Dashing and Star: Byzantine Fault Tolerance with Weak Certificates

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Abstract

State-of-the-art Byzantine fault-tolerant (BFT) protocols assuming partial synchrony such as SBFT and HotStuff use *regular certificates* obtained from $2f+1$ (partial) signatures. We show that one can use *weak certificates* obtained from only $f+1$ signatures to *assist* in designing more robust and more efficient BFT protocols. We design and implement two BFT systems: Pichu (a family of two HotStuff-style BFT protocols) and Sonic (a parallel BFT framework).

We first present Pichu1 that targets both efficiency and robustness using weak certificates. Pichu1 is also network-adaptive in the sense that it can leverage network connection discrepancy to improve performance. We show that Pichu1 outperforms HotStuff in various failure-free and failure scenarios. We then present Pichu2 enabling a *one-phase* fast path by using *strong certificates* from $3f+1$ signatures.

We then leverage weak certificates to build Sonic, a highly scalable BFT framework that delivers transactions from $n-f$ replicas. Sonic compares favorably with existing protocols in terms of liveness, communication, state transfer, scalability, and/or robustness under failures.

We demonstrate that Pichu achieves 47%-107% higher peak throughput than HotStuff for experiments on Amazon EC2. Meanwhile, unlike all known BFT protocols whose performance degrades as f grows large, the peak throughput of Sonic increases as f grows. When deployed in a WAN with 91 replicas across five continents, Sonic achieves an impressive throughput of 256 ktx/sec, 2.38x that of Narwhal.

1 Introduction

Byzantine fault-tolerant state machine replication (BFT) is known as the core building block for permissioned blockchains [5, 6, 14, 29, 30, 34, 37]. This paper focuses on highly efficient, partially synchronous BFT protocols [10, 19]. Almost universally, these protocols rely critically on *regular (quorum) certificates* which, roughly speaking, are sets with at least $2f+1$ messages from different replicas. Recent protocols such as SBFT [25] and HotStuff [39] require using (threshold) signatures for regular certificates as transferable proofs.

This paper demonstrates that one can build BFT systems that outperform existing ones—in one way or another—by using *weak certificates* with at least $f+1$ signatures from different replicas.

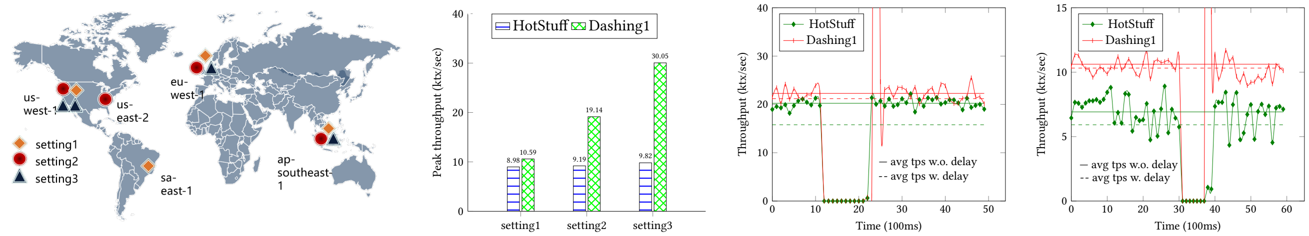
Intuitively, weak certificates may lead to more efficient BFT protocols, because replicas only need to wait for signatures from $f+1$ replicas and combine only $f+1$ signature shares. Indeed, as shown in prior works (e.g., [18]), Byzantine agreement protocols with the $f+1$ threshold can be (much) more efficient than their counterparts with the $2f+1$ threshold. *This paper explores novel usages of weak certificates much beyond this intuition.*

1.1 Dashing: Gaining in Efficiency, Network Adaptivity, and Robustness

In Dashing, we challenge the conventional wisdom and offer new insights into the design of BFT protocols.

• **Using weak certificates.** It is well-known that BFT protocols need to use regular certificates to ensure liveness and safety. So far, weak certificates do not appear to be helpful in building faster BFT protocols. Our first goal is to challenge the intuition and provide a way to exploit weak certificates to assist in the BFT design.

• **Leveraging network connection discrepancy.** When designing and evaluating a partially synchronous BFT, we implicitly assume the simplistic network configuration, where replicas communicate with each other with about the same latency (either all in LANs or WANs). But in practice, the latency discrepancy among different replicas naturally exists. A realistic scenario is that some replicas (say, 1/3 of the replicas) naturally have better connections than the rest of them. This fact is overlooked by existing BFT protocols. We experimentally show in Fig. 1 that HotStuff does not exhibit visible



(a) Distributions of replicas on Amazon EC2 instances.
(b) Peak throughput under three settings for $f = 1$.
(c) Throughput under one-second unexpected delay for $f = 1$.
(d) Throughput under one-second unexpected delay for $f = 10$.

Figure 1. Throughput of HotStuff and Dashing1 in three different settings on Amazon EC2.

performance differences even if we place some replicas in the same region. The result is somewhat expected: intuitively, the safety of BFT depends on the overall BFT network, so the performance of BFT should depend on the overall BFT network. Again, we challenge this intuition, showing that BFT can benefit from network connection discrepancies.

