not use all the multiple paths (like F1 in the two flow experiment) because of link failures or otherwise. Intuitively, restricting the queue length share of these flows should reduce the queue length differential. Packets of flows using all the paths (like F2) continue to use droptail since these packets do not contribute to any queue length differentials.

We envision implementing SRED using packet marking and a topology aware centralized fault manager. When a link fails, the centralized fault manager configures end hosts or ToR routers to mark all packets of flows affected by that failure. Marking can be done using TOS bits in the IP header. Downstream routers employ RED only on marked packets, thus emulating SRED. Other packets are queued and dropped using droptail policy; this limits the ill effects of RED to only those queues which induce queue imbalances in the network. While centralized controllers like Hedera [5] need to respond to new flow arrivals, the fault manager responds only to topology changes. Hence, it can scale well to larger networks.

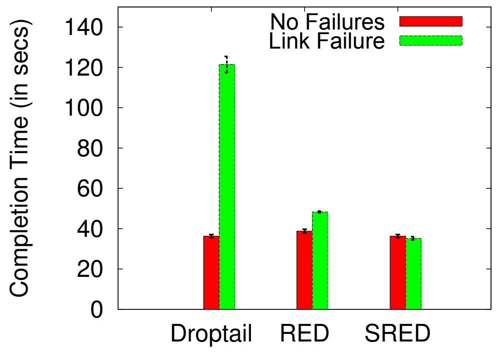
SRED requires changes to switches however. Logically, each output port will need to maintain 2 queues, one using droptail for unmarked packets, while the other using RED for marked packets. We can of course implement SRED using a single physical queue itself; unmarked packets are inserted if space is available while marked packets are queued probabilistically using RED. Only the number of marked packets in the queue are used to calculate the average queue length used by the RED algorithm. We now evaluate this idea but a detailed analysis of SRED is left for future work.



(a) Throughput in 2 flow experiment



(b) Throughput with permutation matrix



(c) Completion time for data transfer

Figure 3.7.: RPS performance with a link failure.

Two Flow Experiment

We repeat the above experiment with RED (threshold$_{max}$ = 20, threshold$_{min}$ = 10, p$_{max}$ = 0.1) enabled at all switches. We observe the throughput of flow F2 under two scenarios: RED is applied to both flows and SRED (RED applied only to F1's packets). We want to show that the reduced throughput experienced by F2 is entirely due to the unequal queue lengths induced by F1's packets. Hence, limiting F1 in the routers' queues should be sufficient to restore F2's throughput. As seen in Figure 3.7(a), F2's throughput falls from 417Mbps in the ideal case to less than 200Mbps when packets from flow F1 are spread over 2 (out of 4) paths. This is expected as F1 creates high queue length differential for packets of flow F2 which are spread over all the 4 paths.

However, when we restrict the queue length of the switches using RED, the throughput of F2 increases (to 360 Mbps) resulting from lesser queue length differential. It is still low (compared to ideal) as we are limiting the link utilization by limiting the total queue length. Under SRED, we can clearly observe that F2 gets close to ideal throughput. When using 2 paths under SRED, flow F1 gets a throughput (not shown in figure) close to 195 Mbps, almost half of the ideal. This is by design since we believe it is acceptable for flows that are directly impacted by the failure to suffer throughput loss, but we want to ensure that other flows not directly impacted by the failure continue unaffected.

Permutation Matrix under Failure

We repeat the experiment with a permutation matrix on a 4:1 oversubscribed topology. As before, we fail the link between A1 and T1. Figure 3.7(b) shows the average throughput of all the flows which are not affected by the failure and spray their packets on all four paths. We also compare their average throughput with that in a topology without failures. With RPS over droptail, the mean throughput of these flows almost halves as compared to that without failure. RPS in presence of RED reduces the average throughput of these flows even in the absence of failures (due to lower link utilization). But, when a failure happens, the mean throughput is not affected by a lot (changes by less than 10%) due to limits on queue length differential.

SRED is exactly like droptail when there are no failures in the network (since no flow is subjected to RED). In case of failures, flows affected by failure (using only 2 out of 4 paths in our case) are handled using RED. These flows get an average throughput of 156Mbps (not shown in graph). Other flows are not affected by the failure at all and they continue to achieve a high share of throughput. The figure clearly shows that the failure has a negligible impact of the average throughput of these flows.

Data Transfer Time under Failure

We repeat the experiment in Section 3.2.3 with unequal traffic splitting. We use the same oversubscribed fat-tree (4:1) as discussed above. However, we transfer 1GB of data between hosts in 3 pods as shown in Figure 3.5. We also inject flow F1 from the 4^{th} pod to a randomly selected host in one of the 3 pods. For this experiment, ignore flow F2 shown in the figure. We plot the average time taken by the last host to complete the transfer in Figure 3.7(c) and compare the with and without failure cases. We average this over 10 runs and the errorbars represent the standard deviation.

A link failure in the 4^{th} pod should not affect the 1GB data transfer because its traffic does not traverse this pod. However, the flow which is injected from the 4^{th} pod creates the queue length differential as it is sprayed over two instead of four core switches (the other two core switches are inaccessible to this flow due to the failure). This imbalance greatly increases the data transfer completion time in RPS with droptail from 36 seconds to 120 seconds. In case of RED, the data transfer completion time increases marginally from 36 to 48 seconds for the same failure scenario. With SRED, a failure in the 4^{th} pod has virtually no effect on the traffic of the 3 pods and the completion time remains the same.

3.4 Related Work

The most related to our work are those mechanisms that rely on flow-level traffic splitting such as ECMP and Hedera [5]. Mahout [6] is a recent scheme that uses end-host mechanisms to identify elephants, and uses flow scheduling schemes similar to Hedera. BCube [18] proposes a server-centric network architecture and source routing mechanism for selecting paths for flows. When a host needs to route a new flow, it probes multiple paths to the destination and selects the one with the highest available bandwidth. Techniques like Hedera, Mahout and BCube which select a path for a flow based on current network conditions suffer from a common problem: When network conditions change over time, the selected path may no longer be the optimal one. To overcome this problem, they periodically re-execute their path selection algorithm. VL2 [1] and Monsoon [38] propose using Valiant Load Balancing (VLB) at a per-flow granularity, but they too do not split an individual flow across multiple paths.

Two research efforts propose traffic splitting at a sub-flow granularity. The first effort, MP-TCP [7], splits a TCP flow into multiple sub-flows at the end hosts, which are routed over different paths in the network using ECMP. The receiving end host aggregates the TCP sub-flows and resequences packets. MP-TCP requires end-host changes which may not be feasible in all environments. It also suffers from high overhead for short flows that dominate data centers. The second effort, FLARE [39], exploits the inherent burstiness of TCP flows to break up a flow into bursts called flowlets, and route each flowlet along a different path. However, FLARE requires each router to maintain some per-flow state and estimate the latency to the destination. We did experiment with some simple variants of FLARE, such as keeping a small number of packets of a flow go through the same path. But we observed that any simple variant of FLARE does not achieve as good a throughput as RPS.

3.5 Summary

We showed how a simple packet-level traffic splitting scheme called RPS not only leads to significantly better load balance and network utilization, but also incurs little packet reordering since it exploits the symmetry in multirooted tree topologies. Furthermore, such schemes have lower complexity and readily implementable, making them an appealing alternative for data center networks. Real data centers also need to deal with failures which may disturb the symmetry, impacting the performance of RPS. We observed that by keeping queue lengths small, this impact can be minimized.

We exploited this observation by proposing a simple queue management scheme called SRED that can cope well with failures.

4 ElastiCon: AN ELASTIC DISTRIBUTED SDN CONTROLLER

SDNs can provide a wide range of services in data center networks. For example, [40] present a number of user cases such as bandwidth-on-demand and virtual data center. A bandwidth-on-demand service can provide guaranteed bandwidth to suit applications demands without requiring any manual configuration. The virtual data center service can aggregate traffic from VMs over the network connections/tunnels that have been discovered and provisioned through SDN. Such services require applications to communicate their network demands to the SDN controller and the SDN controller configures network devices to meet those demands. For large data centers, the SDN controller can become a performance bottleneck for such services.

A few recent papers have explored architectures for building logically centralized but physically distributed SDN controllers [8–10] to solve this problem. While these have focused on building the components necessary to implement a distributed SDN controller, one key limitation of these systems is that the mapping between a switch and a controller is *statically configured*, making it difficult for the control plane to adapt to traffic load variations. Real networks (e.g., data center networks, enterprise networks) exhibit significant variations in both temporal and spatial traffic characteristics. First, along the temporal dimension, it is generally well-known that traffic conditions can depend on the time of day (e.g., less traffic during night), but there are variations even in shorter time scales (e.g., minutes to hours) depending on the applications running in the network. For instance, based on measurements over real data centers in [11], we estimate that the peak-to-median ratio of flow arrival rates is almost 1-2 orders of magnitude[1] (more details in Section 4.1). Second, there are often

[1]This analysis is based on the reactive flow installation although our design supports proactive mode as well.

spatial traffic variations; depending on where applications are generating flows, some switches observe a larger number of flows compared to other portions of the network.

Now, if the switch to controller mapping is static, a controller may become overloaded if the switches mapped to this controller suddenly observe a large number of flows, while other controllers remain underutilized. Furthermore, the load may shift across controllers over time, depending on the temporal and spatial variations in traffic conditions. Hence static mapping can result in sub-optimal performance. One way to improve performance is to over-provision controllers for an expected peak load, but this approach is clearly inefficient due to its high cost and energy consumption, especially considering load variations can be up to two orders of magnitude.

To address this problem, we propose ElastiCon, an *elastic distributed controller architecture* in which the controller pool expands or shrinks dynamically as the aggregate load changes over time. While such an elastic architecture can ensure there are always enough controller resources to manage the traffic load, performance can still be bad if the load is not distributed among these different controllers evenly. For example, if the set of switches that are connected to one controller are generating most of the traffic while the others are not, this can cause the performance to dip significantly even though there are enough controller resources in the overall system. To address this problem, ElastiCon periodically monitors the load on each controller, detects imbalances, and *automatically* balances the load across controllers by migrating some switches from the overloaded controller to a lightly-loaded one. This way, ElastiCon ensures predictable performance even under highly dynamic workloads.

Migrating a switch from one controller to another in a naive fashion can cause disruption to ongoing flows, which can severely impact the various applications running in the data center. Unfortunately, the current de facto SDN standard, OpenFlow does not provide a disruption-free migration operation natively. To address this shortcoming, we propose a new *4-phase migration protocol* that ensures minimal disruption to ongoing flows. Our protocol makes minimal assumptions about the switch architecture and is OpenFlow standard compliant. The basic idea in our protocol involves creating a single trigger event that can help determine the exact moment of handoff between the first controller and second controller. We exploit OpenFlow's "equal mode" semantics to ensure such a single trigger event to be sent to both the controllers that can allow the controllers to perform the handoff in a disruption-free manner without safety or liveness concerns.

Armed with this disruption-free migration primitive, ElastiCon supports the following three main load adaptation operations: First, it monitors the load on all controllers and periodically *load balances* the controllers by optimizing the switch-to-controller mapping. Second, if the aggregate load exceeds the maximum capacity of existing controllers, it *grows* the resource pool by adding new controllers, triggering switch migrations to utilize the new controller resource. Similarly, when the load falls below a particular level, it *shrinks* the resource pool accordingly to consolidate switches onto fewer controllers. For all these actions, ElastiCon uses simple algorithms to decide when and what switches to migrate.

The contributions of this chapter are as follows:

- We propose a migration protocol to guarantee safety, liveness and serializability. We show how these guarantees simplify application-specific modifications for moving state between controllers during switch migration. The serializability guarantee requires buffering messages from the switch during migration. This impacts worst-case message processing delay. Hence, we also explore the trade-off between performance and consistency.
- We propose new algorithms for deciding when to grow or shrink the controller resource pool, and trigger load balancing actions.
- We demonstrate the feasibility of ElastiCon by implementing the enhanced migration protocol and proposed algorithms. We address a practical concern of redirecting switch connections to new controllers when the controller pool is grown or away from controllers when the controller pool needs to be shrunk.
- We show that ElastiCon can ensure that performance remains stable and predictable even under highly dynamic traffic conditions.

4.1 Background and Motivation

The OpenFlow network consists of both switches and a central controller. A switch forwards packets according to *rules* stored in its flow table. The central controller controls each switch by setting up the rules. Multiple application modules can run on top of the core controller module to implement different control logics and network functions. Packet processing rules can be installed in switches either reactively (when a new flow is arrived) or proactively (controller installs rules beforehand). We focus on the performance of the reactive mode in this dissertation. Although proactive rule setup (e.g., DevoFlow [41]) can reduce controller load and flow setup time, it is not often sufficient by itself as only a small number of rules can be cached at switches, because TCAM table sizes in commodity switches tend to be small for cost and power reasons. Reactive mode allows the controller to be aware of the lifetime of each flow from setup to teardown, and hence can potentially offer better visibility than proactive mode. For low-end switches, TCAM space is a major constraint. It may be difficult to install all fine-grained microflow policies proactively. Reactive rule insertion allows such rules to be installed selectively and hence may reduce the TCAM size requirement. Thus, it is important to design the controller for predictable performance irrespective of the traffic dynamics.

Switch–controller communication. The OpenFlow protocol defines the interface and message format between a controller and a switch. When a flow arrives at a switch and does not match any rule in the flow table, the switch buffers the packet and sends a `Packet-In` message to the controller. The `Packet-In` message contains the incoming port number, packet headers and the buffer ID where the packet is stored. The controller may respond with a `Packet-Out` message which contains the buffer ID of the corresponding `Packet-In` message and a set of actions (drop, forward, etc.). For handling subsequent packets of the same flow, the controller may send a `Flow-Mod` message with an add command to instruct the switch to insert rules into its flow table. The rules match the subsequent packets of the same flow and hence allow the packets to be processed at line speed. Controller can also delete rules at a switch by using `Flow-Mod` with delete command. When a rule is deleted either explicitly or due to timeout, the switch sends a `Flow-Removed` message to the controller if the "notification" flag for the flow is set. In general, there is no guarantee on the order of processing of controller messages at a switch. Barrier messages are used to solve the synchronization problem. When the switch receives a `Barrier-Request` message from the controller, it sends a `Barrier-Reply` message back to the controller only after it has finished processing all the messages that it received before the `Barrier-Request`.

Controller architecture. The controller architecture has evolved from the original single-threaded design [42] to the more advanced multi-threaded design [43–46] in recent years. Despite the significant performance improvement over time, the single-controller systems still have limits on scalability and vulnerability. Some research papers have also explored the implementation of distributed controllers across multiple hosts [8–10]. The main focus of these papers is to address the state consistency issues across distributed controller instances, while preserving good performance. Onix, for instance, uses a transactional database for persistent but less dynamic data, and memory-only DHT for data that changes quickly but does not require consistency [8]. Hyperflow replicates the events at all distributed nodes, so that each node can process such events and update their local state [9]. [10] has further elaborated the state distribution trade-offs in SDNs. OpenDaylight [47] is a recent open source distributed SDN controller. Like ElastiCon, it uses a distributed data store for storing state information.

All existing distributed controller designs implicitly assume static mapping between switches and controllers, and hence lack the capability of dynamic load adaptation and elasticity. However, the following back-of-the-envelope calculation using real measurement data shows that there is 1-2 orders of magnitude difference between peak and median flow arrival rates at a switch. In [11], Benson *et al.* show that the minimum inter flow arrival gap is $10\mu s$, while the median ranges roughly from $300\mu s$

to $2ms$ across different data centers that they have measured. Assuming a data center with 100K hosts and 32 hosts/rack, peak flow arrival rate can be up to $300M$ with the median rate between 1.5M and 10M. Assuming 2M packets/sec throughput[2] for one controller [43], it requires only 1-5 controllers to process the median load, but 150 for peak load. If we use static mapping between switches and controllers and install all flow table entries reactively, it requires significant over-provisioning of resources which is inefficient in hardware and power; an elastic controller that can dynamically adapt to traffic load is clearly more desirable.

4.2 Elastic Controller Design

We present the design and architecture of ElastiCon, an elastic distributed SDN controller in this section. We describe the architecture of ElastiCon in three phases: First, we start with a basic distributed controller design that spreads functionality across several nodes by extending Floodlight, a Java-based open source controller [46]. We then describe the 4-phase protocol for disruption-free switch migration, which is one of the core primitives needed for implementing an elastic controller. Finally, we discuss the algorithms we use for elasticity and load adaptation in our design.

4.2.1 Basic Distributed Controller

The key components in our distributed controller design are shown in Figure 4.1. It consists of a cluster of autonomous *controller nodes* that coordinate amongst themselves to provide a consistent control logic for the entire network. The *physical network infrastructure* refers to the switches and links that carry data and control plane traffic. Note that, for simplicity, we have omitted showing the physical topology of the network that includes the hosts and their interconnections with the switches in the network.

[2]This is based on the learning switch application. Throughput is lower for more complex applications, as shown in our experiments.

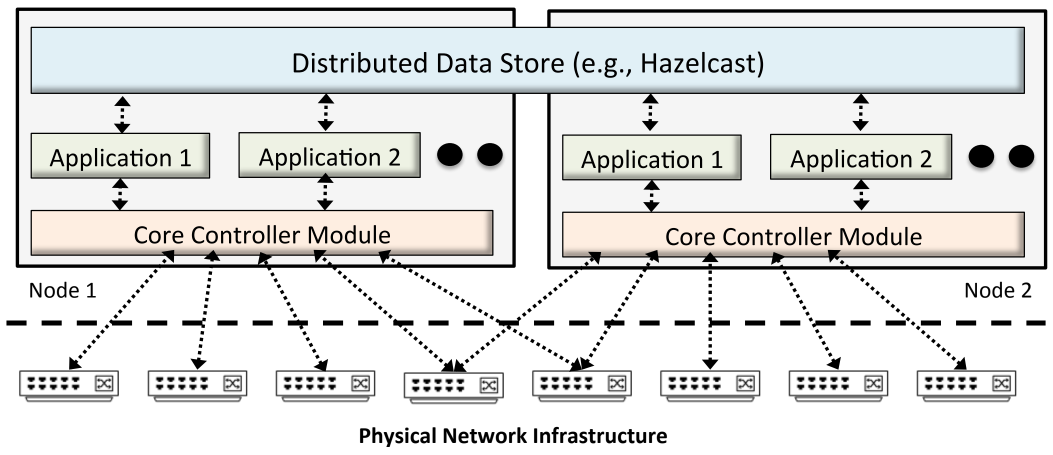


Figure 4.1.: Basic distributed controller architecture.