• **Useful work during asynchrony.** Partially synchronous BFT protocols cannot make progress during asynchrony. They would simply wait until the network becomes synchronous (before view change/leader election occurs) or loop on view changes until a correct leader is selected—in either case, no meaningful progress can be made. The situation is only exacerbated, if the network is intermittently synchronous or adaptively manipulated [32]. Naturally, it seems that there is nothing we can do about the situation: existing partially synchronous BFT protocols are deterministic and subject to the celebrated FLP impossibility result [20]. We take a fresh look at the problem: while one indeed cannot make progress during asynchrony, we do not waste our computation and network bandwidth during asynchrony. We perform "useful" operations such that once the network becomes synchronous, we can commit a large number of cumulative transactions—the longer the asynchrony, the more transactions committed—in some sense, the "best" that one could anticipate.

Dashing1. In Dashing1, we attempt to use weak certificates instead of regular certificates *as much as possible*—during the normal case, during transient failures or network interruptions, during unresponsive replicas (e.g., crashes, slow replicas), and during view changes. Transforming the idea into a fully secure BFT protocol, however, is tricky: we tackled subtle safety and liveness challenges within a view and across views due to the usage of weak certificates. Correspondingly, Dashing1 gains in efficiency and robustness in various scenarios, including during normal cases and across views, and in the presence of transient network interruptions, network connection discrepancies, or unresponsive failures.

As shown in Fig. 1a, we deploy HotStuff and Dashing1 on Amazon EC2 (for $n = 4$) in three different settings: in setting 1, the four replicas are distributed over four continents; in setting 2 and setting 3, we place two replicas in closer locations. In all three settings (Fig. 1b), we find Dashing1 consistently outperforms HotStuff; in setting 2 and setting 3, Dashing1 achieves about 2x and 3x the throughput of HotStuff, respectively. The experiments show that Dashing1 substantially improved performance in the normal case and in the presence of (natural) network connection discrepancies.

We also run experiments for Dashing1 and HotStuff with 1,200 clients in a WAN setting with 4 replicas (Fig. 1c) and 31 replicas (Fig. 1d), respectively. In the experiments, we inject a one-second network delay at $2f$ replicas using the *tc* traffic control command. We report the throughput for a duration of five seconds and six seconds for the experiment with 4 replicas and 31 replicas, respectively. In both experiments, while neither HotStuff nor Dashing1 can make progress during the network delay, the throughput of Dashing1 reaches roughly 10x that of HotStuff when the network recovers. For the experiments for $f = 10$, the average throughput of Dashing1 is 79.3% and 49.1% higher than that of HotStuff with the unexpected network delay (dashed line) and without delay (solid line), respectively. Moreover, Dashing1 achieves roughly the same average throughput as that without delay, while we witness a more visible decrease in throughput for HotStuff. Indeed, *all* transactions corresponding to the wQCs are delivered all at once after the network resumes in Dashing1; in contrast, HotStuff cannot make any progress before the network resumes. Note the throughput gain depends on the network-induced downtime.

Dashing2. We show how to enable a one-phase fast path by leveraging *strong certificates* from $3f + 1$ signatures in our BFT protocols. We demonstrate that such a task is technically challenging—being more subtle than that in SBFT [25]—and offer a secure and efficient solution.

1.2 Star: Gaining in Efficiency and Scalability

We use weak certificates to help design Star, a scalable BFT framework that delivers transactions from $n - f$ replicas using only a single consensus instance. As shown in Table 1, Star has improved prior protocols in terms of message, communication, and state transfer complexity, while achieving

protocols	QC type	message	communication	state transfer	quality
ISS (PBFT) [36]	rQC	$O(n^3)$	$O(Ln^2 + \lambda n^3)$	$\boldsymbol{O(1)}$	yes
Narwhal [15]	rQC	$O(n^3)$	$O(Ln^2 + \lambda n^3)$	$O(k)$	yes
Dumbo-NG [21]	rQC	$\boldsymbol{O(n^2)}$	$\boldsymbol{O(Ln^2 + \lambda n^2)}$	$O(k)$	no
Star (this work)	**wQC**	$\boldsymbol{O(n^2)}$	$\boldsymbol{O(Ln^2 + \lambda n^2)}$	$\boldsymbol{O(1)}$	yes

Table 1. L is the proposal size for each replica and λ is the security parameter. Narwhal provides a variant reducing the messages to $O(n^2)$ but the communication remains $O(Ln^2 + \lambda n^3)$. State transfer denotes the time to obtain a transaction proposed k epochs ago. Quality means if at least a fraction of the transactions in a committed block are from correct replicas.

standard liveness and quality (meaning that at least a non-negligible fraction of the total transactions in a committed block are from correct replicas) guarantees.

More concretely, while Star inherits the architecture that separates bulk data transmission from consensus such that these two processes can be run independently [15], Star uses the more efficient weak certificates for the data transmission layer and importantly, such a layer can be effectively pipelined and provides more efficient communication and state transfer. Moreover, Star associates the layers using an increasing epoch number, which allows us to achieve a strong blockchain quality property.

Simply using PBFT [12] in our underlying consensus layer, the throughput of Star strictly keeps increasing as n grows. When deploying Star and Narwhal [15] (the state-of-the-art protocol) in a WAN with 91 replicas across five continents (Fig. 7o), Star achieves a throughput of 256 ktx/sec, 2.38x that of Narwhal.

1.3 Summary of Contributions

- We design a family of Dashing protocols—Dashing1 and Dashing2—using weak certificates. In particular, Dashing1 gains in improved efficiency and robustness in both failure and failure-free scenarios and in normal cases and across views; unlike prior protocols, Dashing1 excels in performance with transient network interruptions and network connection discrepancies. Dashing2 enables a one-phase fast path for Dashing1 and offers improved latency.
- We provide a new parallel BFT framework (Star) achieving reduced communication and state transfer time and being more scalable than prior ones.
- We implement the BFT protocols (the two Dashing protocols and a Star instantiation). We performed extensive evaluations of the protocols, showing that our protocols outperform existing protocols in various metrics.

2 Related Work

Dashing vs. PBFT. Conventional protocols such as PBFT [11] allow running multiple consensus instances in parallel: a leader can propose new transactions even if previous ones have not been prepared at this point. This makes PBFT look relevant to Dashing: both take approaches to fully utilize the bandwidth. However, in each instance in PBFT, replicas cannot deliver any block if they fail to receive $2f + 1$ matching votes. In Dashing1, replicas may make progress if the leader receives $f + 1$ matching votes (wQCs). Hence, transactions corresponding to wQCs may be delivered in Dashing even if no correct replica receives $2f + 1$ matching votes.

Why pick Narwhal for performance comparison? As shown in [24], Bullshark and Narwhal share almost identical throughput in normal cases, and BullShark offers almost 2x the throughput of Mir-BFT at the same latency. While Tusk appears to have slightly better performance than Narwhal, Tusk does not implement the underlying common coin protocol—a key performance bottleneck.

Detailed comparison between Star and existing protocols. DAG-based purely asynchronous BFT protocols have liveness problems. DAG Rider [28] requires unbounded memory for liveness, while Bullshark [24] and Tusk [15] achieve weak liveness (assuming some form of synchrony).

Concurrently, Dumbo-NG [21] is proposed as an asynchronous protocol. Indeed, while we instantiate Star using a partially synchronous one, Star can be asynchronous if the underlying BFT is asynchronous. However, Dumbo-NG does not use any of the following techniques for efficiency or blockchain quality: 1) weak certificates for better efficiency; 2) associating transmission layer and consensus layer with epoch numbers for better blockchain quality; 3) a constant-time state transfer. In particular, without associating transmission layer with consensus layer, a specific transaction can be delayed or censored due to faster commitments of transactions from faulty replicas. Indeed, faulty replicas can form a long chain with an unbounded number of certificates. Thus, a valid transaction may be processed only after all transactions from faulty replicas are committed; moreover, the fraction of transactions from correct replicas in a block may be made arbitrarily small. Last, Dumbo-NG requires unbounded memory for liveness.

Different from ISS [36] (and Mir-BFT [35]) requiring running n parallel consensus for each epoch, Star only needs a single consensus protocol. ISS relies on a Byzantine failure detector to ensure safety and liveness and replicas need to wait for the slowest consensus instance to terminate (possibly with view changes or until timers run out) before processing transactions; in contrast, Star can process transactions once the single consensus instance completes. Also, Star achieves $O(n^2)$ messages, in contrast to ISS with $O(n^3)$ messages. Last, with crash failures, the throughput of ISS and Mir-BFT may drop to 0 for a long duration; they need to run reconfiguration mechanisms to exclude faulty replicas [35, 36].

BFT with trusted hardware. Using trusted hardware, BFT protocols can use $2f+1$ replicas to tolerate f faulty replicas, and $f+1$ matching votes form a quorum certificate [13, 16, 17, 27]. Instead, our protocols are conventional BFT protocols assuming $n \geq 3f+1$, and use weak certificates for efficiency or robustness.

Separating agreement from execution. The architecture by Yin et al. [38] separates BFT agreement replicas from execution replicas. In contrast, Sonic separates transaction dissemination from agreement to improve the performance and scalability of the system.

3 System Model

BFT. This paper studies Byzantine fault-tolerant state machine replication (BFT) protocols. In a BFT protocol, clients *submit* transactions (requests) and replicas *deliver* them. The client obtains a final response to the submitted transaction from the replica responses. A BFT system with n replicas tolerates $f \leq \lfloor \frac{n-1}{3} \rfloor$ Byzantine failures. The correctness of a BFT protocol is specified as follows:

- **Safety**: If a correct replica *delivers* a transaction tx before *delivering* tx', then no correct replica *delivers* a transaction tx' without first *delivering* tx.
- **Liveness**: If a transaction tx is *submitted* to all correct replicas, then all correct replicas eventually *deliver* tx.

We also need an equivalent primitive, atomic broadcast, as a building block. Atomic broadcast is only syntactically different from BFT. In atomic broadcast, a replica *a-broadcasts* messages and all replicas *a-deliver* messages.

- **Safety**: If a correct replica *a-delivers* a message m before *a-delivering* m', then no correct replica *a-delivers* a message m' without first *a-delivering* m.
- **Liveness**: If a correct replica *a-broadcasts* a message m, then all correct replicas eventually *a-deliver* m.

Note that when describing atomic broadcast, we restrict its API in the sense that only a single replica a-broadcasts a message. One can alternatively allow all replicas to a-broadcast transactions as in asynchronous protocols.

This paper mainly considers the partially synchronous model [19], where there exists an unknown global stabilization time (GST) such that after GST, messages sent between two correct replicas arrive within a fixed delay. One of our protocols (Star) works in purely asynchronous environments if the underlying atomic broadcast is asynchronous.