Typically, each switch connects to one controller. However, for fault-tolerance purposes, it may be connected to more than one controller with one master and the rest as slaves. We assume the control plane is logically isolated from the data plane, and the control plane traffic is not affected by data plane traffic. Each controller node has a *core controller module* that executes all the functions of a centralized controller (i.e., connecting to a switch, event management between a switch and an application). In addition, it coordinates with other controllers to elect a master node for a newly connected switch and orchestrates the migration of a switch to a different controller.

The *distributed data store* provides the glue across the cluster of controllers to enable a logically centralized controller. It stores all switch-specific information that is shared among the controllers. Each controller node also maintains a TCP connection with every other controller node in the form of a full mesh. This full-mesh topology is mainly for simplicity, but as the number of controllers become exceedingly large, one may consider adding a point of indirection, similar to the route-reflector idea in scaling BGP connections in ISP networks. For today's data centers, maintaining a full mesh across a few 100 controllers does not pose any scaling concerns. A controller node uses this TCP connection for various controller-to-controller messages, such as when sending messages to a switch controlled by another node or coordinating actions during switch migration. The *application* module implements the control logic of network applications, responsible for controlling the switches for which its controller is the master. The fact that state is maintained distributed data store makes switch migration easier and also helps fast recovery from controller failures.

4.2.2 4-Phase Switch Migration Protocol

If we use a single SDN controller, since all switches are always connected to this controller, there is no break in the control plane processing. Moving to a distributed controller architecture does not necessarily pose a problem so long as the switch-to-controller mapping stays static. However, such an architecture, which is employed by previously proposed distributed controllers, cannot adapt to the load imbalances caused by spatial and temporal variations in traffic conditions. Once a controller becomes overloaded, the response time for control plane messages becomes too high, thus impacting flows and applications running in the data center. We can mitigate such imbalances by dynamically shifting load between existing controllers or by adding new nodes to the controller pool. The basic granularity at which one can shift load is at a switch-level; simply migrate a switch from an overloaded controller to a lightly loaded one.

Unfortunately, there is no native support for safely migrating switches in existing de facto SDN standard, OpenFlow, without which one cannot guarantee that there is no impact to traffic during migration. In particular, there are three standard properties any migration protocol needs to provide—*liveness*, *safety* and *serializability.*

- **Liveness.** At least one controller is active for a switch at all times. Otherwise, a new flow that arrives at a switch cannot be properly routed causing disruption to that application. In addition, if a controller has issued a command to a switch, it needs to remain active until the switch finishes processing that command.
- **Safety.** Exactly one controller processes every asynchronous message from the switch; duplicate processing of asynchronous messages such as `Packet-In` could

result in duplicate entries in the flow table, or even worse, inconsistency in the distributed data store.

- **Serializability.** The controller processes events in the order in which they are transmitted by the switch; if events are processed in a different order, the controller's view of the network may be inconsistent with the state of the network. For instance, if a link goes down and comes back up, the switch will generate a port status down message followed by a port status up. However, if these events are processed in the wrong order, the controller may assume that the link is permanently down.

Now, consider the following naive protocol that OpenFlow readily provides: The target controller can be first put in the slave mode for the switch (see Section 4 for implementation details). The target controller then simply sends a `Role-Request` message to the switch indicating that it wants to become the master. The switch would set that controller as the master and the previous master as slave. Such a naive and intuitive protocol can cause serious disruption to traffic since it can violate the liveness property. Assume that the switch had sent a `Packet-In` message to the initial master. If the switch receives the `Role-Request` message from the slave before the `Packet-Out` message from the initial master, then the switch will ignore the `Packet-Out` message since it is designed to ignore messages from any controller which is not the master/equal. Ideally, the switch can buffer all these `Packet-In` requests and try retransmitting the `Packet-In` message to the new master, but that makes the switch design complicated, which is not desirable.

In our protocol design, we assume we cannot modify the switch. There are two additional issues: First, the OpenFlow standard clearly states that a switch may process messages not necessarily in the order they are received, mainly to allow multi-threaded implementations. We need to factor this in our protocol design. Second, the standard does not specify explicitly whether the ordering of messages transmitted by the switch remains consistent across two controllers that are in master/equal mode. This assumption, which is clearly logical, is required for our protocol to work; allowing

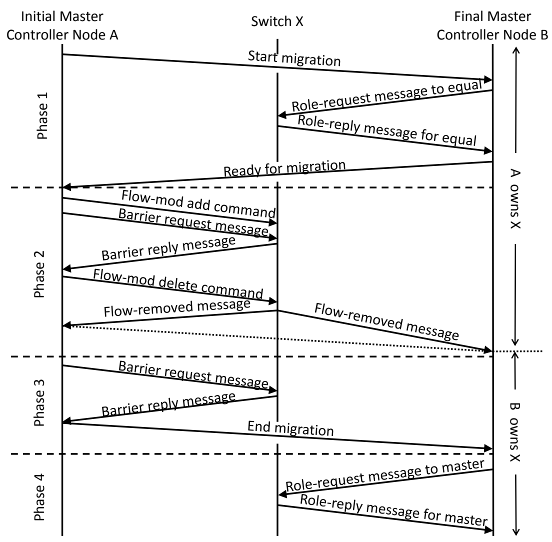


Figure 4.2.: Message exchanges for switch migration.

arbitrary reordering of messages across two controllers will make an already hard problem significantly harder. For ease of exposition, we use X to denote the switch, which is being migrated from initial controller A to target controller B. We first outline the key ideas that provide the desired guarantees and then describe the protocol in detail.

Liveness. To guarantee liveness, we first transition the target controller B to equal mode. After that, we transition initial controller A from master to slave mode and then transition controller B to master mode. This ensures guarantees liveness since at least one controller is active (master or equal mode) at a time.

Safety. Using an intermediate equal mode for the controller B solves the liveness problem but it may violate the safety property since both controllers may process messages from the switch causing inconsistencies and duplicate messages. To guarantee safety, we create a *single trigger event* to stop message processing in the first controller and start the same in the second one. Fortunately, we can exploit the fact that `Flow-Removed` messages are transmitted to all controllers operating in the equal mode. We therefore simply insert a dummy flow entry into the switch and then remove the flow entry, which will provide a single `Flow-Removed` event to both the controllers to signal handoff.

Serializability. To guarantee serializability, the controller A should complete processing its last message before the controller B can process its first message. However, the first message for the B may arrive before A completes processing its last message. So, we cache messages at B until the A has finished processing its last message and committed it to the switch.

Our protocol operates in four phases described below (shown in Figure 4.2). We now describe each phase in detail and highlight a trade-off between performance and serializability.

Phase 1: Change role of the target to equal. In the first phase, target B's role is first changed to equal mode for the switch X. Initial master A initiates this phase by sending a start migration message to B using a proprietary message on the controller-to-controller channel. B sends `Role-Request` message to the switch informing that it is an equal. After B receives a `Role-Reply` message from the switch, it informs the initial master A that its role change is completed. After B changes its role to equal, it receives control messages (e.g., `Packet-In`) from the switch, but ignores them and does not respond.

Phase 2: Insert and remove a dummy flow. To determine an exact instant for the migration, A sends a dummy (but well-known) `Flow-Mod` command to X to add a new flow table entry that does not match any incoming packet. Then, it sends another `Flow-Mod` command to delete this flow table entry; in response, the switch sends a `Flow-Removed` message to both controllers since B is in the equal mode. This `Flow-Removed` event signals a handoff of switch X from A to B, and henceforth, only B will process all messages transmitted by switch. Here, our assumption that both controllers in equal mode receive messages from the switch in the same order is needed to guarantee the safety property. An additional barrier message is required after the insertion of the dummy flow and before the dummy flow is deleted to prevent any chance of processing the delete message before the insert.

Although B processes all messages after the `Flow-Removed` message, it does not do so immediately. It caches all the messages after the `Flow-Removed` message and begins processing them in the next phase. This is needed to guarantee the serializability property. Processing of messages from the north-bound interface can continue uninterrupted.

Phase 3: Flush pending requests with a barrier. Now, B has taken over responsibility of switch X, but A has not detached from X yet. However, it cannot just detach immediately from the switch since there may be pending requests at A that arrived before the `Flow-Removed` message, for which A is still the owner. Controller A processes all messages that arrived before `Flow-Removed` and transmits their responses. Then, it transmits a `Barrier-Request` and waits for the `Barrier-Reply`. Receiving a `Barrier-Reply` from switch X indicates that X has finished processing all messages that it received before the `Barrier-Request` messages. So, only after receiving the `Barrier-Reply` message, controller A signals "end migration" to the final master B. The "end migration" message is a signal to B that A has finished processing all its messages and committed them to the switch. Once B receives the "end migration" message, it processes all the cached messages in the order that they were received. Note that delay in end migration message can potentially cause message processing latency at B. This delay can be avoided if we do not need to guarantee serializability. In that case B can start processing `Packet-In` messages right after receiving `Flow-Removed`.

Phase 4: Make target controller final master. Here, A would have already detached from X and has signaled to B to become the new master, which it does by by sending a `Role-Request` message to the switch. It also updates the distributed data store to indicate this. The switch sets A to slave when it receives the `Role-Request` message from the final master B after which it processes all messages from the switch.

Performance-Serializability Trade-off. Buffering messages from the switch at the end of phase 2 is needed to guarantee serializability. It ensures that B begins processing messages only after A has completed processing messages before the

`Flow-Removed` message. The duration for buffering messages will depend on the network latencies, message loss ratio, controller processing times, etc. In our experiments, we observed that messages were never buffered for more than 50msec. However, the worst case will depend on many network characteristics and may be larger. While buffering is needed to guarantee serializability, it has two undesirable side-effects. First, the controller will be unable to respond to events from the switch while messages are being buffered. Second, buffered messages will be processed late and may be irrelevant by the time they are processed. So, the network operator should choose between two configurations of the migration protocol depending on network characteristics and application requirements. The "consistency configuration" buffers messages as described above and provides all three guarantees. The "performance configuration" does not buffer messages. It does not provide serializability but responds faster to switch events during migration.

4.2.3 Application State Migration

Safety, liveness and serializability guarantees of the migration protocol simplify controller application changes to support switch migration. The three guarantees together ensure that applications do not miss any asynchronous events and do not have to check for duplicate or reordered asynchronous messages from the switch before processing them. We describe the modifications to the applications and their interface with the core controller module below. We have implemented them for the routing applications in ElastiCon.

We added two methods to the interface between the core controller module and each application module. The first method, named "switch_emigrate", is invoked at the initial master controller (controller A in the above example). The core controller module invokes this method for each application after it has finished processing all messages before the `Flow-Removed` message from the switch. The method returns after the application has flushed all switch-specific state to the distributed

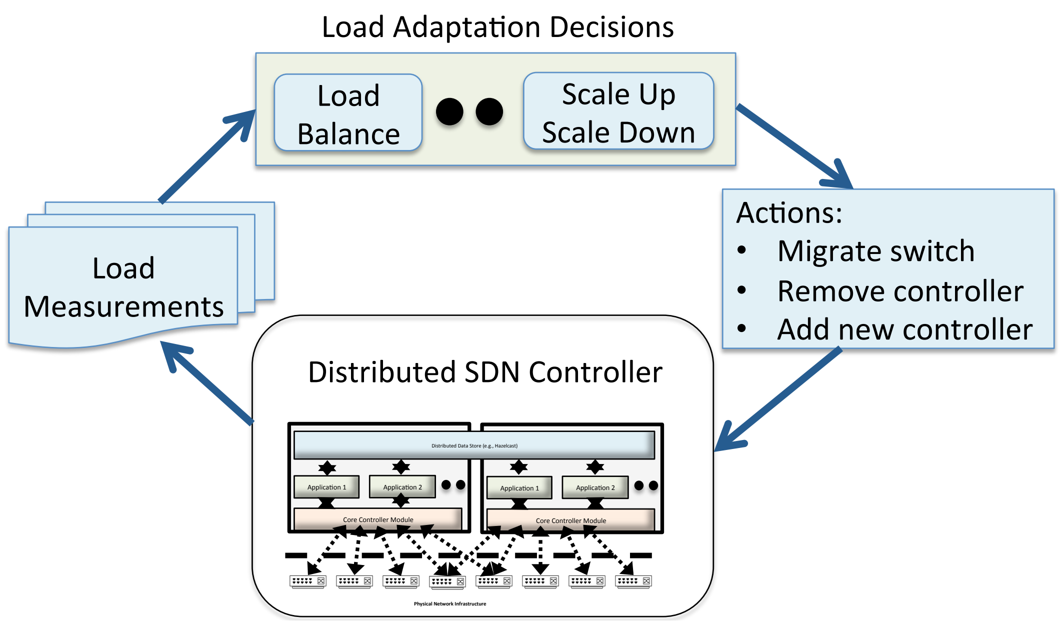


Figure 4.3.: Load adaptation in ElastiCon.

data store. Applications also stop any switch-specific execution (like timers). The controller sends the “end migration” message only after all applications execute their “switch_emigrate” method. The second method, “switch_immigrate”, is invoked at the target master controller (controller B in the above example) for each application. It is invoked after the controller receives the “end migration” message. Each application reads switch-specific state from the distributed data store to populate local data structures and starts switch-specific execution. The distributed data store should guarantee that the controller reads the state written in the “switch_emigrate” method earlier. The controller starts processing cached asynchronous messages after all applications have executed their “switch_immigrate” methods.

State-transfer between applications can also be performed over TCP connections between applications instead of using the distributed data store. The above design simplified our implementation since we reused the interface between the application and the distributed data store. Using this disruption-free migration protocol as a basic primitive, we now look at load adaptation aspects of ElastiCon.

4.2.4 Load Adaptation

There are three key operations we envision for load adaptation in ElastiCon. If the aggregate traffic load is greater (smaller) than aggregate controller capacity, we need to *scale up (down)* the controller pool. In addition, we need to periodically *load balance* the controllers by migrating switches to newer controllers to adapt to traffic load imbalances. We show our basic approach to achieve this in Figure 4.3. It consists of three steps:

- Periodically collect load measurements at each controller node.
- Determine if the current number of controller nodes is sufficient to handle the current load. If not, add or remove controller nodes. In addition, if any controller is getting overloaded, but aggregate load is within the capacity, we need to trigger load balancing actions.
- Finally, adjust the switch to controller mapping by adding or removing the controllers and triggering switch migrations as needed.

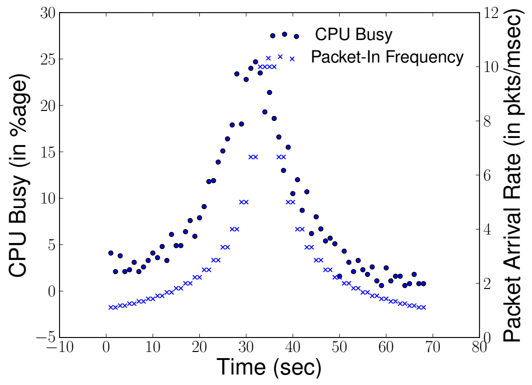


Figure 4.4.: CPU vs. packet frequency.

Load Measurement

The most direct way to measure the load on a controller is by sampling response time of the controller at the switches. This response time will include both computation and network latency. However, switches may not support response time measurements, since that requires maintaining some amount of extra state at the switches that may or may not be feasible. Since the controller is more programmable, ElastiCon maintains a load measurement module on each controller to periodically report the CPU utilization and network I/O rates at the controller. Our experiments show that the CPU is typically the throughput bottleneck and CPU load is roughly in proportion to the message rate (see Figure 4.4). The module also reports the average message arrival rate from each switch connected to the controller. This aids the load balancer in first dissecting the contribution of each switch to the overall CPU utilization, and helps making optimal switch to controller mapping decisions. We assume that the fraction of controller resources used by a switch is proportional to its fraction of the total messages received at the controller, which is typically true due to the almost linear relationship between throughput and messages. The load measurement module averages load estimates over small time intervals (we use three seconds) to avoid triggering switch migrations due to short-term load spikes.

Algorithm 1 Load Adaptation Algorithm

```
while True do
    GET_INPUTS()
    $migration\_set \leftarrow$ DOREBALANCING()
    if $migration\_set == NULL$ then
        if DORESIZING() then
            if CHECKRESIZING() then
                $migration\_set \leftarrow$ DOREBALANCING()
            else
                REVERTRESIZING()
            end if
        end if
    end if
    EXECUTE_POWER_ON_CONTROLLER()
    EXECUTE_MIGRATIONS($migration\_set$)
    EXECUTE_POWER_OFF_CONTROLLER()
    SLEEP(3)
end while
```

Adaptation Decision Computation

The load adaptation algorithm determines if the current distributed controller pool is sufficient to handle the current network load. It sets a high and low thresholds to determine whether the distributed controller needs to be scaled up or down. Difference between these thresholds should be large enough to prevent frequent scale changes. Then, the algorithm finds an optimal switch to controller mapping constrained by the controller capacity while minimizing the number of switch migrations. Some CPU cycles and network bandwidth should also be reserved for switches connected to a controller in slave mode. Switches in slave mode impose very little load on the controller typically, but some headroom should be reserved to allow switch migrations.

While one can formulate and solve an optimization problem (e.g., linear program) that can generate an optimal assignment of switch-to-controller mappings, it is not clear such formulations are useful for our setting in practice. First, optimal balancing is not the primary objective as much as performance (e.g., in the form of response time). Usually, as long as a controller is not too overloaded, there is not much performance difference between different CPU utilization values. For example, 10% and 20% CPU utilization results in almost similar controller response time. Thus, fine-grained optimization is not critical in practice. Second, optimal balancing may result in too many migrations that is not desirable. Of course, one can factor this in the cost function, but then it requires another (artificial) weighting of these two functions, which then becomes somewhat arbitrary. Finally, optimization problems are also computationally intensive and since the traffic changes quickly, the benefits of the optimized switch-controller mapping are short-lived. So, a computationally light-weight algorithm that can be run frequently is likely to have at least similar if not better performance than optimization. Perhaps, this is the main reason why distributed resource management (DRM) algorithms used in real world for load balancing cluster workloads by migrating virtual machines (VMs) do not solve any such optimization problems and rely on a more simpler feedback loop [15]. We adopt a similar approach in our setting.