Cryptographic building blocks. We define a (t,n) threshold signature scheme with the following algorithms (*tgen*, *tsign*, *tcombine*, *tverify*). *tgen* outputs a threshold signature public key and a vector of n private keys. A signature signing algorithm *tsign* takes as input a message m and a private key sk_i and outputs a partial signature σ_i. A combining algorithm *tcombine* takes as input pk, a message m, and a set of t valid partial signatures, and outputs a signature σ. A signature verification algorithm *tverify* takes as input pk, a message m, and a signature σ, and outputs a single bit. We require the robustness and unforgeability properties for threshold signatures. When describing the algorithms, we leave the verification of partial signatures and threshold signatures implicit. Dedicated threshold signatures can be realized using pairings [7, 8]. One can also use a group of conventional signatures to build a (t,n) threshold signature for efficiency, as used in various libraries such as HotStuff [2, 39], Jolteon and Ditto [22], and Wendy [23]. The approach is also preferred for our protocols, as many of our protocols have more than one threshold. (Otherwise, one should use different threshold signatures for different thresholds.)

Byzantine quorums and quorum certificates. We consider a system with n replicas, of which at most f are Byzantine faulty. We assume $n \geq 3f+1$ for our protocols, but for simplicity, let $n = 3f+1$. A Byzantine quorum consists of $\lceil \frac{n+f+1}{2} \rceil$ replicas, or simply $2f+1$ if $n = 3f+1$. We call it a *regular quorum*. Slightly abusing notation, we additionally define two different types of quorums: a *weak quorum* consisting of $f+1$ replicas and a *strong quorum* consisting of $n = 3f+1$ replicas. A message with signatures signed by a weak quorum, a regular quorum, and a strong quorum is called a *weak (quorum) certificate* (wQC), a *regular (quorum) certificate* (rQC), and a *strong (quorum) certificate* (sQC), respectively. A certificate can be a threshold signature with a threshold t or a set of t digital signatures.

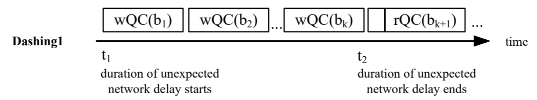
4 The Family of Dashing Protocols

4.1 Overview of (Chained) HotStuff

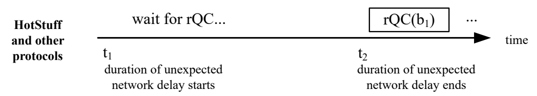
HotStuff describes the syntax of leader-based BFT replication using the language of trees over blocks for leader-based protocols. Here we use a slightly more general notation, where multiple blocks, rather than just one block, may be delivered within a view until a view change.

Each replica stores a tree of blocks. Each block b contains a hash pointer pl to its parent block. A branch led by a given block b is the path from b all the way to the root of the tree (i.e., the *genesis* block). The *height* for b is the number of blocks on the branch led by b. A block b' is an extension of block b, if b is on the branch led by b'. Two blocks are conflicting if neither is an extension of the other. During the protocol, a monotonically growing branch becomes committed. A safe BFT ensures that no two correct replicas commit two conflicting blocks.

HotStuff uses three phases (*prepare phase*, *precommit phase*, and *commit phase*) to deliver a block. In the prepare phase, the leader broadcasts a proposal (a block) b to all replicas and waits for signed responses (also called votes) from a quorum of $n-f$ replicas to form a threshold signature as a quorum certificate (*prepareQC*). In the precommit phase, the leader broadcasts *prepareQC* and waits for responses to form *precommitQC*. In the commit phase, the leader broadcasts *precommitQC*, waits to form *commitQC*, and broadcasts it.

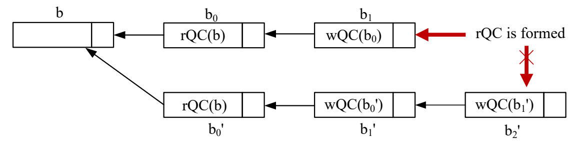


(a) Dashing1.

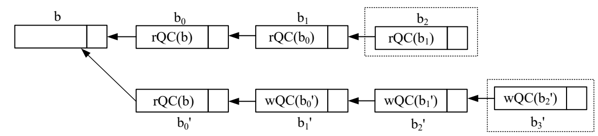


(b) Conventional BFT (e.g., HotStuff).

Figure 2. The way how Dashing1 and a regular BFT handle unexpected network delays, respectively.



(a) The safety challenge within a view.



(b) The safety challenge across views.

Figure 3. Challenges of building BFT from weak certificates.

Upon receiving the *precommitQC*, a replica becomes *locked* on b. Upon receiving the *commitQC*, a replica delivers b.

During view changes, each replica sends the leader its latest *prepareQC*. Upon receiving a quorum of $n - f$ such messages, the leader selects the QC of the largest height and extends the block for the QC using a new proposal.

Throughout the paper, we use the chained version for HotStuff and Dashing, where phases are overlapping and pipelined.

4.2 Overview of Dashing1

Dashing1 in a nutshell. In Dashing1, we use weak certificates (signatures from $f + 1$ replicas) to improve both efficiency and robustness. The main idea is to use weak certificates *as much as possible*, during normal cases, across views, and in the presence of transient network interruptions and connection discrepancies. In any of the above cases, we allow replicas to "proceed" with weak certificates.

As an example, consider a system with seven replicas, p_1 to p_7. The leader p_1 can only receive messages from p_2 and p_3, but not from other replicas. During the network interruption, replicas in existing partially synchronous BFTs cannot make meaningful progress. They have to wait until a regular certificate is formed, or until a view change occurs. In contrast, Dashing1 allows replicas to make meaningful progress and accumulate proposals under unexpected delays.

Fig. 2a describes the way how Dashing1 and a regular BFT protocol (e.g., HotStuff or PBFT) handle unexpected network delays. For both protocols, starting from time t_1, the leader could not form an rQC until t_2 when the network becomes synchronous again. During the network delay, the leader in conventional BFT protocol simply waits for its rQC. In contrast, in Dashing1 the leader can form a sequence of wQCs for blocks $b_1, \cdots, b_k$. Then when the network becomes synchronous, replicas can receive all the messages and catch up with the leader in a very short period of time. After generating three rQCs for block b_{k+1}, blocks $b_1, \cdots, b_{k+1}$ are committed simultaneously. Hence, we make full use of our computation and network bandwidth during expected network delays.

With our design, Dashing1 can naturally leverage network connection discrepancies for adaptive performance. The performance of Dashing1 depends on the group of fast replicas (1/3 of total replicas) rather than the Byzantine quorum of replicas (the overall network condition).

Challenges and our design. Transforming the idea into a fully secure BFT protocol, however, is non-trivial. First, a faulty leader may easily create forks and generate conflicting weak certificates. To prevent the forks from growing exponentially, we can ask each correct replica to vote for at most one block at each height.

Second, we need to ensure that the protocol achieves safety within a view even if wQCs are used. Namely, if forks are formed, an rQC can only be formed for at most one of the branches. As shown in Fig. 3a, b_1 and b_2' are conflicting blocks and an rQC is formed for b_1. Here we need to ensure that an rQC will never be formed for b_2'. We solve the problem by enforcing a constraint: if a replica receives a proposal for block b_2' that extends a block b_1' with a wQC, the replica votes for b_2' if and only if it has previously voted for the parent block b_1'. Then due to the quorum intersection, rQC for b_2' cannot be formed within the view.

Third, the protocol should achieve safety across views. During view changes, we ask each replica to send its *highest* wQC to the new leader and the new leader can select a branch led by a wQC to extend. However, the new leader may not choose the *right* wQC. As shown in Fig. 3b, rQCs are formed for b_0, b_1, and b_2, while wQCs for b_0', b_1', b_2', and b_3' are formed too (a "fork"). Note that an rQC for b_2 is also the *commitQC* for b_0. If a view change occurs and the leader selects the highest weak certificate (a wQC for b_3'), a *conflicting proposal* with the committed block b_0 will be proposed.

To address this issue, for any block b, we additionally define a *stable block* as the highest block with an rQC on the branch led by b. After a new leader collects the certificates from $2f + 1$ replicas, it will select a *safe block* to extend based on the highest rQC and the wQC with the highest stable block. In this example, as the stable block of b_3' is b and b is lower than b_2, the leader will create a proposal extending b_2. Correct replicas can check whether the new leader selects the right branch according to their locked blocks.

One more (liveness) challenge is about timers. In Dashing1, besides the regular view change timer Δ_1, the leader additionally maintains a timer Δ_2. After forming a wQC for block b with $f+1$ matching votes, the leader starts a timer Δ_2. When Δ_2 expires or an rQC for b is formed, the leader continues to propose a new block. Therefore, we need to be careful about Δ_2. Fortunately, an overly large Δ_2 does not cause any (performance) issues, as the leader will propose a new block once $n-f$ votes are received. Namely, even if we set an overly large Δ_2, Dashing1 would remain at least as efficient as HotStuff and is still optimistically responsive. Also, we comment that in settings with natural network discrepancies, we set Δ_2 according to concrete network connection conditions.

4.3 Notation for the Dashing Protocols

Blocks. A block b is of the form $\langle req, pl, sl, view, height\rangle$. We use $b.x$ to represent the element x in block b. Fixing a block b, $b.pl$ is the hash digest of b's parent block, $b.height$ is the number of blocks on the branch led by b, and $b.view$ is the view in which b is proposed. Note that different from the prior notation, sl is a new element in b. Formally, $b.sl$ denotes the hash digest of b's stable block (the highest block with a regular certificate on the branch led by b). For simplicity, we also use $b.parent$ and $b.stable$ to represent the parent block and the stable block of b, respectively.

Messages. Messages transmitted among replicas are of the form $\langle$TYPE, $block$, $justify\rangle$. We use three message types—GENERIC, VIEW-CHANGE, and NEW-VIEW. GENERIC messages are used in normal operations. VIEW-CHANGE and NEW-VIEW messages are used during view change: VIEW-CHANGE messages are sent by replicas to the next leader, while NEW-VIEW messages are sent by the new leader to the replicas. The $justify$ field stores certificates to validate the $block$. Fields may be set as $\perp$.

Functions and notation for QCs. A QC for message m is also called a QC for $m.block$. Fixing a QC qc for a block b, let QCBLOCK(qc) return the block b.

To hide implementation details of the QCs, we let QCVOTE(m) denote the output of a partial signing algorithm for m or a conventional signing algorithm and let QCCREATE(M) be a QC generated from signatures in M. QCCREATE(M) may be a wQC, an rQC, or an sQC.

Rank of QCs and blocks. Following the notion in [22], we now define the $rank()$ function for QCs and blocks. $rank()$ does not return a concrete number. Instead, it takes as input two blocks or QCs and outputs whether the rank of a block/QC is higher than the other one. The rank of two blocks/QCs is first compared by the view number, then by the height.