Our load-adaptation decision process proceeds in two phases, as shown in Algorithm 1. First, during the rebalancing step the load adaptation module evenly distributes the current load across all available controllers. After rebalancing, if the load on one or more controllers exceeds the upper (or lower) threshold, the load adaptation module grows (or shrinks) the controller pool.

Input to the Algorithm. A load adaptation module within ElastiCon periodically receives inputs from the load measurement module on each controller. The input contains the total CPU usage by the controller process in MHz. It also contains a count of the number of packets received from each switch of which that controller is the master. The packet count is used to estimate the fraction of the load on the controller due to a particular switch. The load adaptation module stores a moving window of the past inputs for each controller. We define utilization of a controller as the sum of the mean and standard deviation of CPU usage over the stored values for that controller. The rebalancing and resizing algorithms never use instantaneous CPU load. Instead they use CPU utilization to ensure that they always leave some headroom for temporal spikes in instantaneous CPU load. Also, the amount of headroom at a controller will be correlated to the variation in CPU load for that controller.

Output of the Algorithm. After processing the inputs, the load adaptation module may perform one or more of the following actions: powering off a controller, powering on a controller, or migrating a switch from one controller to another.

Main Loop of the Algorithm. First, the load adaptation module receives the inputs from all controllers and augments them to its stored state. All functions except the EXECUTE_* functions only modify this stored state and they do not affect the state of the controllers. After that, the EXECUTE_* functions determine the changes to the stored state and send migration and power on/off commands to the appropriate controllers.

There are two main subroutines in the rebalancing algorithm: doRebalancing and doResizing. doRebalancing distributes the current load evenly among the controllers. doResizing adds or removes controllers accordingly to the current load. doResizing is invoked after doRebalancing since resizing the controller pool is a more intrusive operation than rebalancing the controller load, and hence should be avoided when possible. Although one can estimate average load per controller without actually doing rebalancing and then determine whether resizing is needed or not, this often suffers from estimation errors.

If the first invocation of doRebalancing generates any migrations, we execute those migrations and iterate over the main loop again after 3 seconds. If there are no migrations (indicating that the controllers are evenly loaded), ElastiCon generates resizing (i.e., controller power on/off) decisions by invoking doResizing. The power off decision needs to be verified to ensure that the switches connected to the powered off controller can be redistributed among the remaining controllers without overloading any one of them. This is done in the checkResizing function. This function uses a simple first-fit algorithm to redistribute the switches. While other more sophisticated functions can be used, our experience indicates first-fit is quite effective most of the time. If this function fails, the (stored) network state is reverted. Otherwise, ElastiCon calls doRebalancing to evenly distribute the switch load. Finally, the execute_* functions implement the state changes made to the network by the previous function calls. Since a migration changes the load of two controllers, all stored inputs for the controllers involved in a migration are discarded. The main loop is executed every 3 seconds to allow for decisions from the previous iteration to take effect.

Rebalancing. The rebalancing algorithm, described in Algorithm 2, tries to balance the average utilization of all controllers. We use the standard deviation of utilization across all the controllers as a balancing metric. In each iteration, it calls the get_best_migration function to identify the migration that leads to the most reduction in standard deviation of utilization across controllers. This function tries every possible migration in the network and estimates the standard deviation of uti-

Algorithm 2 The rebalancing algorithm

```
procedure doRebalancing()
    $migration\_set \leftarrow NULL$
    while True do
        $best\_migration \leftarrow$ GET_BEST_MIGRATION()
        if $best\_migration.std\_dev\_improvement \geq THRESHOLD$ then
            $migration\_set$.INSERT($best\_migration$)
        else
            return $migration\_set$
        end if
    end while
end procedure
```

lization for each scenario. It returns the migration which has the smallest estimated standard deviation. To estimate the standard deviation, this function needs to know the load imposed by every switch on its master controller. Within each scenario, after a hypothetical migration, the function calculates the utilization of each controller by adding the fractional utilizations due to the switches connected to it. It then finds the standard deviation across the utilization of all the controllers. If reduction in standard deviation by the best migration it finds exceeds the minimum reduction threshold, ElastiCon adds that migration to the set of migrations. If no such migration is found or the best migration does not lead to sufficient reduction in standard deviation, it exits.

Resizing. The resizing algorithm, shown in Algorithm 3, tries to keep the utilization of every controller between two preset high and low thresholds. Each invocation of the resizing algorithm generates either a power on, or power off, or no decision at all. The resizing algorithm is conservative in generating decisions to prevent oscillations. Also, it is more aggressive in power on decisions than power off. This is because when the utilization exceeds the high threshold, the network performance may suffer unless additional controllers are put in place quickly. However, when the utilization goes below the low threshold, network performance does not suffer. Removing controllers only consolidates the workload over fewer controllers sufficient to

Algorithm 3 The resizing algorithm

```
procedure doResizing()
    for all $c$ $in$ $controller\_list$ do
        if $c.util \geq HIGH\_UTIL\_THRESH$ then
            switch_on_controller()
            return True
        end if
    end for
    $counter \leftarrow 0$
    for all $c$ $in$ $controller\_list$ do
        if $c.util \leq LOW\_UTIL\_THRESH$ then
            $counter \leftarrow counter + 1$
        end if
    end for
    if $counter \geq 2$ then
        switch_off_controller()
        return True
    else
        return False
    end if
end procedure
```

handle existing traffic conditions, mainly for power and other secondary concerns than network performance. Thus, we generate a power on decision when any controller exceeds the high threshold while requiring at least two controllers to fall below the low threshold for generating a power off decision. Triggering a decision when just one or two controllers cross the threshold might seem like we aggressively add or remove controller. But, our decisions are quite conservative because the resizing algorithm is executed only when the load is evenly distributed across all controllers. So, if a controller crosses the threshold, it indicates that all controllers are close to the threshold.

Extending Load Adaptation Algorithms

The load adaptation algorithms described above can be easily extended to satisfy additional requirements or constraints. Here we describe two such potential extensions to show the broad applicability and flexibility of the algorithm.

Controller Location. To reduce control plane latency, it may be better to assign a switch to a closeby controller. We can accommodate this requirement in ElastiCon by contraining migrations and controller additions and removals. To do so, in every iteration of the rebalancing algorithm (Algorithm 2), we consider only migrations to controllers close to the switch. This distance can be estimated based on topology information or controller to switch latency measurements. If the operator wants to set switch-controller mapping based on physical distance (in number of hops), he/she can use the network topology. The operator should use latency measurements when he/she wants to set switch-controller mapping based on logical distance (in milliseconds). Similarly, in the resizing algorithm (Algorithm 3), the new controllers added should be close to the overloaded controllers so that switches can migrate away from the overloaded controller. The first-fit algorithm used in CHECKRESIZING function should also be modified such that a switch can only "fit" in a closeby controller.

Switch Grouping. Assigning neighboring switches to the same controller may reduce inter-controller communication during flow setup and hence improve control plane efficiency. Graph partitioning algorithms can be used to partition the network into switch groups; and the result can be fed into ElastiCon. ElastiCon can be modified to treat each group as a single entity during migration and resizing, so that the switches of the same group are always controlled by the same controller except for short intervals during migration. The load measurements module should be modified to combine load readings of switches of a group and present it as a single entity to the load adaptation algorithm. When the rebalancing algorithm determines that the entity needs to be migrated, the EXECUTE_* functions should migrate all the switches of the group.

Adaptation Action

Following the adaptation decision, adaptation actions are executed to transform the network configuration (i.e., switch to controller mapping). A switch is migrated to a former slave by following the steps in our 4-phase migration protocol described before. In case of controller addition or removal, one or more switches may need to be reassigned to new master controllers that they are not currently connected to. This can be done by replacing one of the existing slave controllers' IP address of the switch with that of the new controller using the `edit-config` operation of OpenFlow Management and Configuration Protocol [48]. Once the connection between the new controller and the switch is established, we then invoke the migration procedure to swap the old master with the new slave controller. If a switch does not support updating controller IP addresses at runtime, other workarounds based on controller IP address virtualization are also possible (discussed in Section 4.3).

4.3 Implementation

In this section, we present further details on how we implement ElastiCon by modifying and adding components to the centralized Floodlight controller.

Distributed Data Store. We use Hazelcast to implement the distributed data store. Although other NoSql databases may have also worked here, we find Hazelcast a good choice due to its performance and flexibility. Hazelcast provides strong consistency, transaction support, and event notifications. Its in-memory data storage and distributed architecture ensures both low latency data access and high availability. Persistent data can be configured to write to disk. It is written in Java, which makes it easy for integration with Floodlight. We include the Hazelcast libraries in the Floodlight executable. The first Hazelcast node forms a new distributed data store. Subsequently, each Hazelcast node is configured with the IP addresses and ports of several peers. At least one of the peers needs to be active for the new node to join the distributed data store.

Controller. When a controller boots up, it publishes its own local data and retrieves data of other controllers by accessing Hazelcast. One such data is the IP address and TCP port of each controller needed for inter-controller communication. This allows the controllers to set up direct TCP connections with each other, so that they can invoke each other to set up paths for flows.

The switch to master controller mapping is also stored in Hazelcast using the unique switch datapath-id as the key. We have modified the core controller in Floodlight to allow a controller to act in different roles for different switches. The initial switch to master mapping can be determined in one of two ways. In the first method, the load adapter module running in the controller (described later) reads in the mapping from a configuration file and stores the information in Hazelcast. We also implement an ad hoc strategy by letting the controllers try to acquire a lock in Hazelcast when a switch connects to them. Only one controller can succeed in acquiring the lock; it then declares itself as the master for the switch.

Load Adaptation Module. The load measurement module is integrated into the controller. We use SIGAR API [49] to retrieve the CPU usage of the controller process. We enhanced the REST API of the controller to include CPU usage queries. The adaptation decision algorithm run on a separate host. It communicates with all controllers over the REST API. It requires the REST port and IP address of one of the controllers. Using that, it queries the controller for the IP address and REST port of all other controllers and switch-to-controller mappings of all switches in the network. In each iteration, the program queries the CPU usage information from each controller and sends migration requests to the master controller of a switch when the switch needs to be migrated.

Adding and Removing Controllers. Migration of a switch to a newly connected controller is done in two steps. First, we replace the IP address and TCP port number of one of the slave controllers of the switch with those of the new controller. This can be done by using the `edit-config` operation of OpenFlow Management and Configuration Protocol [48]. Once the connection between the new controller and the switch is established, we then invoke the migration procedure to swap the old master with the new slave controller.

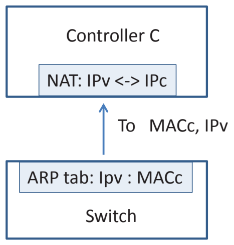


Figure 4.5.: Controller virtual IP address binding

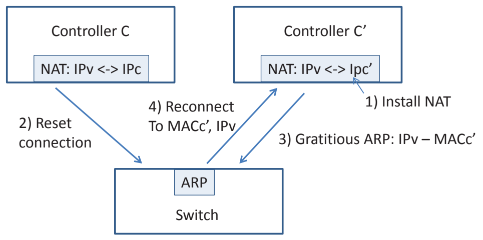
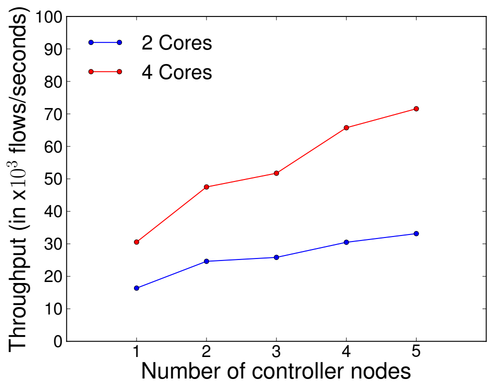
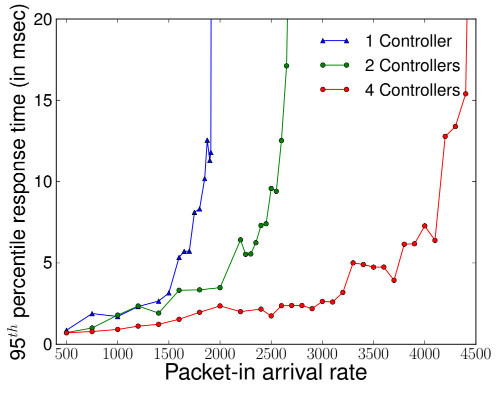


Figure 4.6.: Controller binding change

If a switch does not support updating controller IP addresses at runtime, we can use the following procedure as a workaround, which is suitable when the control plane is configured to use the same layer 2 network (e.g., on the same VLAN). All switches are configured to use a set of virtual controller IP addresses, which will be mapped to the real controller IP addresses at runtime according to load condition. Such mapping can be realized by using ARP and Network Address Translation (NAT), as shown in Figure 4.5. When the virtual controller IP address ip_v for a switch is mapped to controller C's IP address ip_c, we use gratuitous ARP to bind the MAC address of the controller C with ip_v, so that the packets to ip_v can reach controller C. At controller C, we do NAT from ip_v to ip_c, so that the packets can be handled by the controller transparently.



(a) Controller throughput.



(b) Response time.

Figure 4.7.: Performance with varying number of controller nodes.

Figure 4.6 shows how such binding can be changed when we need to replace controller C with controller C'. We first send a TCP reset message from C to disconnect the switch from the controller, and then use gratuitous ARP to bind MAC address of C' with ip_v. Note that connection reset to C is only done when C is not a master controller to avoid disruption in normal switch operation. When the switch tries to reconnect to ip_v, the message will reach C' instead of C. We then do a NAT from ip_v to $ip_{c'}$ at controller C' as before. Note that if the gratuitous ARP does not reach the switch before the reconnection request is sent, controller C simply rejects the reconnection request and the switch ultimately gets connected to controller C'.

4.4 Evaluation

In this section, we evaluate the performance of our ElastiCon prototype using an emulated SDN-based data center network testbed. We first describe the enhanced Mininet testbed that we used to carry out the evaluation, and then present our experimental results.

4.4.1 Enhanced Mininet Testbed

Our experimental testbed is built on top of Mininet [50], which emulates a network of Open vSwitches [51]. Open vSwitch is a software-based virtual Openflow switch. It implements the data plane in kernel and the control plane as a user space process. Mininet has been widely used to demonstrate the functionalities, but not the performance, of a controller because of the overhead of emulating data flows. First, actual packets need to be exchanged between the vSwitch instances to emulate packet flows. Second, a flow arrival resulting in sending a `Packet-In` to the controller incurs kernel to user space context switch overhead in the Open vSwitch. From our initial experiments we observe that these overheads significantly reduce the maximum flow arrival rate that Mininet can emulate, which in turn slows down the control plane traffic generation capability of the testbed. Note that for the evaluation of ElastiCon, we are primarily concerned with the control plane traffic load and need not emulate the high overhead data plane. We achieve this by modifying Open vSwitch to inject `Packet-In` messages to the controller without actually transmitting packets on the data plane. We also log and drop `Flow-Mod` messages to avoid the additional overhead of inserting them in the flow table. Although we do not use the data plane during our experiments, we do not disable it. So, the controller generated messages (like LLDPs, ARPs) are still transmitted on the emulated network.

In order to experiment with larger networks we deployed multiple hosts to emulate the testbed. We modified Mininet to run the Open vSwitch instances on different hosts. We created GRE tunnels between the hosts running Open vSwitch instances to emulate links of the data center network. Since we do not actually transmit packets in the emulated network, the latency/bandwidth characteristics of these GRE tunnels do not impact our results. They are used only to transmit link-discovery messages to enable the controllers to construct a network topology. To isolate the switch to controller traffic from the emulated data plane of the network, we run Open vSwitch on hosts with two Ethernet ports. One port of each host is connected to a gigabit

Ethernet switch and is used to carry the emulated data plane traffic. The other port of each host is connected to the hosts that run the controller. We isolated the inter-controller traffic from the controller-switch traffic too by running the controller on dual-port hosts.

4.4.2 Experimental Results

We report on the performance of ElastiCon using the routing application. All experiments are conducted on k=4 fat tree emulated on the testbed. We use 4 hosts to emulate the entire network. Each host emulates a pod and a core switch. Before starting the experiment, the emulated end hosts ping each other so that the routing application can learn the location of all end hosts in the emulated network.

Throughput. We send 10,000 back-to-back `Packet-In` messages and plot the throughput of ElastiCon with varying number of controller nodes (Figure 4.7(a)). We repeat the experiment while pinning the controllers to two cores of the quad-core server. We observe two trends in the results. First, adding controller nodes increases the throughput almost linearly. This is because there is no data sharing between controllers while responding to `Packet-In` messages. Second, the throughput reduces when we restrict the controllers to two cores indicating that CPU is indeed the bottleneck.

Response time. We plot the response time behavior for `Packet-In` messages with changing flow arrival rate (see Figure 4.7(b)). We repeat the experiment while changing the number of controller nodes. As expected, we observe that response time increases marginally up to a certain point. Once the packet generation rate exceeds the capacity of the processor, queuing causes response time to shoot up. This point is reached at a higher packet-generation rate when ElastiCon has more nodes.

Migration time. The time taken to migrate a switch is critical for the load balancing protocol to work efficiently. We define migration time for controller A as the time between sending the "start migration" message and "end migration" message.