Local state at replicas. Each replica maintains the following state parameters, including the current view number $cview$, the highest rQC QC_r, the highest wQC QC_w, the locked block lb, and the last voted block vb.

Algorithm 1: Utilities

```
procedure CREATEBLOCK(b', v, req, qc)
    b.pl ← hash(b'), b.parent ← b', b.height ← b'.height+1
    b.req ← req, b.view = v
    ┆ if qc is a wQC then
    ┆     b.sl ← b'.sl, b.stable ← b'.stable, return b
    if qc is an rQC then b.sl ← b.pl, b.stable ← b' return b
procedure STATEUPDATE(QC_w, QC_r, lb, qc)
    b' ← QCBLOCK(qc), b'' ← b'.parent, b* ← b''.parent,
    v ← b'.view, b_0 ← QCBLOCK(QC_w), b_high ← QCBLOCK(QC_r)
    if qc is an rQC then
        if rank(b') > rank(b_high) then QC_r ← qc
        if b'.stable = b'' and rank(b'') > rank(lb) then lb ← b''
        if b'.stable = b'' and b''.stable = b* and
        b''.view = b*.view = v then
            deliver the transactions in b* and ancestors of b*
    ┆ if qc is a wQC and rank(b'.stable) ≥ rank(b_0.stable) then
    ┆     QC_w ← qc
```

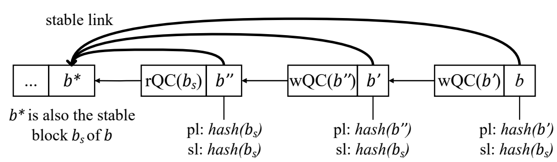


Figure 4. Illustration of the relationships of blocks in Algorithm 2.

4.4 Dashing1

We present in Algorithm 2 and Algorithm 3 the normal case protocol and view change protocol of Dashing1, respectively. The utility functions are presented in Algorithm 1. We largely follow the description of HotStuff and highlight how Dashing1 supports wQCs in dotted boxes.

Normal case protocol (Algorithm 2). In each phase, the leader broadcasts a message and waits for signed responses from replicas. At lines 9-10, the leader first proposes a new block b and broadcasts a $\langle$GENERIC, b, $qc_{high}\rangle$ message, where qc_{high} is the last QC it received (either a wQC or an rQC). The leader waits for the votes from the replicas. After collecting $f+1$ matching votes, the leader starts a timer Δ_2 (ln 6) to determine if the leader should stop waiting for more votes and propose a new block. Namely, the leader can propose a new block if either one of the two conditions is met: 1) Δ_2 expires; 2) an rQC for b is formed. Then the leader combines the signatures in the votes into qc_{high} for the next phase.

Upon receiving a $\langle$GENERIC, b, $\pi\rangle$ message from the leader, each replica p_i first verifies whether b is well-formed (ln 13-16), i.e., b has a higher rank than its parent block b' and $b.height = b'.height + 1$. Let b'' denote the parent of b', we distinguish two cases. For ease of understanding, we illustrate in Fig. 4 the relationships of b, b', b'', and b^*.

- If the π field is a wQC for b' (ln 17), p_i verifies if the stable block of b and b' are the same such that b indeed extends b'. p_i also verifies if b, b' and b'' are all proposed in the

Algorithm 2: Normal case protocol of Dashing1 for p_i

1 **initialization:** $cview \leftarrow 1$, vb, QC_w, QC_r, lb are initialized to $\perp$
2 Start a timer Δ_1 for the first request in the queue of pending transactions
3 ▷ **GENERIC phase:**
4 **as** a leader
5 **wait for** votes for b: $M \leftarrow \{\sigma | \sigma$ is a signature for $\langle \text{GENERIC}, b, \perp \rangle \}$
6 **upon** $|M| = f+1$ **then** set a start timer Δ_2
7 **upon** Δ_2 *timeout* or receiving $n-f$ matching messages **then**
8 $qc_{high} \leftarrow \text{QCCREATE}(M)$
9 $b \leftarrow \text{CREATEBLOCK}(b, cview, req, qc_{high})$
10 broadcast $m = \langle \text{GENERIC}, b, qc_{high} \rangle$
11 **as** a replica
12 **wait for** $m = \langle \text{GENERIC}, b, \pi \rangle$ from $\text{LEADER}(cview)$
13 $b' \leftarrow b.parent$, $b'' \leftarrow b'.parent$, $b_s \leftarrow b.stable$
14 $m \leftarrow \langle \text{GENERIC}, b, \perp \rangle$
15 **if** $rank(b') \geq rank(b)$ **or** $b.height \neq b'.height + 1$
16 **discard the message**
17 **if** π is a wQC for b' **and** $b_s = b'.stable$ **and** $b_s.view = b'.view = b''.view = cview$ **and** $b' = vb$ **then** $vb \leftarrow b$, $\text{STATEUPDATE}(QC_w, QC_r, lb, \pi)$
18 **if** π is an rQC for b' **and** $b_s = b'$ **and** $rank(b') \geq rank(vb)$
19 $vb \leftarrow b$, $\text{STATEUPDATE}(QC_w, QC_r, lb, \pi)$
20 **if** $vb = b$ **then** send $\text{QCVOTE}(m)$ to $\text{LEADER}(cview)$
21 ▷ **NEW-VIEW phase:** switch to this line if Δ_1 *timeout* occurs
22 **as** a replica
23 $cview \leftarrow cview + 1$
24 send $\langle \text{VIEW-CHANGE}, \perp, (QC_r, QC_w) \rangle$ to $\text{LEADER}(cview)$