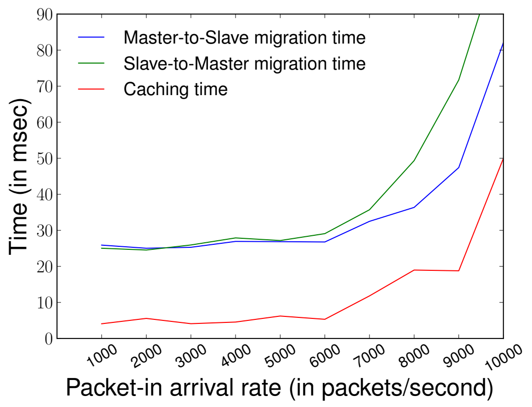
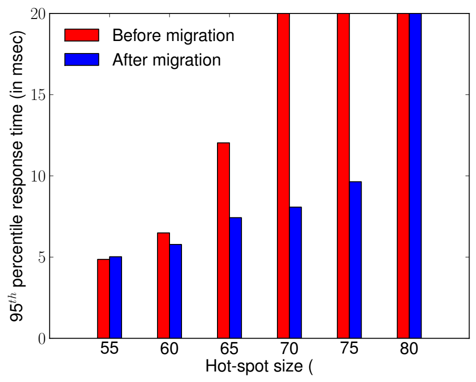


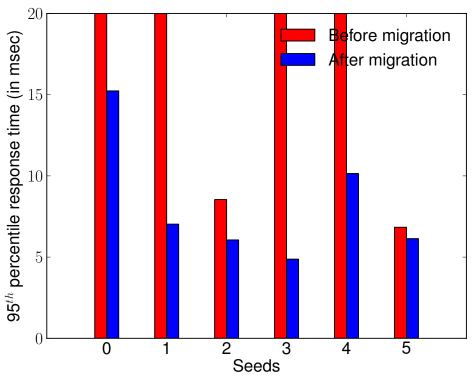
Figure 4.8.: Migration time

We define migration time for controller B as the time between receiving the "start migration" and sending the `Role-Request` to change to master. In a network with 3 controllers, we perform 200 migrations and observe the migration time for each migration at both controllers. We also observe the time for which controller B caches messages from the switch. We plot the 95th percentile of the migration and caching times in Figure 4.8. The plot shows that the migration time is minimal (few tens of milliseconds) and increases marginally as the load on the controller increases. The caching time is even smaller (around 5ms). This keeps memory usage of the message cache small (few KBs).

Automatic rebalancing under hot-spot traffic. We use a N=4 fat tree to evaluate the effect of the automatic load balancing algorithm. Three of the four pods of the fat tree are evenly loaded, while the flow arrival rate in the fourth pod is higher than that in the other three. We configure ElastiCon with four controllers, one assigned to all the switches of each pod. The master controller of the fourth pod is obviously more heavily loaded than the other three. Figure 4.9(a) shows the 95th percentile of the response time of all `Packet-In` messages before and after rebalancing. The `Packet-In` message rate in the fourth pod is varied on the X-axis. We truncate the y-axis at 20ms, so a bar at 20ms is actually much higher.



(a) Single hot-spot traffic pattern.



(b) Pareto distributed traffic pattern.

Figure 4.9.: Benefit of automatic rebalancing. We truncate the y-axis at 20ms, so a bar at 20ms is actually much higher.

We observe that as traffic gets more skewed (i.e., the `Packet-In` rate in the fourth pod increases), we see a larger benefit by doing rebalancing corresponding to the 65-75% bars. At 70-80% hot-spot, the system is unstable. The 95th percentile can be arbitrarily high depending on the amount of time the experiment is run before rebalancing, since the one of the controllers is overloaded (i.e., the `Packet-In` rate exceeds the saturation throughput). At 80% hot-spot, rebalancing by itself does not help as seen by the blue bar exceeding 20ms since there is no way to fit the workload among existing controllers.

Automatic rebalancing under Pareto distribution. We also evaluate the benefit of the rebalancing algorithm in the case where multiple hot spots may appear randomly following a Pareto distribution. As before, we use a N=4 fat tree with 4 controllers. The network generates 24,000 `Packet-In` messages per second. The message arrival rate is distributed across all the switches in the network using a Pareto distribution. We repeat the traffic pattern with 6 different seeds. We start with a random assignment of switches to controllers and apply the rebalancing algorithm. Figure 4.9(b) shows the 95th percentile response time with random assignment and with rebalancing.

Since a Pareto distribution is highly skewed, the improvement varies widely depending on the seed. If the distribution generated by a seed is more skewed, rebalancing is likely to deliver better response times over a random switch to controller assignment. But, if the Pareto distribution evenly distributes traffic across switches (see seeds #2 and #5), random assignment does almost as well as rebalancing. In the Figure 4.9(b), we can observe that for all cases, rebalancing at least ensures that there is no controller that is severely overloaded while at least in four cases, random load balancing led to significant overload as evidenced by the high red bar.

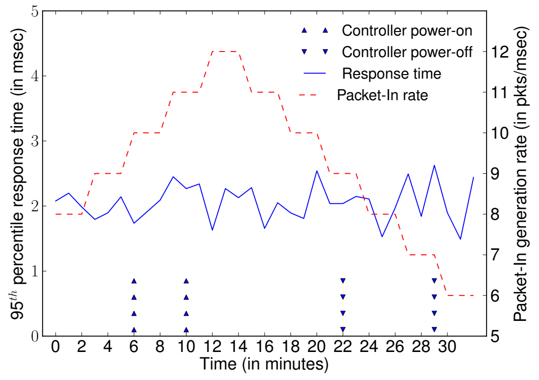


Figure 4.10.: Growing and shrinking ElastiCon

Effect of resizing. We demonstrate how the resizing algorithm adapts the controllers as the number of `Packet-In` messages increases and decreases. We begin with a network with 2 controllers and an aggregate `Packet-In` rate of 8,000 packets per second. We increase the `Packet-In` rate in steps of 1,000 packets per second every 3 minutes until it reaches 12,000 packets per second. We then reduce it in steps of 1,000 packets per second every 3 minutes until it comes down to 6,000 packets per second. At all times, the `Packet-In` messages are equally distributed across switches, just for simplicity. We observe 95th percentile of the response time at each minute for the duration of the experiment. We also note the times at which ElastiCon adds and removes controllers to adapt to changes in load. The results are shown in Figure 4.10.

We observe that ElastiCon adds a controller at the 6^{th} and 10^{th} minute of the experiment as the `Packet-In` rate rises. It removes controllers at the 22^{nd} and 29^{th} minute as the traffic falls. Also, we observe that the response time remains around 2ms for the entire duration of the experiment although the `Packet-In` rate rises and falls. Also, ElastiCon adds the controllers at 10,000 and 11,000 `Packet-In` messages per second and removes them at 9,000 and 7,000 `Packet-In` messages per second. As described earlier, this is because ElastiCon aggressively adds controllers and conservatively removes them.

4.5 Summary

We presented our design of ElastiCon, a distributed elastic SDN controller. We designed and implemented algorithms for switch migration, controller load balancing and elasticity which form the core of the controller. We enhanced Mininet and used it to demonstrate the efficacy of those algorithms.

5 FlowBricks: A FRAMEWORK FOR COMPOSING HETEROGENEOUS SDN CONTROLLERS

The popularity of SDNs has led to many open-source [42,44,46,47] and proprietary [8] implementations of SDN controllers. Each controller implementation supports a different set of services and is optimized for different performance metrics. For example, Beacon [46] is optimized for latency while Onix [8] provides higher throughput due to its distributed architecture. Network operators face the onerous task of selecting a controller implementation that can meet all current and future network services and performance requirements. In this chapter, we propose a framework, FlowBricks, to address this problem. It allows network operators to create an SDN control plane by combining services running on different controller platforms.

There are four very strong incentives to integrate *heterogeneous* control planes:

1. Modern networks are intelligent, and require implementation of sophisticated services such as advanced VPN, deep packet inspection, firewalling, intrusion detection – to name just a few. Moreover, this list continues to grow, increasing the need for methods to implement new network policies. However, not all services may be available on the same controller platform. It is also unlikely that one controller vendor will have the best-in-class implementation for all services. Hence, network operators can be forced to choose between not deploying a service or moving to another controller platform, which is expensive, disruptive, or even simply infeasible.

2. Even if services can be easily ported from one controller to another, each controller will have different performance characteristics. Some controllers may be suitable for high scalability while others may have low response times. In such cases, it may be desirable to run different services on different controllers to

match the controller performance characteristics with service requirements. For instance, services which reactively insert flow table entries to route flows need low response times and services which passively sample packets in the network might need to scale to keep up with network traffic load.

3. Some services may not require global network knowledge and may benefit from proximity to the data plane. Such services can be implemented on a controller platform in the network element itself while others can be implemented on the centralized controller. As we show later, FlowBricks can also be deployed on the every switch combine switch-local and centralized services.

4. In traditional networks, new network functionality can be implemented through middleboxes that support integration of services in a "bump-in-the-wire" manner (e.g., firewalls) [52]. Though middleboxes are transparent to existing services, in the long run, network operators prefer to integrate services supported by middleboxes into routers and switches in order to reuse existing hardware accelerators for packet processing and significantly reduce power and space demands. As a result, network management can be simplified and become less expensive, thus further motivating new abstractions that enable composition of services from heterogeneous controllers.

Thus, network operators are in need of constructors for flexible implementation of policies (that consist of services from various controller vendors, ideally, transparently to the services of integrated controllers) which allow for flexible and sound assemblage.

Standardizing the northbound and east-west interfaces to the controller is a technically feasible but impractical approach to combine services running on different controllers. In this paper, we show that a standardized southbound API is sufficient to combine services. We demonstrate this using OpenFlow as an example. While doing so, we found that the southbound API needs to convey certain information explicitly and specify certain switch behavior which OpenFlow does not. We describe these stipulations later which are essentially the properties of a southbound API that are needed for correctly implementing FlowBricks.

Previous research has tackled two types of service composition: *parallel* and *serial*. Parallel composition gives the illusion that each service operates on its own copy of the packet. Then, a set union of modifications from all services is applied to the packet. In the serial case, services operate on a packet in sequence. So, each service operates on a packet that has already been modified by prior services. Frenetic [53] does parallel composition of services while Pyretic [28] supports both serial and parallel composition. However, both assume that all services are running on the same controller. Flowvisor [54] slices the network and flows and assigns each slice to a different controller. It supports heterogeneous controllers but does not allow applying services from different controllers on the same traffic. [55] describes the design of an SDN hypervisor that is capable of combining services from heterogeneous SDN controllers similar to FlowBricks. However, their technique for composing forwarding table rules has two undesirable consequences. First, it leads to an exponential increase in number of forwarding rules in the datapath which makes their solution infeasible for switches with limited TCAM space. Second, it does not allow demultiplexing of northbound control plane messages. Hence, the SDN hypervisor will have to broadcast northbound OpenFlow messages (like `Packet-In` messages) to all controllers which can lead to incorrect behavior. The core contribution of this paper is a new framework named FlowBricks for integrating services from heterogeneous SDN controllers. The paper is organized as follows:

- We present our complete design of FlowBricks (Section 5.2). This includes the architecture of FlowBricks, policy definitions and algorithms for combining flow tables and other OpenFlow features from heterogeneous SDN controllers.
- We point out requirements in OpenFlow that impact the realization of FlowBricks (Section 5.3).
- FlowBricks introduces two performance overheads. First, it introduces additional flow table lookups for every packet on the datapath. Second, routing every mes-

sage through FlowBricks may impact throughput and latency of the control plane. We describe techniques to reduce the impact of FlowBricks on control plane and data plane performance(Section 5.4).

- We describe our experiments with FlowBricks involving 20 different combinations of five services implemented on four different controllers. We experimentally evaluate the technique to mitigate the impact of FlowBricks on control plane performance. (Section 5.5).

We begin with a brief review of OpenFlow terminology and switch forwarding behavior (Section 5.1).

5.1 Background: Packet Forwarding in OpenFlow

In this section, we recall terminology, switch components and forwarding behavior specified by OpenFlow 1.1 [56]. In short, a switch consists of flow tables and an action set.

Flow Entry. Each flow table contains one of more flow entries. Each flow entry contains a set of match fields for matching packets, a priority, and a set of instructions. When a packet hits a flow table, it is matched with all the flow entries in the table and exactly one is selected (if the packet matches multiple flow entries, the one with the highest priority is selected). Then, the instructions in the instruction set of that flow entry are executed. A controller may associate an idle timeout interval and a hard timeout interval with each flow entry. If no packet has matched the flow entry in the last idle timeout interval, or the hard timeout interval has elapsed since the flow entry was inserted, the switch removes the entry.

Instructions. An instruction results in changes to the packet, action set and/or pipeline processing. The `Apply-Actions` instruction contains a list of actions which are immediately applied to the packet being processed. The `Write-Actions` instruction contains a list of actions which are inserted into the action set and the `Clear-Actions` instruction removes all actions from the action set. The `Goto-Table` instruction indicates the next table in the pipeline processing. When a packet matches a flow entry, the instructions in the instruction set of that flow entry are executed.

Actions. An action describes packet handling. This includes forwarding a packet to a specific output port, pushing or popping tags, and modifying packet header fields.

Action Set. A set of actions are applied to a packet after all flow table processing has been completed. Being a *set*, the action set cannot contain more than one action of each type. A flow entry can populate the action set with the `Write-Actions` instruction and clear it with the `Clear-Actions` instruction.

Group Table. A group table consists of group entries which are identified by a unique 32-bit identifier. A group entry specifies more complex forwarding like flooding, multicast and link aggregation. Flow table entries can point a packet to a group entry using the `Group` action and its unique group entry identifier.

The OpenFlow pipeline processing for a packet starts at flow table 0. The packet is matched against the flow entries of flow table 0 to select a flow entry. Then, the instruction set associated with that flow entry is executed. If the instruction set contains a `Goto-Table` instruction, the packet is directed to another flow table and the same process is repeated. If the instruction set does not contain a `Goto-Table` instruction, pipeline processing stops and the actions in the action set are applied to the packet.

5.2 FlowBricks Design

To restate, FlowBricks aims to serially concatenate both proactive and reactive services from heterogeneous controllers onto the same traffic. In this section, we first describe the high-level system architecture. We then show how policies are defined in FlowBricks and describe a technique to combine flow table pipelines of controllers to realize these policies.

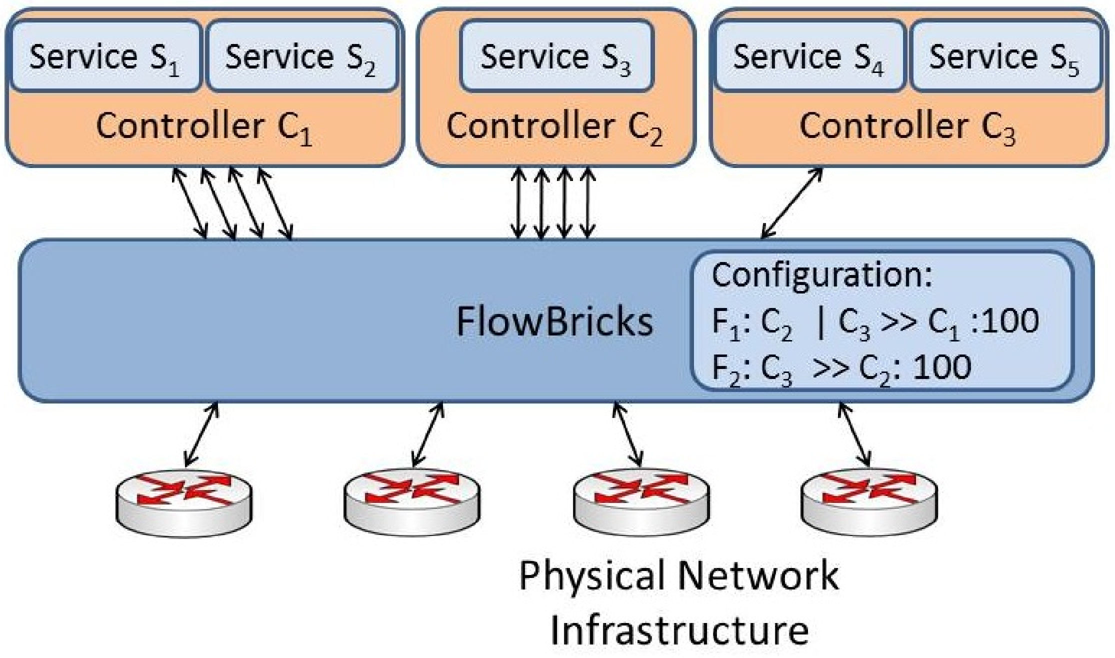


Figure 5.1.: FlowBricks system architecture.

5.2.1 System Architecture

The standardized communication protocol between the controller and switches presents a general way of integrating heterogeneous controllers. So, we implement FlowBricks as a flowbricks between the heterogeneous control and switches as shown in Figure 5.1. From the switches' perspective, FlowBricks is the control plane and from the controllers' perspective it is the forwarding plane. All switches are configured with the IP address and TCP port number of FlowBricks as the controller. Switches initiate a connection with FlowBricks and FlowBricks in turn initiates connections (one for each switch) with each controller. Each controller is configured with a set of services. We cannot assume that controllers can share state with each other. So, the set of services configured at each controller should be independent of those running on other controllers. Controllers send southbound control plane messages to FlowBricks. FlowBricks modifies these messages and forwards them to the switch that corresponds to the connection on which the message was received from the controller. Messages from the controllers to switches are modified such that the datapath configured on the switches combines the services from all controllers. The services are combined according to a policy configured in FlowBricks by the network operator. Northbound control plane messages from switches are forwarded by FlowBricks to one or more controllers using internal state and fields in the message (more details later).

5.2.2 Policy Definition

The policy configured on FlowBricks specifies how services from controllers are applied to traffic on the datapath. We use the | and >> operators for parallel and serial composition of heterogeneous controllers in a policy. These are similar to syntactic elements used in Pyretic for composing services (for a single controller).

The policy is specified on a per flow[1] basis. A policy is described by a flow, an ordered set of controllers whose services should be applied to that flow and a priority. The policy is configured in FlowBricks and the controllers themselves are unaware of the policy. For instance, three controllers $\mathtt{C}_1$, $\mathtt{C}_2$, $\mathtt{C}_3$ may be composed as follows:

$$\mathtt{F}_1 : \mathtt{C}_1 | \mathtt{C}_2 >> \mathtt{C}_3 : 100 \tag{5.1}$$

$$\mathtt{F}_2 : \mathtt{C}_2 >> \mathtt{C}_1 : 99 \tag{5.2}$$

This describes FlowBricks's policy for two flows, $\mathtt{F}_1$ and $\mathtt{F}_2$. FlowBricks applies services of $\mathtt{C}_1$, $\mathtt{C}_2$ and $\mathtt{C}_3$ to packets of flow $\mathtt{F}_1$ in that sequence. It applies services of $\mathtt{C}_2$ and $\mathtt{C}_1$ to packets of flow $\mathtt{F}_2$ in sequence. $\mathtt{F}_1$ has a priority of 100 while $\mathtt{F}_2$ has priority 99. A higher number indicates a higher priority. So, packets which match both flow definitions will be treated as $\mathtt{F}_1$'s packets. Controllers can be concatenated in two ways. A network operator may want to specify that a controller's flow tables should complete all their processing and apply actions (modifications) to the packet before the packet is processed by the next controller's flow tables. This is serial

[1]A flow can be defined on any fields of the packet header. For instance, a flow can be defined as packets with the same VLAN tag or packets destined for the same subnet.

composition of controllers and is represented as >>. Otherwise, the operator may wish to forward the packet to the flow tables of the next controller without applying the actions of the previous controller. This is parallel composition and is represented by |. During parallel composition, the actions generated by a controller's flow tables are added to an action set and the unmodified packet is matched with the following controller's flow tables. This accumalation of actions in the action set continues until reaching the end of the policy or a serial composition operator in the policy. At this point, the actions accumulated in the action set are applied to the packet.

The serial and parallel composition operators are an intuitive and powerful way to compose services, as shown in [28]. For example, consider a network administrator who wants to deploy traffic monitoring, network address translation (NAT), and routing services implemented on three different controllers ($\texttt{C}_1$, $\texttt{C}_2$, and $\texttt{C}_3$ respectively) for all HTTP traffic. The traffic monitoring service and NAT should see the unmodified packets while the routing service should be applied to packets after their addresses have been modified by NAT. One way to achieve this using the serial and parallel composition operators is shown in the equation below:

$$\mathit{http} : \texttt{C}_1|\texttt{C}_2{>>}\texttt{C}_3 : 100 \tag{5.3}$$

In this policy, $\texttt{C}_1$ (traffic monitoring) and $\texttt{C}_2$ (NAT) are composed with the parallel composition operator (|). The umodified packet will be matched with their flow tables and the actions will be stored in an action set. Then, these actions will be applied to the packet for serial composition (>>) before the packet is matched with $\texttt{C}_3$'s flow tables.

5.2.3 Constraints on Combining Flow Table Pipelines

Controllers $\texttt{C}_1$, $\texttt{C}_2$ and $\texttt{C}_3$ use a sequence of messages to install their flow table pipelines on the switch. FlowBricks modifies these messages such that a combined flow table pipeline is installed on the switch. The combined flow table pipeline should apply the services of controllers to packets according to the policy configured by the network operator. Creating a combined flow table pipeline involves linking flow tables from different controllers or combining flow entries from controllers' flow tables into a single flow table. This can be done in many ways. In this section, outline constraints which have to be met by any technique for correctly combining flow table pipelines.

a. *Combining flow table entries:* One way to combine two flow tables into a single flow table involves computing the cross product of the flow tables [57]. Pyretic uses this technique to combine flow tables from different services. However, flow table entries may have configurations like counters or idle timeout. It is impossible to assign a correct idle timeout to a flow table entry obtained by combining two flow table entries with different idle timeouts. In the most general case, each controller may assign different idle timeout values to its flow table entries. This would prevent FlowBricks from computing the cross product of flow tables.

b. *Duplicating flow table entries:* Some techniques to combine flow tables may require that each flow in the policy have its own set of flow tables. For instance, consider the policy in 5.1 and 5.2. Packets of F_1 may be processed by flow tables 1 to 4 in the combined pipeline while packets of F_2 may be processed by flow tables 5 to 7. This would require duplicating flow table entries of C_1 and C_2 since they should be matched against packets of both flows. However, this would lead to incorrect behavior for flow table entries that have idle timeouts. If a controller's flow table entry with an idle timeout is duplicated by FlowBricks and inserted into two flow tables on the switch by FlowBricks, one copy of the flow table entry may be removed due to the idle timeout while the other remains. This may lead to unexpected switch behavior from the controller's perspective since packets of one flow continue to be matched against a flow table entry while the packets of another flow don't.

c. *Modifying flow table pipeline:* FlowBricks may need to modify the combined flow table pipeline on the switch when the administrator changes the policy. However, this change should not require moving a controller's flow table entry from one flow table to another. Doing so would reset the idle timer in the switch leading to wrong behavior. Similarly, if a controller adds or removes a flow table entry, FlowBricks should not add or remove the flow table entries of other controllers.

d. *Forwarding `Packet-In` messages:* A switch can send a packet from the datapath to the controller using a `Packet-In` message. When FlowBricks receives a `Packet-In` message from the switch, it needs to forward the message to the controller whose flow table entry generated that `Packet-In` message. FlowBricks needs to identify this controller using just the fields in the `Packet-In` message. The `Packet-In` message has a table-id field which holds the table-id of the flow table entry which generated the `Packet-In` message. To use the table-id field for forwarding `Packet-In` messages, FlowBricks should not insert flow table entries of different controllers into the same flow table on the switch. To rephrase, each flow table on the switch should contain flow table entries from just one controller to enable demultiplexing of `Packet-In` messages to controllers using table-id.

We now propose a technique to combine flow table pipelines which obeys the above constraints.

5.2.4 Combining Flow Table Pipelines

In this section, we describe a technique to combine flow table pipelines which obeys the above constraints. Figure 5.2 shows the flow table pipeline obtained by combining the flow table pipelines of $\mathtt{C}_1$, $\mathtt{C}_2$ and $\mathtt{C}_3$ using this technique. The combined flow table pipeline contains flow tables from all controllers. In addition, the pipeline contains one flow table (which we call *transitional* flow table) for each controller. The

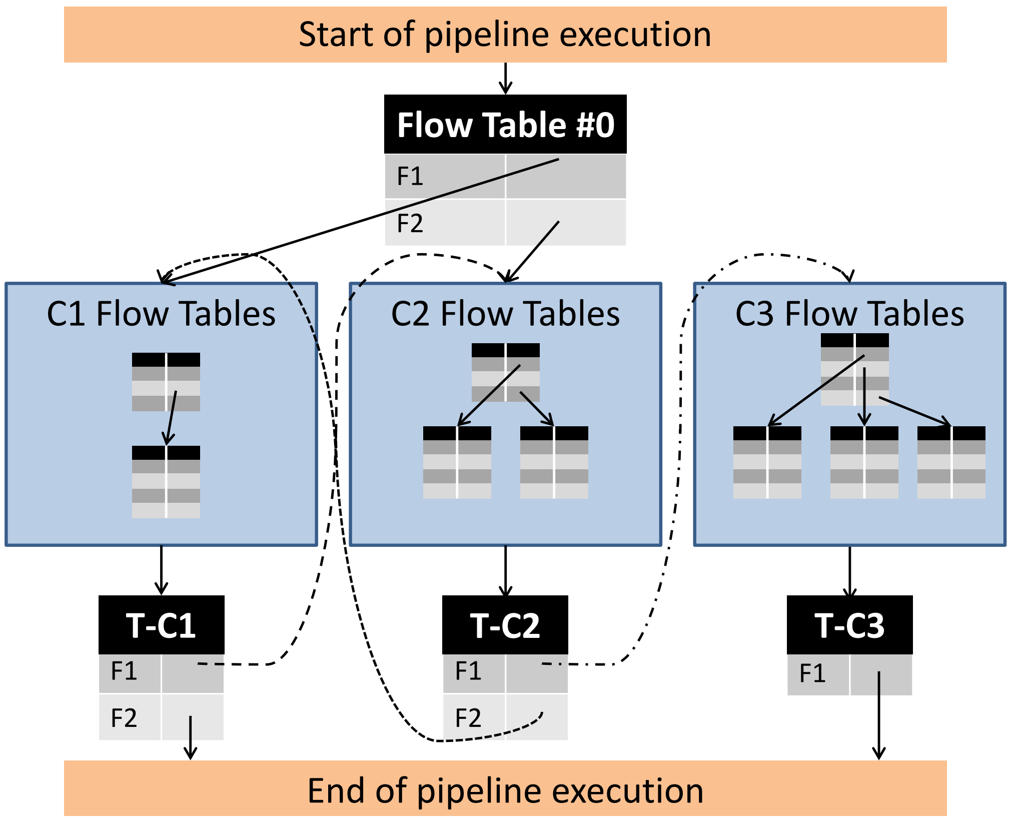


Figure 5.2.: Pipeline configured on switch by FlowBricks

transitional flow table directs pipeline execution in the desired sequence between the controllers' flow tables. Transitional flow tables are labeled $\mathtt{T\text{-}C}_1$, $\mathtt{T\text{-}C}_2$ and $\mathtt{T\text{-}C}_3$ in the figure. FlowBricks modifies each controller's flow tables to ensure that pipeline processing proceeds from a controller's flow tables to its transitional flow table. Each transitional flow table contains one flow table entry for every flow in the policy. The flow table entry directs pipeline execution to the flow tables of the next controller in the policy for that flow.

An Example

Figure 5.2 shows the combined flow table pipeline for the policy in equations 5.1 and 5.2. Assume a packet belonging to $\mathtt{F}_1$ arrives at the switch. Flow table pipeline processing for the packet begins at flow table 0. The packet will match the flow table entry for F_1 in flow table 0. This flow table entry directs execution to C_1's flow tables since it is the first controller in F_1's policy. After the packet has been processed by C_1's flow table, it proceeds to T-C_1. The entry for F_1 in T-C_1 directs execution to C_2's flow table. From there execution for F_1's packets proceeds to T-C_2, then to C_3's flow tables and finally to T-C_3. The flow table entry for F_1 in T-C_3 terminates pipeline execution since C_3 is the last controller in F_1's policy. For packet belonging to F_2, pipeline execution also begins at flow table 0. But, F_2's packets will match F_2's flow table entry in flow table 0. This entry directs execution to C_2's flow tables. From there, execution for F_2's packets proceeds to T-C_2, then C_1's flow tables and finally to T-C_1.

Flow table number map

Each controller inserts its flow table entries into flow tables which are numbered starting with 0. If FlowBricks forwards messages from the controller to the switch without modification, a flow table on the switch may contain flow table entries from multiple controllers. But, each controller should have its own set of flow tables in the combined flow table pipeline. To achieve this, FlowBricks needs to map each controller's flow tables onto unique flow tables on the switch. For this, FlowBricks maintains a flow table number map. This data structure maps the controller-id, flow table number pair to a unique flow table number on the switch. When FlowBricks receives a message containing a "flow table number" field from the controller, it does a lookup on the flow table number map and replaces the controller's flow table number with the unique flow table number before forwarding the message to the switch. A message from the switch to FlowBricks may also contain a flow table number field. In such a message, FlowBricks needs to replace the flow table number with the controller's flow table number. For this it does a reverse lookup on the flow table number map using the flow table number in the message. This lookup returns a controller-id and the controller's flow table number. FlowBricks inserts the controller's flow table number in the message and forwards the message to that controller.

Transitional flow tables

We now describe how FlowBricks determines instructions, match fields and priority of the flow entries in transitional flow tables.

Instructions. Flow entries in a controller's flow tables contain instructions to add actions to the action set while leaving the packet unmodified. The flow table entries in a transitional flow table contain different instructions depending upon whether they direct execution across a serial or parallel composition operator. Transitional flow table entries that direct execution across a parallel composition operator contain a `Goto-Table` instruction to direct execution to the following controller's flow tables. For example, the entry for F_1 in $T\text{-}C_1$ will contain a `Goto-Table` instruction to C_2's first flow table. This way, the packet remains unmodified and the actions generated by the following controller's (C_2 in this case) flow tables are also added to the action set.

Transitional flow table entries that direct execution across a serial composition operator contain an instruction to apply all the actions accumalated in the action set along with a `Goto-Table` instruction. For example, the flow entry for F_1 in $T\text{-}C_2$ will direct execution across a serial composition operator. So, it will contain an instruction to apply the actions in the actions set. At that point, the action set will contain actions inserted by C_1 and C_2's flow tables entries. F_1's flow table entry in $T\text{-}C_2$ will also contain a `Goto-Table` instruction to direct pipeline execution to C_3's first flow table.

Matching Packets. Flows in FlowBricks's policy are defined using fields of the packet header. However, the fields in the packet header may be modified by flow table entries. So, to identify the flow for a packet, we cannot match the fields in the packet header. We illustrate the problem with an example. Let $\texttt{F}_3$ be defined as packets with destination IP address X. The policy for $\texttt{F}_3$ is shown below.

$$\texttt{F}_3 : \texttt{C}_1{>>}\texttt{C}_2{>>}\texttt{C}_3 : 98 \tag{5.4}$$

Assume that $\texttt{C}_1$'s flow table entries rewrite the destination IP address. The IP address modifications will be applied to the packet at the serial composition operator between $\texttt{C}_1$ and $\texttt{C}_2$. So, the flow table entry in the transitional flow tables following $\texttt{C}_2$ cannot identify $\texttt{F}_3$'s packets by matching the destination IP address with X.

To work around this problem, we use the metadata field of a packet. OpenFlow associates a metadata with every packet being processed in the flow table pipeline. It also provides instructions to modify the bits in the metadata. The match for a packet in a flow table entry can also include bits of the metadata. FlowBricks generates a unique identifier for every flow in the policy. The flow table entries in the first flow table (flow table #0) of the pipeline match packets with the fields in the packet header. This is not a problem since the packet has not yet been modified. These flow table entries also write the unique flow identifier into the packet's metadata field. Flow table entries in subsequent transitional flow tables match a packet with the metadata rather than the packet header fields.

Priority. The policy definition in FlowBricks assigns a priority to each flow. FlowBricks assigns the same priority to the corressponding flow table entry in the transitional flow table. For instance, $\texttt{F}_1$'s entry in $\texttt{T-C}_1$, $\texttt{T-C}_2$ and $\texttt{T-C}_3$ will have priority 100, since that was $\texttt{F}_1$'s priority in equation 5.1.

Group Tables

Controllers may insert group entries into the group table. Each group entry is identified by a unique group identifier. Like the flow table number map, FlowBricks maintains a mapping from each controller's group identifiers to globally unique group identifiers. For southbound messages, FlowBricks replaces the controller's group identifier with a globally unique group identifier by doing a lookup on this map. This replacement is done for all messages that contain a group identifier. This includes messages that modify the group table and messages that contain the `Group` action which requests packet processing through a specified group. For northbound messages, it performs the reverse operation using reverse-lookup on the same map.

Handling Policy Updates

The networks administrator may update the policy configured in FlowBricks at any time. This may involve adding/removing a flow to the policy, changing the processing for an exisiting flow or add/removing a controller. All these update involve changes to only the transitional flow tables. However, these changes need to be done atomically to avoid incorrect packet processing during update. For switch-local changes, updates can be done atomically using the `Bundle` messages introduced in OpenFlow 1.4 [58]. Mechanisms proposed in [59] can be used for updates that span multiple switches.

5.2.5 OpenFlow Message Processing

In this section, we describe message processing for some common OpenFlow messages. The processing of other messages usually involves just translation of flow table numbers.

Flow-Mod Messages. OpenFlow uses a `Flow-Mod` message to insert a flow table entry in a switch. The `Flow-Mod` message contains the number of the flow table where the flow table entry should be inserted. FlowBricks modifies the flow table number using the flow table number map described in Section 5.2.4. A `Flow-Mod` message also contains a list of instructions which are executed when a packet matches the flow table entry. The absence of a `Goto-Table` instruction from the instruction list of a `Flow-Mod` message indicates that the controller expects pipeline processing to terminate at that flow table entry. For such flow table entries, FlowBricks inserts a `Goto-Table` instruction to direct pipeline execution to a transitional flow table.

Packet-In Messages. When a packet does not match any flow table entry, the switch generates a `Packet-In` message to the controller. A controller can configure the number of bytes in the `Packet-In` message and buffering for the packet in one of two ways. The controller can request the switch to buffer the packet and send a fixed number of bytes of the packet in the `Packet-In` message. Otherwise, it can request the switch to send the entire packet in the `Packet-In`. FlowBricks configures the switch to buffer the packet and send a fixed number of bytes of the packet in the `Packet-In` message. To handle controllers that have requested the entire packet in the `Packet-In` message, FlowBricks sets the fixed number of bytes in the `Packet-In` message equal to the maximum transmission unit (MTU) of the network.

The `Packet-In` message also contains a flow table number. This is the flow table number of the flow entry that generated the `Packet-In` message. Using the flow table number, FlowBricks performs a reverse lookup on the controller to switch flow table map. It replaces the flow table number in the `Packet-In` message with the controller's flow table number and forwards the `Packet-In` message to the corresponding controller (i.e. the controller which inserted flow table entries into that flow table).

Packet-Out Messages. A `Packet-Out` message may contain the entire packet or a buffer identifier where the packet is stored. In either case, the FlowBricks just forwards the packet to the switch without modification.

Barrier and Statistics Messages. FlowBricks forwards barrier and statistics request message from the controller to the switch and their replies in the reverse direction. The transaction ID field of the request messages need to be unique. Since the transaction ID field of the request messages is populated by the controller, two controllers may send a request message with the same transaction ID to the same switch. To avoid this problem, FlowBricks replaces the transaction ID in the request messages with a unique value before forwarding it to the switch and does the reverse operation for the replies. This ensures uniqueness of the transaction ID field in a manner that is transparent to the controllers.

5.3 OpenFlow Limitations

We also point out stipulations in OpenFlow 1.3 which impede implementation of FlowBricks using the flow table concatenation algorithm described earlier. For ease of exposition, we assume that FlowBricks is configured with flows shown in eqs. (5.1) and (5.2).

Packet Drops. OpenFlow does not have an explicit action to drop packets. OpenFlow specifies an `Output` action that takes an output port as parameter. At the end of pipeline execution, if the action set contains the `Output` action, the packet is transmitted on that output port. Otherwise, the packet is dropped. Any flow entry can force a packet drop by clearing the action set and terminating pipeline execution. Consider $\texttt{F}_1$ in Equation 5.1. $\texttt{C}_1$ can never drop a packet since FlowBricks will not let pipeline execution terminate at any of $\texttt{C}_1$'s flow tables. It will always add a `Goto-Table` instruction to $\texttt{C}_1$ flow entries to continue execution to $\texttt{T-C}_1$.

A simple way to get around this problem is to introduce a *Drop* action in the OpenFlow specification. Flow table entries that want to drop a packet, should insert the `Drop` action into the action set. If the action set contains a drop action when it is applied to the packet, the packet is dropped. Otherwise, it is sent on the output port specified by the `Output` action.

When FlowBricks sees an instruction to insert a `Drop` action in a flow table entry, it can deduce that the flow table entry wants to drop the packet. So, FlowBricks should not insert a `Goto-Table` instruction in that flow table entry. This will terminate pipeline execution without processing flow tables from subsequent controllers.

Executing Action Set. The action set is executed after pipeline processing has finished. However, if policy definition contains a serial composition operator between two controllers (between $\texttt{C}_2$ and $\texttt{C}_3$ for $\texttt{F}_1$), then the switch should apply the actions in the action set after the packet has finished matching the flow tables of controllers before the serial composition operator. Currently, there is no way of doing this in OpenFlow.