Algorithm 3: View change protocol of Dashing1 for p_i

1 ▷ **VIEW-CHANGE phase**
2 **as** a new leader //M is a set of $n-f$ VIEW-CHANGE messages
3 $qc_{high} \leftarrow$ the rQC of highest rank contained in M
4 $b_1 \leftarrow \text{QCBLOCK}(qc_{high})$
5 **for** $m \in M$ **if** a wQC $qc_d \in m.justify$ **and** $\text{QCBLOCK}(qc_d) = d$ **and** $rank(d.stable) > rank(b_0.stable)$ **then** $vc \leftarrow qc_d$, $b_0 \leftarrow d$ **if** $rank(b_0.stable) \geq rank(b_1)$ **then** $b \leftarrow \text{CREATEBLOCK}(b_0, cview, req, vc)$, broadcast $m = \langle \text{GENERIC}, b, vc \rangle$
6 **else then**
7 $b \leftarrow \text{CREATEBLOCK}(b_1, cview, req, qc_{high})$
8 broadcast $m = \langle \text{GENERIC}, b, qc_{high} \rangle$
9 //switch to normal case protocol
10 **as** a replica
11 **wait for** $m = \langle \text{GENERIC}, b, \pi \rangle$ from $\text{LEADER}(cview)$
12 $b' \leftarrow b.parent$, $b_s \leftarrow b.stable$, $m \leftarrow \langle \text{GENERIC}, b, \perp \rangle$
13 **if** $b'.view \geq cview$ **or** $rank(b') \geq rank(b)$ **or** $b.height \neq b'.height + 1$ **then discard the message**
14 **if** π is a wQC for b' **and** $b_s = b'.stable$ **and** $rank(b_s) \geq rank(lb)$ **then** $vb \leftarrow b$, $\text{STATEUPDATE}(QC_w, QC_r, lb, \pi)$
15 **if** π is an rQC for b' **and** $b_s = b'$ **and** $rank(b_s) \geq rank(lb)$
16 **then** $vb \leftarrow b$, $\text{STATEUPDATE}(QC_w, QC_r, lb, \pi)$
17 **if** $vb = b$ **then** send $\text{QCVOTE}(m)$ to $\text{LEADER}(cview)$
18 //switch to normal case protocol. Three consecutive rQCs are required for the first block proposed during the view change.
19 ▷ **NEW-VIEW phase:** switch to NEW-VIEW phase if Δ_1 times out

same view and p_i has voted for b'. If so, p_i updates its local parameter QC_w to π (Algorithm 1, ln 15).

- If π is an rQC for b' (ln 18-19), p_i verifies if b's parent block b' has a higher rank than vb. If so, p_i updates its QC_r to π and generates a signature. If b'' has an rQC and b'' has a higher rank than the locked block of p_i, then p_i updates its lb to b''. If p_i has received an rQC for both b'' and b^* (the parent block of b''), then p_i commits block b^* and delivers the transactions in b^* (Algorithm 1, ln 6-14).

In both cases, the replica updates its vb to b and sends its vote (a signature for m) to the leader (ln 20).

View change protocol (Algorithm 3). Every replica starts timer Δ_1 for the first transaction in its queue. If the transaction is not processed before Δ_1 expires, the replica triggers view change. In particular, the replica sends a $\langle$VIEW-CHANGE, $\perp, (QC_r, QC_w)\rangle$ message to the new leader (Algorithm 2, ln 23-24). Upon receiving $n-f$ VIEW-CHANGE messages, the leader first obtains a block b_1 with an rQC that has the highest rank (ln 4). The leader then obtains a block b_0 with a wQC vc such that among all the blocks with weak QCs, b_0 has the highest stable block (first part of ln 5). Then the leader checks if the rank of the stable block of b_0 is no less than that of b_1 (second part of ln 5). If so, the leader creates a new block b extending b_0 and broadcasts b to all replicas. Otherwise, the leader extends b_1, and creates and broadcasts block b to the replicas (ln 7 and ln 8).

Upon receiving a $\langle$GENERIC, $b, \pi\rangle$ message from a new leader, each replica p_i verifies if the proposed block b extends a block of a prior view (ln 13). Then p_i votes for b if either of the following conditions is satisfied: 1) b extends a block b' with a wQC (ln 14), the stable blocks of b and b' are the same block (denoted as b_s), and the rank of b_s is no less than that of the locked block of p_i; 2) b extends a block b' with an rQC (ln 15-16), and the rank of the stable block of b is no less than that of the locked block of p_i.

For the first block proposed in a new view, the leader needs to collect three consecutive rQCs after replicas switch to the normal case protocol (ln 17). As discussed in Sec. 4.2, this rule is crucial for dealing with the liveness challenge caused by the timer Δ_2. Moreover, one may optionally enforce an additional rule such that the leader should commit at least one block after proposing a "sufficient" number of blocks with wQCs (say, 50 blocks).