This problem is addressed by introducing an `Execute-Actions` instruction. This instruction executes all the actions stored in the action set and clear the action set. FlowBricks inserts this instruction in those entries of the transitional flow tables which direct execution across a serial composition operator in the policy definition.

Sequence of Flow Tables. Switch flow table numbers start with 0. OpenFlow states that packet processing cannot go from a higher flow table number to a lower one. This limits the possible rules that can be composed in FlowBricks. For instance, it is impossible to simultaneously implement F_1 and F_2 (eqs. (5.1) and (5.2)) in FlowBricks. If C_1's flow tables have lower flow table numbers than those of C_2, then, F_2's policy cannot be configured on the switch since it will involve going from a higher to lower flow table number. Similarly, if C_1's flow tables have higher flow table numbers than those of C_2, F_1's policy cannot be configured on the switch.

As a simple workaround, FlowBricks could duplicate flow tables of C_1's flow tables. One set of C_1's flow tables could have lower flow table numbers than those of C_2's. Another set could have higher flow table numbers than C_2. However, this workaround is infeasible if C_1's flow table entries have soft timeout interval. OpenFlow specifies that a flow table entry will time out and be removed from the flow table when no packet matches the flow table entry for a pre-configured soft timeout interval. If we duplicate C_1's flow table entries, one set of entries may soft timeout and be removed while the other set continues to match packets. This would leave the switch in an inconsistent state.

Parallel Compostion Limitations. Consider a policy that composes two controllers in parallel: $C_1 \mid C_2$. C_1 and C_2's flow tables should add actions to the action set and these actions should be applied to the packet at the end of the pipeline. There are two potential problems with this. First, OpenFlow allows flow table entries to directly apply actions to packets (without adding them to the action set) using the `Apply-Actions` instruction. If C_1's flow table entries use the `Apply-Actions` instruction, C_2's flow tables will be matched with the modified packet which violates parallel composition. Second, C_2's flow tables may use the `Clear-Actions` instruction to remove actions inserted in the action set by $\texttt{C}_1$. Both these errors can be reported by FlowBricks using simple checks. In some cases, the network operator could reorder the controllers in the policy to achieve the desired behavior. For instance, the policy $\texttt{C}_2 \mid \texttt{C}_1$ will work correctly if $\texttt{C}_1$ uses `Apply-Actions` and $\texttt{C}_2$ uses `Clear-Actions`.

5.4 Performance Considerations

The ability to compose a SDN control plane from heterogeneous controllers comes at a cost. FlowBricks inserts transitional flow tables into the flow table pipeline which adds flow table lookups to packet processing. The latency between the controller and switches increases since message need to routed via FlowBricks. We propose ways to mitigate these overheads introduced by FlowBricks.

5.4.1 Reducing Number of Flow Table Lookups

The number of flow table lookup operations required for packet processing determines the line rate that the switch can support. FlowBricks adds one flow table lookup in the transitional flow table per controller. This reduces the line rate supported by the switch. [60, 61] propose techniques efficiently to reduce the number of lookups at the expense of TCAM space. However, these techniques cannot be directly applied in FlowBricks due to constraints introduced by idle timeouts and counters associated with flow table entries. Here, we propose computing the cross product of two sets of flow table entries to avoid an lookups in transitional flow tables. This is similar to the technique used by Pyretic to combine policies but is applied selectively to flow table entries that do not have idle timeouts or counters associated with them. We first describe the technique with an example and then outline the algorithm to implement the technique.

Example

Figure 5.3 shows unmodified $\texttt{C}_1$ flow tables on the left and $\texttt{C}_1$ flow tables after taking cross product with $\texttt{T-C}_1$ on the right. Initially, $\texttt{C}_1$ has two flow tables with two flow table entries each. Assume that flow table entry for $\texttt{F}_D$ has an idle timeout. To avoid the flow table lookup for $\texttt{T-C}_1$, we iterate over $(p_i, q_j) \in P \times Q$ where P is the set of flow table entries in $\texttt{T-C}_1$ and Q is the set of flow table entries in $\texttt{C}_1$'s flow tables that contain `Goto-Table` instructions to $\texttt{T-C}_1$ and do not have idle timeouts or counters associated with them. In Figure 5.3, $P = \{\texttt{F}_1, \texttt{F}_2\}$ and $Q = \{\texttt{F}_A, \texttt{F}_C\}$. For each (p_i, q_j), generate a flow table entry by computing the intersection of the flow patterns, sum of the priorities and union of the instruction sets. Insert this new flow table entry in q_j's flow table. The resulting flow tables will avoid the additional lookup in $\texttt{T-C}_1$ for packets of flows $\texttt{F}_A$ and $\texttt{F}_C$ as shown on the right side in Figure 5.3.

Algorithm

A controller sends a `Flow-Mod` message to insert or delete a flow table entry. To configure the modified flow table pipeline on the switch, FlowBricks needs to processes `Flow-Mod` messages differently. We first describe how FlowBricks processes `Flow-Mod` messages that insert a flow table entry.

a. When FlowBricks receives a `Flow-Mod` message from a controller to add a flow table entry, it processes the message as described in Section 5.2.5. If FlowBricks did not add a `Goto-Table` instruction to direct pipeline execution to a transitional flow table, then processing of the `Flow-Mod` messages stops and the message is sent to the switch. If the `Goto-Table` instruction was added the following steps are executed.

b. FlowBricks checks if the flow table entry being inserted by the `Flow-Mod` message has an idle timeout or counters associated with it. If it does, then the additional

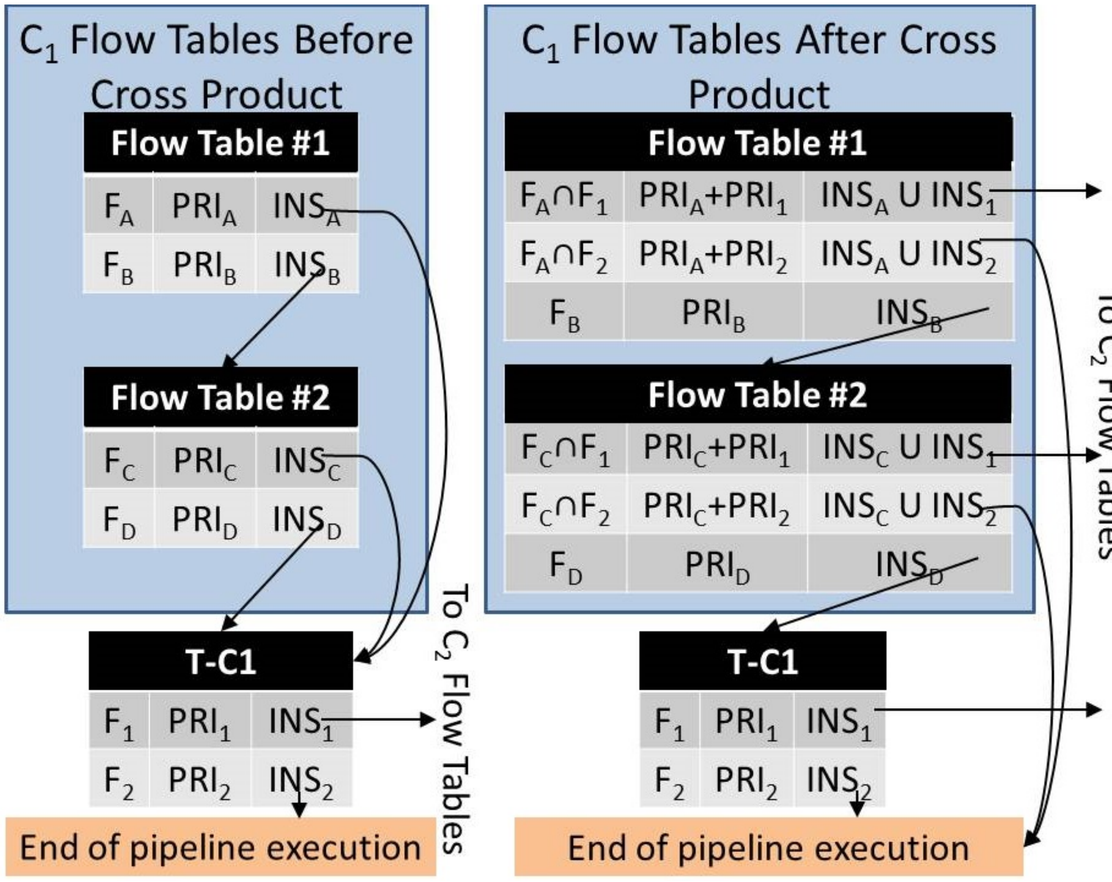

C1 Flow Tables Before Cross Product
Flow Table #1
FA PRIA INSA
FB PRIB INSB
Flow Table #2
FC PRIC INSC
FD PRID INSD
T-C1
F1 PRI1 INS1
F2 PRI2 INS2
End of pipeline execution
To C2 Flow Tables
C1 Flow Tables After Cross Product
Flow Table #1
FA∩F1 PRIA+PRI1 INSA U INS1
FA∩F2 PRIA+PRI2 INSA U INS2
FB PRIB INSB
Flow Table #2
FC∩F1 PRIC+PRI1 INSC U INS1
FC∩F2 PRIC+PRI2 INSC U INS2
FD PRID INSD
T-C1
F1 PRI1 INS1
F2 PRI2 INS2
End of pipeline execution
To C2 Flow Tables

Figure 5.3.: $\texttt{C}_1$ flow tables before and after computing cross product with $\texttt{T-C}_1$.

lookup in the transitional flow table cannot be avoided for packets that match this flow table entry. FlowBricks skips the following steps and forwards the `Flow-Mod` message to the switch.

c. FlowBricks removes the `Goto-Table` instruction from the flow table entry. Then, FlowBricks iterates over every entry in the transitional flow table of the controller that sent the `Flow-Mod` message. For each entry in the transitional flow table, FlowBricks generates a new flow table entry by intersecting the flow pattern, summing the priority and computing the union of the instruction set with the `Flow-Mod` message. If the transitional flow table has n flow table entries, the above iteration will generate n `Flow-Mod` messages. If the original flow table entry had a hard timeout, configure the same hard timeout to all n flow table entries generated by taking the intersection with the transitional flow table. Finally, all n `Flow-Mod` messages are encapsulated in a `Bundle` message to ensure that they are applied atomically on the switch.

Note that we do not violate any of the constraints in Section 5.2.3. We only combine a controller's flow table entries with its transitional flow table entries for those entries that do not have idle timeouts and counters. Also, each flow table on the switch will still contain just one controller's flow table entries since P and Q never contain flow table entries from two different controllers.

Avoiding the additional lookup comes at the cost of increase in number of flow table entries. For example, consider a transitional flow table with n flow table entries. Suppose that m flow table entries direct pipeline execution to that transitional flow table using the `Goto-Table` instruction. The cross product of the two sets of flow table entries will yield $n * m$ flow table entries instead of $n + m$ flow table entries.

In addition to the above change, two more changes are needed. First, a `Flow-Mod` message to delete a flow table entry needs to the delete the n entries that were generated by the corresponding `Flow-Mod` to insert a flow table entry. Second, a modification to the policy configured in FlowBricks will require modifying all the flow table entries which were generated by computing the cross product with the flow table entries in the transitional flow tables. Both these changes can be done atomically using the `Bundle` message provided in OpenFlow.

5.4.2 Deployment Alternatives

A key feature of FlowBricks is that it does not require global knowledge of the network. Hence, even if the controller is centralized, FlowBricks can be implmented in a distributed manner. In the extreme case, we could even have a separate instance of FlowBricks for each switch. This gives us a lot of flexibility in deploying FlowBricks to minimize its impact on performance metrics like control plane throughput and latency.

A central module is still needed to push policy changes to all distributed instances of FlowBricks. In this section, we describe three alternatives to deploy FlowBricks.

a. *Colocate with controllers* FlowBricks and all the controllers can be run on the save server. Such a setup will help reduce latency because the traffic between FlowBricks and the controller will remain within the same host. It can offer further latency benefits if the controller can be configured to use inter-process communication methods offered by the OS instead of TCP/IP. However, if the controller processes are CPU-intensive, then the CPU may become a bottleneck.

b. *On a dedicated server* FlowBricks can be deployed on a dedicated server or a set of dedicated server. This adds constrol plane latency since each packet from the switch will be routed to the controller via FlowBricks's server. However, for CPU-intensive controllers or deployments where we expect a lot of control plane traffic, it is desirable to run FlowBricks on dedicated server.

c. *Colocate with the switch* A third option is colocating FlowBricks with the switch software. This can be easily achieved for software switches (like Open vSwitch) which run on end hosts. For physical switches, the switch vendor can integrate FlowBricks's in the software running on the switch. If such a deployment can be achieved, it offers the both throughput and latency benefits. In this setup, each switch runs its own instance of FlowBricks and controllers can run on different servers. This prevents CPU bottlenecks. Traffic between the switch and FlowBricks remains within the same host, thus providing the latency benefits.

In Section 5.5, we explore the trade-off between latency and throughput for the three deployment scenarios described above.

5.4.3 Using FlowBricks with ElastiCon

FlowBricks and ElastiCon, are both SDN controller architectures to address scalability of the SDN control plane along with number of network services and control plane resource demands. They can both be combined in two ways. First, a service which requires control plane resources to scale in response to traffic demands could be run on ElastiCon as shown in Figure 5.4. Other services, running on different controllers may be purely reactive and not require the benefits provided by ElastiCon.

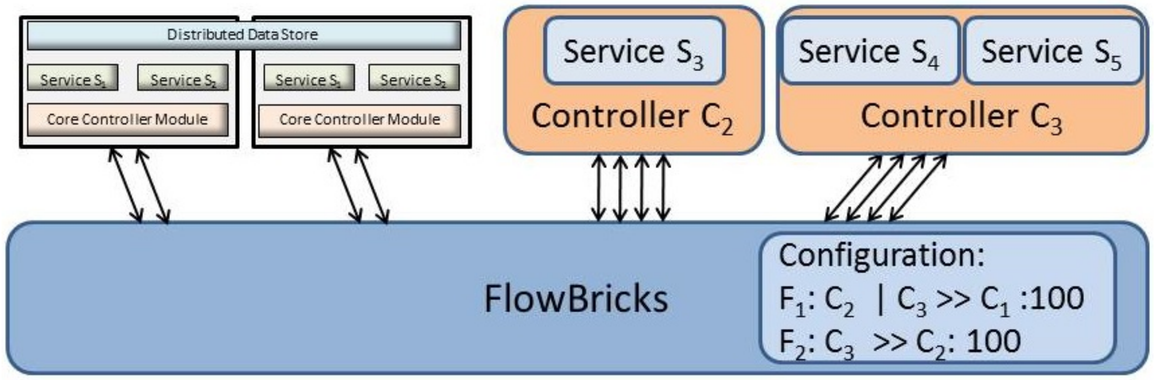


Figure 5.4.: Using ElastiCon with FlowBricks.

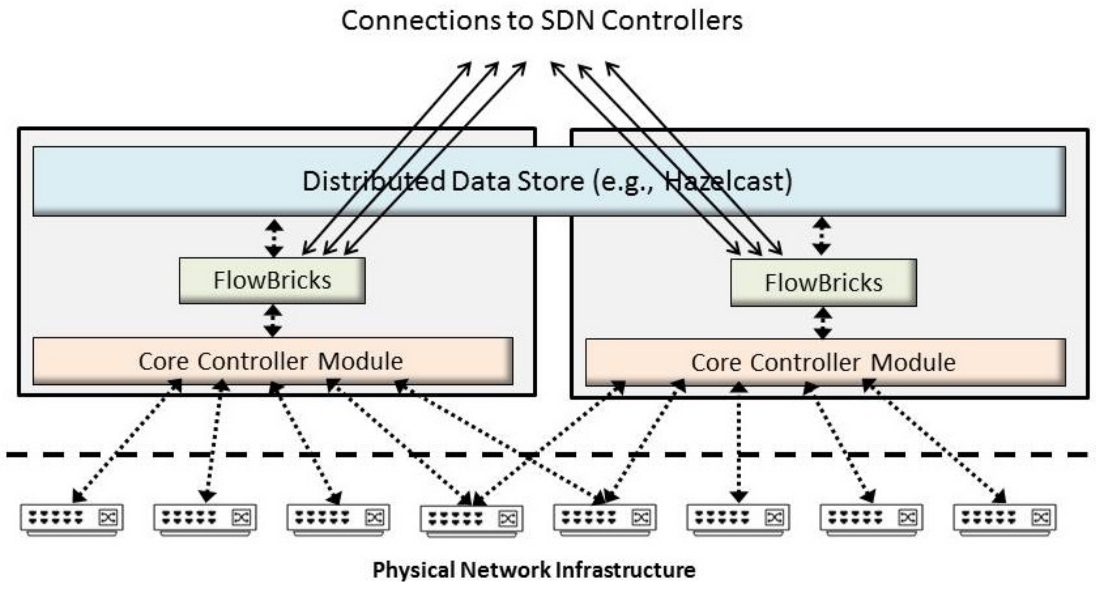

Connections to SDN Controllers
Distributed Data Store (e.g., Hazelcast)
FlowBricks
FlowBricks
Core Controller Module
Core Controller Module
Physical Network Infrastructure

Figure 5.5.: FlowBricks as an application in ElastiCon.

Also, ElastiCon may be used to improve the scalability of FlowBricks itself Figure 5.5. The application in ElastiCon would be implement the service composition algorithm described in Section 5.2. It would also be responsible for maintaining TCP connections with the controllers while the core module would maintain the switch TCP connections. ElastiCon would allow dynamic addition and removal of flowbricks servers in response to control plane resource demands. When a server is added to FlowBricks, the core controller module would dynamically migrate switches to the new server. The application would be responsible for migrating TCP connections to the heterogeneous controllers and reading and writing state (like flow table number map) to the distributed key-value store.

5.5 Evaluation

In this section, we describe our implementation and use it to evaluate the overhead imposed by FlowBricks.

5.5.1 Implementation

We considered implementing FlowBricks as a plugin for Pyretic. However, Pyretic's core module computes a cross product to combine flow tables. During cross product of flow tables, flow table entry features (like timeouts) are lost. This prevented us from using Pyretic to implement FlowBricks.

FlowBricks can be implemented as a module in any controller (we chose Floodlight [46]). We reused Floodlight's core module for communication with switches. The switches in the network are configured with the IP address and TCP port of Floodlight. They establish a connection with Floodlight's core module. The core module sends a notification to the FlowBricks module when a new switch connects to Floodlight.

The FlowBricks module is configured with a policy and the IP addresses and TCP ports of the heterogeneous controllers. The module implements the switch-side communication of the OpenFlow protocol. When it receives a notification for a new switch from the core module, it establishes a new TCP connection with each controller. The controllers send switch configuration messages over this connection. FlowBricks module modifies these messages as described in previous sections and forwards them to the switch.

For our experiments, we used Open vSwitch (OVS), an open source OpenFlow-compliant software switch. We modified OVS to address the limitation of the OpenFlow protocol described in Section 5.3. In particular, we allowed pipeline processing to transition from a higher to lower flow table number and added support for `Execute-Actions` instruction and `Drop` packet action.

5.5.2 Examples

To demostrate the utility and flexibility of FlowBricks, we experimented with various combinations of services and controller implementation. Table 5.1 shows the services that we used and Table 5.2 shows the policies for combining them. We briefly describe them below.

Table 5.1.: Services in FlowBricks

Service	**Abbrev.**	**Controllers**	**Actions**
Learning Switch	LS	Floodlight, Pox, Beacon	Packet forwarding
Shortest Path Routing	SPR	Floodlight, Pox, Beacon	Topology discovery, packet forwarding
Quality of Service	QOS	Floodlight	Set IP DSCP bits
Access Control	AC	Floodlight	Packet filtering
Address Rewriting	NAT	Pyretic	Set IP addresses
Deep Packet Inspection	DPI	Pyretic	Send packet to controller
ARP Responder	ARPR	Pox	Send ARP response

Table 5.2.: Policies in FlowBricks

Policy
* : AC >> LS†: 100
LLDP : SPR†: 100 * : AC >> SPR†: 99
* : DPI >> SPR†: 100
* : QOS | LS†: 100
ARP : DPI >> ARPR : 100 * : DPI >> LS†: 99
ARP : AC >> ARPR : 100 LLDP : SPR†: 99 * : AC >> SPR†: 98
LLDP : SPR†: 100 * : AC >> QOS | NAT >> SPR†: 99

† Implementations of this service were available on multiple controllers. We verified each policy using all combinations of service implementations.

Learning Switch. The learning switch application learns MAC addresses of hosts and installs rules reactively when new flows arrive. If the service has not learnt the location of the destination, it floods the packet.

Shortest Path Routing. This service uses LLDP messages to discover the topology of the network. For policies invovling the shortest path routing service, we configured FlowBricks to forward apply the routing services to LLDP packets. It also learns MAC addresses from packets sent by end hosts. When a new flow arrives, it uses the topology and destination MAC address to compute the shortest path to the destination. It then installs flow table entries on all switches along the route.

Access Control. We implemented this service to permit communication only between certain pairs of hosts in the network. The service proactively installs rules which match packets between host pairs that are allowed to communicate with each other. Other packets match a low priority rule that drops packets using the `Drop` instruction.

Quality of Service. This service inspects the packet header fields and sets the type of service field of the IP header according to a configured policy. This field can be used by downstream services and switches to assign packets to queues.

Address Rewriting. This service rewrites IP addresses to emulate a NAT middlebox. We modified this service to decouple it from the routing service. As a consequence, we also had to disable some checks in Pyretic's core module which ignore rules that do not forward a packet to an output port.

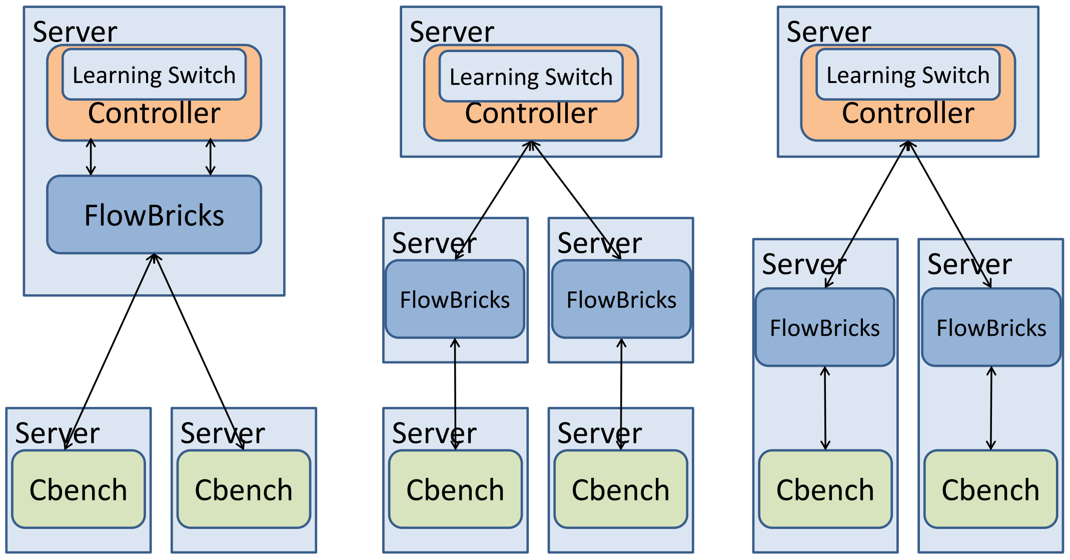
Deep Packet Inspection. Pyretic provides a sample implementation of this service. It sends every packet to the controller and prints it to the console. Like the address rewriting service, we modified one line of code to decouple this service from the routing service.

ARP Responder. This service responds to ARP request messages instead of flooding them in the network. It install rules to redirect all ARP messages to the controller. The service learns MAC addresses of hosts and responds to ARP requests for MAC address of known hosts.

We experimented with various combinations of services as shown in Table 5.2. For each policy, we ran each service on a separate controller. We configured the policy in FlowBricks and used Mininet [50] to emulate a tree topology. We verified that the datapath was correctly configured using Open vSwitch [51] utilities. We inspected packets using tcpdump [62] to ensure that they were being modified and forwarded correctly. For services whose implementations were available on multiple controllers, we verified the policy with all combinations of implementations.

5.5.3 FlowBricks Overhead

Figure 5.7 shows CDF time taken by the control plane to respond to `Packet-In` messages as observed at the switch. The blue line shows the response time with FlowBricks and two separate controllers running one application each. The green line shows the response time of a controller running both applications. FlowBricks causes



(a) Co-locating with controller. (b) Dedicated servers. (c) Co-locating with switches.

Figure 5.6.: Setup used for comparing the deployment alternatives.

a two-fold increase in response time. Inserting FlowBricks doubles the communication overhead (transmission and message parsing time) for every message. So, the increase in control plane response time is explained almost entirely by the fact that all control plane traffic needs to be redirected through FlowBricks. However, the 95th percentile response time of 2.2ms with FlowBricks is still well below the acceptable flow setup time of 5-10ms for LAN environments [63].

5.5.4 Performance Comparison

Since the increase in response time is almost entirely due to the additional hop introduced by FlowBricks in the control plane, we expect the three deployment scenarios (Section 5.4) to have different performance characteristics. We empirically quantify them in this section.

Setup. For performance comparision of the three deployment alternatives described in Section 5.4.2, we used a setup consisting of a Floodlight controller running the

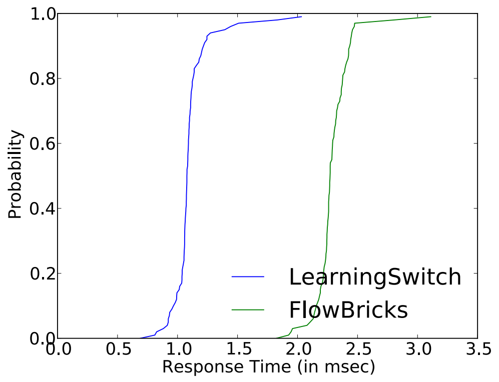
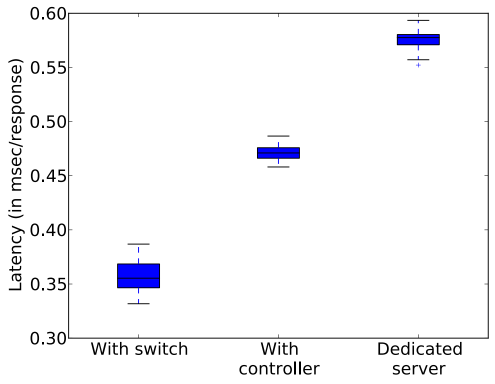
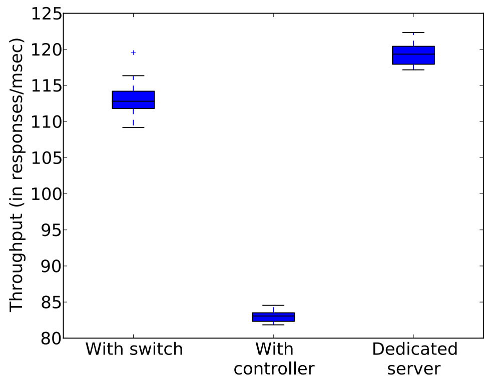


Figure 5.7.: CDF of response time with and without FlowBricks.



(a) Response time.



(b) Throughput.

Figure 5.8.: Performance comparison of deployment scenarios

learning switch service as shown in Figure 5.6. We configured FlowBricks with the following policy: $* : LS : 100$. FlowBricks applies the learning switch service to all traffic in the network. To emulate the network, we using two instances of Cbench [64]. We configured each Cbench instance with the IP address and TCP port of FlowBricks. Cbench generates `Packet-In` message and measures the throughput and response time of the corresponding `Packet-Out` messages from the control plane, which in our case includes both the controller and FlowBricks.

Response Time. To measure the response time, Cbench sends one `Packet-In` messages per switch and waits for a response from the controller. When it receives a response, it measures the response time and it sends another `Packet-In` message. This continues for the duration of the experiment. Figure 5.8(a) shows the response time of the control plane as measured by Cbench. As seen in the figure, deploying FlowBricks and controller on separate servers has the highest response time since each control plane message traverses the network two times. Colocating FlowBricks with the controller reduces the response time. Colocating FlowBricks with the switch, further reduces the response time, since we now have two instances of FlowBricks instead of one.

Throughput. Cbench measures the throughput of the control plane by ensuring that the controller is always processing `Packet-In` messages for the duration of the experiment. The ratio of the messages processed to the duration of the experiment gives the throughput of the controller. Figure 5.8(b) shows the throughput observed in the three scenarios. As expected, running FlowBricks and the controller on separate servers gives the highest throughput. When FlowBricks was colocated with the switches, it reduced the throughput by approximately 10%. This is probably because the Cbench process consumed CPU cycles and thus reduced FlowBricks's throughput. The lowest throughput was observed when FlowBricks was colocated with the controller.

Summary. The above results show that colocating FlowBricks with the switch software leads to the lowest response time. Also, the throughput is comparable to that achieved by running FlowBricks on a dedicated server. For software-switches like Open vSwitch, running an instance of FlowBricks with every instance of the software switch on end hosts is possible. However, it may not be feasible to run FlowBricks on a physical switch unless the switch vendor allows it. So, for physical switches, choosing between a dedicated server for FlowBricks and colocating FlowBricks with the controller involves a trade-off between response time and throughput. For deployments that expect a lot of control plane traffic, the server running the controllers is likely to become a bottleneck. In a such a scenario, deploying controllers and FlowBricks on separate dedicated servers is preferred. However, for deployments where response time to network events is more critical, it may be better to colocate FlowBricks with the controller.

5.6 Related Work

The first SDN controller was single threaded [42]. Since then, more advanced multi-threaded controllers [45, 46] have been developed. More recently, physically distributed SDN controllers [8, 9] have been proposed to handle large networks which are beyond the capability of a single server. This allows the operator to add and remove features at runtime. However, all the above implementations are monolithic controllers which focus on improving performance.

Some earlier work has focused on making controllers more flexible. Beacon [44] allows dynamic addition and removal of controller modules. This allows the operator to add and remove features at runtime. However, all modules need to be written in Java and use Beacon's API. Yanc [65] is a platform which exposes network configuration and state using the file system. Controller applications are separate processes. This allows applications to be written in any language but still requires application vendors to use Yanc's file system layout. HotSwap [66] provides a mechanism to upgrade from one controller version to the next or move between controller vendors. It does so by replaying network events to bring the new and old controllers to a consistent state. However, at a given time, services from just one controller can be applied to the traffic in the network.

Frenetic [53] and Pyretic [28] provide a query language for describing high-level packet-forwarding policies for parallel and serial cases of service integration. Frenetic and Pyretic programs easily integrate services but cannot be generalized across controllers from several vendors.

Our system architecture resembles FlowVisor [54] at a high level. FlowVisor allows a network operator to slice the global flow space and assign a controller to each slice.

FlowVisor needs to match a packet only against the flow tables of the controller of its slice. Since FlowBricks deploys services from multiple controllers on the *same* flows, a given packet needs to be matched against the flow tables of all controllers. This introduces the problem of combining all controllers' flow tables in the datapath, and changes how FlowBricks processes messages which is the focus of this paper.

An SDN hypervisor [55] has been proposed to address the same problem as FlowBricks. It combines policies by calculating the cross product of rules of each policy. As the authors themselves point out, this mechanism does not handle flow table entry timeouts. Calculating the cross product leads to an exponential increase in TCAM space requirement for the combined policy. This could make it infeasible to deploy the combined policy on switches with limited TCAM space. Also, the SDN hypervisor does not support multiple flow tables since it addresses OpenFlow 1.0 which has a single flow table.

5.7 Summary

The SDN paradigm increases the potential for flexible network systems design and implementation. We address the problem of composing services implemented on controllers from different vendors. We introduced a framework to integrate heterogeneous controllers using only the standardized controller to switch communication protocol. To demonstrate the feasibility of this framework, we presented its design using a simple technique to combine flow tables from different OpenFlow-based controllers without modifying the controllers themselves.

6 CONCLUSIONS

In this thesis, we explored techniques to improve the data and control plane performance of data center networks. In particular, we focused on networks that are organized in multi-rooted tree topologies and employ the SDN paradigm. We proposed techniques that are compatible with existing network protocols and can be readily deployed in data centers. We empirically verified that our techniques improve the network performance metrics like throughput and latency and consequently impact application performance too.

We showed how a simple packet-level traffic splitting scheme called RPS not only leads to significantly better load balance and network utilization, but also incurs little packet reordering since it exploits the symmetry in these networks. Furthermore, such schemes have lower complexity and readily implementable, making them an appealing alternative for data center networks. Real data centers also need to deal with failures which may disturb the symmetry, impacting the performance of RPS. We observed that by keeping queue lengths small, this impact can be minimized. We exploited this observation by proposing a simple queue management scheme called SRED that can cope well with failures.

To improve scalability along the control plane, we presented our design of ElastiCon, a distributed elastic SDN controller. We designed and implemented algorithms for switch migration, controller load balancing and elasticity which form the core of the controller. We enhanced Mininet and used it to demonstrate the efficacy of those algorithms.

Finally, we proposed FlowBricks, a framework that allows integration of services running on heterogeneous controllers in a way that is transparent to controllers and does not require any additional standardization beyond a southbound API.

6.1 Future Directions

In this dissertation, we propose and empirically demonstrate that techniques that improve the scalability of data center networks. However, we do not address fault tolerance while proposing these techniques. Also, the ability to easily integrate services independently of existing services in an SDN controller presents the opportunity to develop new services.

6.1.1 Fault Tolerance

Our current design of FlowBricks and ElastiCon does not address issues caused by failures, although we believe fault tolerance mechanisms can easily fit into these architectures. For ElastiCon, this may require running three or more controllers in equal role for each switch and using a consensus protocol between them to ensure there is always at least one master even if the new master crashes. In FlowBricks, we want to explore algorithms for making FlowBricks stateless. If FlowBricks is stateless, a new instance of FlowBricks can be triggered when an instance crashes. However, this would involve changing the southbound API to include some state-information with every action in a flow table entry.

6.1.2 New SDN Services

Going forward we plan to develop controllers with management and monitoring services that can be plugged into FlowBricks at runtime to monitor performance and debug the network. For example, new services can be added to the beginning and end of the policies in FlowBricks. The service in the beginning of the policy would insert new packets in the network using the `Packet-Out` messages and the service at the end could verify that the packet was correctly modified by intermediate services. We also plan to integrate existing techniques [27, 59] into FlowBricks to guarantee consistency and correctness of composed services.

REFERENCES

REFERENCES

[1] Albert Greenberg, James R. Hamilton, Navendu Jain, Srikanth Kandula, Changhoon Kim, Parantap Lahiri, David A. Maltz, Parveen Patel, and Sudipta Sengupta. VL2: A Scalable and Flexible Data Center Network. In *Proceedings of the ACM SIGCOMM 2009 Conference on Data Communication*, SIGCOMM '09, pages 51–62, New York, NY, USA, 2009. ACM.

[2] Mohammad Al-Fares, Alexander Loukissas, and Amin Vahdat. A Scalable, Commodity Data Center Network Architecture. In *Proceedings of the ACM SIGCOMM 2008 Conference on Data Communication*, SIGCOMM '08, pages 63–74, New York, NY, USA, 2008. ACM.

[3] Srikanth Kandula, Sudipta Sengupta, Albert Greenberg, Parveen Patel, and Ronnie Chaiken. The Nature of Data Center Traffic: Measurements & Analysis. In *Proceedings of the 9th ACM SIGCOMM Conference on Internet Measurement Conference*, IMC '09, pages 202–208, New York, NY, USA, 2009. ACM.

[4] Theophilus Benson, Ashok Anand, Aditya Akella, and Ming Zhang. Understanding Data Center Traffic Characteristics. In *Proceedings of the 1st ACM Workshop on Research on Enterprise Networking*, WREN '09, pages 65–72, New York, NY, USA, 2009. ACM.

[5] Mohammad Al-Fares, Sivasankar Radhakrishnan, Barath Raghavan, Nelson Huang, and Amin Vahdat. Hedera: Dynamic Flow Scheduling for Data Center Networks. In *Proceedings of the 7th USENIX Symposium on Networked Systems Design and Implementation*, NSDI '10, pages 281–296, Berkeley, CA, USA, 2010. USENIX Association.

[6] Andrew R. Curtis, Wonho Kim, and Praveen Yalagandula. Mahout: Low-Overhead Datacenter Traffic Management using End-Host-Based Elephant Detection. In *Proceedings of the 30th IEEE International Conference on Computer Communications*, INFOCOM '11, pages 1629–1637. IEEE, 2011.

[7] Costin Raiciu, Sebastien Barre, Christopher Pluntke, Adam Greenhalgh, Damon Wischik, and Mark Handley. Improving Datacenter Performance and Robustness with Multipath TCP. In *Proceedings of the ACM SIGCOMM 2012 Conference on Data Communication*, SIGCOMM '11, pages 266–277, New York, NY, USA, 2011. ACM.

[8] Teemu Koponen, Martin Casado, Natasha Gude, Jeremy Stribling, Leon Poutievski, Min Zhu, Rajiv Ramanathan, Yuichiro Iwata, Hiroaki Inoue, Takayuki Hama, and Scott Shenker. Onix: A Distributed Control Platform for Large-scale Production Networks. In *Proceedings of the 9th USENIX Symposium on Operating Systems Design and Implementation*, OSDI '10, pages 351–364, Berkeley, CA, USA, 2010. USENIX Association.

[9] Amin Tootoonchian and Yashar Ganjali. HyperFlow: A Distributed Control Plane for OpenFlow. In *Proceedings of the 2010 Internet Network Management Conference on Research on Enterprise Networking*, INM/WREN '10, pages 3–3, Berkeley, CA, USA, 2010. USENIX Association.

[10] Dan Levin, Andreas Wundsam, Brandon Heller, Nikhil Handigol, and Anja Feldmann. Logically Centralized?: State Distribution Trade-offs in Software Defined Networks. In *Proceedings of the 1st ACM SIGCOMM Workshop on Hot Topics in Software Defined Networks*, HotSDN '12, pages 1–6, New York, NY, USA, 2012. ACM.

[11] Theophilus Benson, Aditya Akella, and David A. Maltz. Network Traffic Characteristics of Data Centers in the Wild. In *Proceedings of the 10th ACM SIGCOMM Conference on Internet Measurement*, IMC '10, pages 267–280, New York, NY, USA, 2010. ACM.

[12] Sushant Jain, Alok Kumar, Subhasree Mandal, Joon Ong, Leon Poutievski, Arjun Singh, Subbaiah Venkata, Jim Wanderer, Junlan Zhou, Min Zhu, Jon Zolla, Urs Hölzle, Stephen Stuart, and Amin Vahdat. B4: Experience with a Globally-deployed Software Defined Wan. In *Proceedings of the ACM SIGCOMM 2013 Conference on Data Communication*, SIGCOMM '13, pages 3–14, New York, NY, USA, 2013. ACM.

[13] Sally Floyd and Van Jacobson. Random Early Detection Gateways for Congestion Avoidance. *IEEE/ACM Transactions on Networking*, 1(4):397–413, August 1993.

[14] Doug Beaver, Sanjeev Kumar, Harry C. Li, Jason Sobel, and Peter Vajgel. Finding a Needle in Haystack: Facebook's Photo Storage. In *Proceedings of the 9th USENIX Conference on Operating Systems Design and Implementation*, OSDI'10, pages 1–8, Berkeley, CA, USA, 2010. USENIX Association.

[15] Ajay Gulati, Anne Holler, Minwen Ji, Ganesha Shanmuganathan, Carl Waldspurger, and Xiaoyun Zhu. VMware Distributed Resource Management: Design, Implementation and Lessons Learned. Technical report, VMWare, Inc, Palo Alto, California, 2012.

[16] Rishi Kapoor, George Porter, Malveeka Tewari, Geoffrey M. Voelker, and Amin Vahdat. Chronos: Predictable Low Latency for Data Center Applications. In *Proceedings of the 3rd ACM Symposium on Cloud Computing*, SoCC '12, pages 9:1–9:14, New York, NY, USA, 2012. ACM.

[17] George Porter, Richard Strong, Nathan Farrington, Alex Forencich, Pang Chen-Sun, Tajana Rosing, Yeshaiahu Fainman, George Papen, and Amin Vahdat. Integrating Microsecond Circuit Switching into the Data Center. In *Proceedings of the ACM SIGCOMM 2013 Conference on Data Communication*, SIGCOMM '13, pages 447–458, New York, NY, USA, 2013. ACM.

[18] Chuanxiong Guo, Guohan Lu, Dan Li, Haitao Wu, Xuan Zhang, Yunfeng Shi, Chen Tian, Yongguang Zhang, and Songwu Lu. BCube: A High Performance, Server-centric Network Architecture for Modular Data Centers. In *Proceedings of the ACM SIGCOMM 2009 Conference on Data Communication*, SIGCOMM '09, pages 63–74, New York, NY, USA, 2009. ACM.

[19] Jeffrey Dean and Sanjay Ghemawat. MapReduce: Simplified Data Processing on Large Clusters. *Communications of the ACM*, 51(1):107–113, January 2008.

[20] Haitao Wu, Zhenqian Feng, Chuanxiong Guo, and Yongguang Zhang. ICTCP: Incast Congestion Control for TCP in Data Center Networks. In *Proceedings of the 6th International Conference on Emerging Networking Experiments and Technologies*, Co-NEXT '10, pages 13:1–13:12, New York, NY, USA, 2010. ACM.

[21] Mohammad Alizadeh, Albert Greenberg, David A. Maltz, Jitendra Padhye, Parveen Patel, Balaji Prabhakar, Sudipta Sengupta, and Murari Sridharan. Data Center TCP (DCTCP). In *Proceedings of the ACM SIGCOMM 2010 Conference on Data Communication*, SIGCOMM '10, pages 63–74, New York, NY, USA, 2010. ACM.

[22] Balajee Vamanan, Jahangir Hasan, and T.N. Vijaykumar. Deadline-aware Datacenter TCP (D2TCP). In *Proceedings of the ACM SIGCOMM 2012 Conference on Data Communication*, SIGCOMM '12, pages 115–126, New York, NY, USA, 2012. ACM.

[23] Nick McKeown, Tom Anderson, Hari Balakrishnan, Guru Parulkar, et al. OpenFlow: Enabling Innovation in Campus Networks. *SIGCOMM Computer Communication Review*, 38(2), March 2008.

[24] Minlan Yu, Lavanya Jose, and Rui Miao. Software Defined Traffic Measurement with OpenSketch. In *Proceedings of the 10th USENIX Symposium on Networked Systems Design and Implementation*, NSDI '13, pages 29–42, Berkeley, CA, USA, 2013. USENIX Association.

[25] Vimalkumar Jeyakumar, Mohammad Alizadeh, Yilong Geng, Changhoon Kim, and David Mazires. Millions of Little Minions: Using Packets for Low Latency Network Programming and Visibility. In *Proceedings of the ACM SIGCOMM 2014 Conference on Data Communication*, SIGCOMM '14, pages 3–14, New York, NY, USA, 2014. ACM.

[26] Nikhil Handigol, Brandon Heller, Vimalkumar Jeyakumar, David Mazières, and Nick McKeown. I Know What Your Packet Did Last Hop: Using Packet Histories to Troubleshoot Networks. In *Proceedings of the 11th USENIX Symposium on Networked Systems Design and Implementation*, NSDI '14, pages 71–85, Berkeley, CA, USA, 2014. USENIX Association.

[27] Ahmed Khurshid, Xuan Zou, Wenxuan Zhou, Matthew Caesar, and P. Brighten Godfrey. VeriFlow: Verifying Network-wide Invariants in Real Time. In *Proceedings of the 10th USENIX Conference on Networked Systems Design and Implementation*, NSDI '13, pages 15–28, Berkeley, CA, USA, 2013. USENIX Association.

[28] Christopher Monsanto, Joshua Reicha, Nate Foster, Jennifer Rexford, and David Walker. Composing Software-Defined Networks. In *Proceedings of the 10th USENIX Symposium on Networked Systems Design and Implementation*, NSDI '13, pages 1–13, Berkeley, CA, USA, 2013. USENIX Association.

[29] Zafar Ayyub Qazi, Cheng-Chun Tu, Luis Chiang, Rui Miao, Vyas Sekar, and Minlan Yu. SIMPLE-fying Middlebox Policy Enforcement Using SDN. In *Proceedings of the ACM SIGCOMM 2013 Conference on Data Communication*, SIGCOMM '13, pages 27–38, New York, NY, USA, 2013. ACM.

[30] Per packet load balancing. `http://www.cisco.com/en/US/docs/ios/12_0s/feature/guide/pplb.html`. Accessed July 2012.

[31] M. Laor and L. Gendel. The Effect of Packet Reordering in a Backbone Link on Application Throughput. *Network, IEEE*, 16(5):28–36, sep 2002.

[32] Mohammad Alizadeh, Abdul Kabbani, Tom Edsall, Balaji Prabhakar, Amin Vahdat, and Masato Yasuda. Less Is More: Trading a Little Bandwidth for Ultra-Low Latency in the Data Center. In *Proceedings of the 9th USENIX Symposium on Networked Systems Design and Implementation*, NSDI '12, pages 253–266, Berkeley, CA, USA, 2012. USENIX Association.

[33] Ethan Blanton and Mark Allman. Using TCP DSACKs and SCTP Duplicate Transmission Sequence Numbers (TSNs) to Detect Spurious Retransmissions. Request for Comments (Experimental) 3708, Internet Engineering Task Force, February 2004.

[34] Sumitha Bhandarkar, A. L. Narasimha Reddy, Mark Allman, and Ethan Blanton. Improving the Robustness of TCP to Non-Congestion Events. Request for Comments (Experimental) 4653, Internet Engineering Task Force, August 2006.

[35] David Zats, Tathagata Das, Prashanth Mohan, Dhruba Borthakur, and Randy Katz. DeTail: Reducing the Flow Completion Time Tail in Datacenter Networks. In *Proceedings of the ACM SIGCOMM 2012 Conference on Data Communication*, SIGCOMM '12, pages 139–150, New York, NY, USA, 2012. ACM.

[36] Sebastien Barr. MultiPath TCP in the Linux Kernel. `https://scm.info.ucl.ac.be/trac/mptcp/wiki/install`. Accessed July 2012.

[37] K. K. Ramakrishnan, Sally Floyd, and David Black. The Addition of Explicit Congestion Notification (ECN) to IP. Request for Comments (Proposed Standard) 3168, Internet Engineering Task Force, September 2001.

[38] Albert Greenberg, Parantap Lahiri, David A. Maltz, Parveen Patel, and Sudipta Sengupta. Towards a Next Generation Data Center Architecture: Scalability and Commoditization. In *Proceedings of the ACM Workshop on Programmable Routers for Extensible Services of Tomorrow*, PRESTO '08, pages 57–62, New York, NY, USA, 2008. ACM.

[39] Shan Sinha, Srikanth Kandula, and Dina Katabi. Harnessing TCPs Burstiness using Flowlet Switching. In *Proceedings of the 3rd ACM Workshop on Hot Topics in Networks*, HotNets-III, New York, NY, USA, 2004. ACM.

[40] Ping Pan and Thomas Nadeau. Software-Defined Network (SDN) Problem Statement and Use Cases for Data Center Applications. Internet-Draft (Standards Track), Internet Engineering Task Force, March 2012.

[41] Andrew R. Curtis, Jeffrey C. Mogul, Jean Tourrilhes, Praveen Yalagandula, Puneet Sharma, and Sujata Banerjee. DevoFlow: Scaling Flow Management for High-performance Networks. In *Proceedings of the ACM SIGCOMM 2011 Conference on Data Communication*, SIGCOMM '11, pages 254–265, New York, NY, USA, 2011. ACM.

[42] Natasha Gude, Teemu Koponen, Justin Pettit, Ben Pfaff, Martín Casado, Nick McKeown, and Scott Shenker. NOX: Towards an Operating System for Networks. *SIGCOMM Computer Commununication Review*, 38(3):105–110, 2008.

[43] Amin Tootoonchian, Sergey Gorbunov, Yashar Ganjali, Martin Casado, and Rob Sherwood. On Controller Performance in Software-Defined Networks. In *Proceedings of the 2nd USENIX Workshop on Hot Topics in Management of Internet, Cloud, and Enterprise Networks and Services*, HotICE '12, Berkeley, CA, 2012. USENIX Association.

[44] David Erickson. The Beacon OpenFlow Controller. In *Proceedings of the 2nd ACM SIGCOMM Workshop on Hot Topics in Software Defined Networks*, HotSDN '13, pages 13–18, New York, NY, USA, 2013. ACM.

[45] Z. Cai, A. L. Cox, and T. S. E. Ng. Maestro: A System for Scalable OpenFlow Control. Technical report, Computer Science Department, Rice University, Houston, Texas, 2010.

[46] Floodlight. `http://www.projectfloodlight.org/floodlight/`. Accessed March 2013.

[47] OpenDaylight. `http://www.opendaylight.org/`. Accessed May 2014.

[48] Open Networking Foundation. OpenFlow Management and Configuration Protocol (OF-Config 1.1), June 2012.

[49] Hyperic SIGAR API. `http://www.hyperic.com/products/sigar`. Accessed May 2014.

[50] Nikhil Handigol, Brandon Heller, Vimalkumar Jeyakumar, Bob Lantz, and Nick McKeown. Reproducible Network Experiments Using Container-based Emulation. In *Proceedings of the 8th International Conference on Emerging Networking Experiments and Technologies*, CoNEXT '12, pages 253–264, New York, NY, USA, 2012. ACM.

[51] Ben Pfaff, Justin Pettit, Keith Amidon, Martin Casado, Teemu Koponen, and Scott Shenker. Extending Networking into the Virtualization Layer. In *Proceedings of the 8th ACM Workshop on Hot Topics in Networks*, HotNets-VIII, pages 1–6, New York, NY, USA, 2009. ACM.

[52] Barath Raghavan, Martín Casado, Teemu Koponen, Sylvia Ratnasamy, Ali Ghodsi, and Scott Shenker. Software-defined Internet Architecture: Decoupling Architecture from Infrastructure. In *Proceedings of the 11th ACM Workshop on Hot Topics in Networks*, HotNets-XI, pages 43–48, New York, NY, USA, 2012. ACM.

[53] Nate Foster, Rob Harrison, Michael J. Freedman, Christopher Monsanto, Jennifer Rexford, Alec Story, and David Walker. Frenetic: A Network Programming Language. In *Proceedings of the 16th ACM SIGPLAN International Conference on Functional Programming*, ICFP '11, pages 279–291, New York, NY, USA, 2011. ACM.

[54] Rob Sherwood, Glen Gibb, Kok-Kiong Yap, Guido Appenzeller, Martin Casado, Nick McKeown, and Guru M. Parulkar. Can the Production Network Be the Testbed? In *Proceedings of the 9th USENIX Symposium on Operating Systems Design and Implementation*, OSDI '10, pages 365–378, Berkeley, CA, USA, 2010. USENIX Association.

[55] X. Jin, J. Rexford, and D. Walker. Incremental Update for a Compositional SDN Hypervisor. In *Proceedings of the 3rd ACM SIGCOMM Workshop on Hot Topics in Software Defined Networks*, HotSDN '14, pages 187–192, New York, NY, USA, 2014. ACM.

[56] Open Networking Foundation. OpenFlow Switch Specification (Version 1.1.0), February 2011.

[57] Carolyn Jane Anderson, Nate Foster, Arjun Guha, Jean-Baptiste Jeannin, Dexter Kozen, Cole Schlesinger, and David Walker. NetKAT: Semantic Foundations for Networks. In *Proceedings of the 41st ACM SIGPLAN-SIGACT Symposium on Principles of Programming Languages*, POPL '14, pages 113–126, New York, NY, USA, 2014. ACM.

[58] Open Networking Foundation. OpenFlow Switch Specification (Version 1.4.0), October 2013.

[59] Mark Reitblatt, Nate Foster, Jennifer Rexford, Cole Schlesinger, and David Walker. Abstractions for Network Update. In *Proceedings of the ACM SIGCOMM 2012 Conference on Data Communication*, SIGCOMM '12, pages 323–334, New York, NY, USA, 2012. ACM.

[60] Alexander Kesselman, Kirill Kogan, Sergey Nemzer, and Michael Segal. Space and Speed Tradeoffs in TCAM Hierarchical Packet Classification. *Journal of Computer System Sciences*, 79(1):111–121, February 2013.

[61] Kirill Kogan, Sergey Nikolenko, Ori Rottenstreich, William Culhane, and Patrick Eugster. SAX-PAC (Scalable And eXpressive PAcket Classification). In *Proceedings of the 2014 ACM Conference on SIGCOMM*, SIGCOMM '14, pages 15–26, New York, NY, USA, 2014. ACM.

[62] Henry Van Styn. Tcpdump Fu. *Linux Journal*, 2011(210):90–97, October 2011.

[63] M. Kobayashi, S. Seetharamn, G. Parulkar, G. Appenzeller, J. Little, J. van Reijendam, P. Weissmann, and N. McKeown. Maturing of OpenFlow and Software-Defined Networking through Deployments. *Computer Networks*, 61(0):151–175, March 2014.

[64] Cbench. `http://www.openflowhub.org/display/floodlightcontroller/Cbench+(New)`. Accessed September 2014.

[65] Matthew Monaco, Oliver Michel, and Eric Keller. Applying Operating System Principles to SDN Controller Design. In *Proceedings of the 12th ACM Workshop on Hot Topics in Networks*, HotNets-XII, pages 2:1–2:7, New York, NY, USA, 2013. ACM.

[66] Laurent Vanbever, Joshua Reich, Theophilus Benson, Nate Foster, and Jennifer Rexford. HotSwap: Correct and Efficient Controller Upgrades for Software-Defined Networks. In *Proceedings of the 2nd ACM SIGCOMM Workshop on Hot Topics in Software Defined Networking*, HotSDN '13, pages 133–138, New York, NY, USA, 2013. ACM.

VITA

VITA

Advait Abhay Dixit received his B.Tech in computer science and engineering from Indian Institute of Technology, Guwahati, India in 2003. He received his M.S. in computer science from the University of California, Los Angeles in 2004 where his research focused on sensor networks. He started his graduate studies in the Computer Science Department at Purdue University in 2010, where he worked on various aspects of data center networking. During the course of his graduate studies, he interned at Google Inc, NEC Labs America and Bell Labs. He spent one year as a research assistant in the Rosen Center for Advanced Computing working on grid computing applications while pursuing his Ph.D.