State transfer. As in HotStuff, replicas in Dashing1 may need to perform state transfer with other replicas to obtain the QCs or transactions included in the QCs. For the state transfer of QCs, if a replica learns that a block b with height

h is committed but it has not received any QCs between height h' of its latest committed block and h, the replica has to synchronize all the QCs for blocks between h' and h on the branch led by b. For the state transfer of the transactions for each QC, the replica needs to obtain the proposal from other replicas such that the hash of the proposal matches that in the QC.

Correctness. Below we briefly argue how Dashing1 addresses the challenges mentioned in Sec. 4.2. For safety within a view, as every correct replica votes for a block b with wQC only if it has voted for all the blocks on the branch led by b, correct replicas will never commit blocks on two conflicting branches. For safety across views, the stable block introduced ensures that a new leader always proposes blocks extending the *right* branch, where the blocks are non-conflicting with any committed block. For liveness, the crux is to show that after GST, each block proposed by a correct leader can be accepted by all correct replicas. Indeed, the leader extends either the highest rQC received from other replicas or a wQC extending the highest rQC; eventually all correct replicas will vote for the block. We provide the full proof in Appendix A.

4.5 Dashing2

We show in Dashing2 how to further enable a fast path using sQCs. Intuitively, supporting a $3f+1$ threshold may allow replicas to deliver the transactions in a single phase: if the leader collects an sQC for a block and broadcasts to the replicas, replicas can directly commit the block.

While prior works have demonstrated how to design secure BFT protocols using strong quorums [3, 4, 25], integrating sQCs in Dashing1, however, has its unique challenges due to the usage of wQCs. Indeed, as a block supported by an sQC may be extended from a block with only a weak certificate, replicas cannot directly commit the block upon receiving an sQC. As depicted in Fig. 5, two conflicting blocks b and b' are proposed in the same view (view 1) with the same height. Moreover, an rQC is formed for b and a wQC is formed for b'. Besides, a wQC for block b_1' that extends b_0' is formed. Suppose now a view change occurs, the new leader in view 2 extends b_1' and proposes b_2'. Replicas can vote for b_2', so an sQC can be formed and at least one correct replica commits b_2'. Then we consider a scenario where another view change occurs and replicas enter view 3. As there is no guarantee on how many correct replicas have received the sQC for b_2', the new leader in view 3 may choose to extend b_0. And b_0 can be later committed in view 3, in which case safety is violated as b_2' is committed in view 2. As a view change may occur at any moment, replicas cannot directly commit a block when an sQC is received.

We thus make several major changes on top of Dashing1 to address the challenge. First, in normal cases, if replica p_i receives an sQC for a block b that extends a block with an rQC/sQC, p_i immediately commits b and our protocol

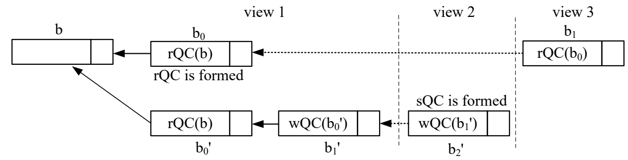


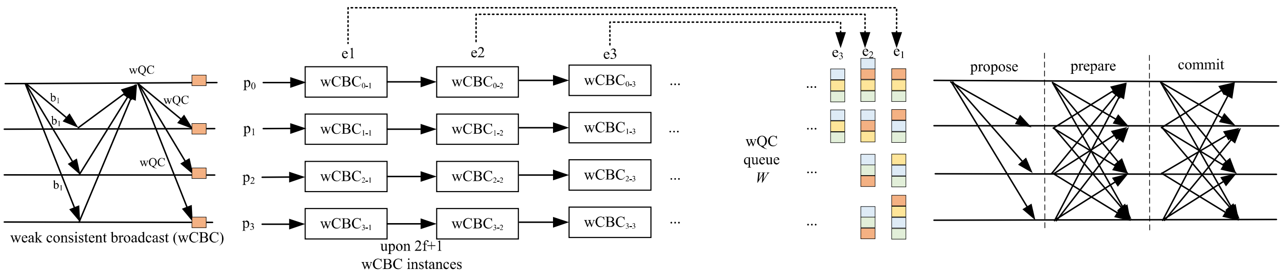
Figure 5. Challenge of integrating strong certificates in Dashing2.

admits a fast path in this way. However, if block b extends a block with a wQC, we prevent p_i from committing b in the fast path. In this case, p_i should wait for two consecutive rQCs for b before committing b. Second, Dashing2 follows the two-phase commit rule that if a replica receives an rQC for both a block b and b' (the parent block of b), block b' can be committed. Third, we modify the view change protocol. For the first block b proposed after each view change, the leader forms an rQC rather than wQC or sQC to start the normal case operations. Also, during view change, the NEW-VIEW message from the new leader includes a set of at least $n-f$ VIEW-CHANGE messages. Upon receiving the NEW-VIEW message with a proposal, a correct replica verifies the proposal by performing a computation as the one used by the new leader to create the proposal. Replicas resume normal operations only after the NEW-VIEW message is verified. Indeed, the view change protocol now becomes similar to that in Fast-HotStuff [26] and Jolteon [22]. Hence, Dashing2 has $O(n^2)$ authenticators and $O(n)$ messages.

Note that like BFT protocols using strong quorums [3, 4, 25], Dashing2 does not achieve optimistic responsiveness (which is unavoidable due to the one-phase fast path). We show the pseudocode of Dashing2 in Appendix B and proof of correctness in Appendix B.1.

5 The Star Framework

We present Star that allows replicas to concurrently propose transactions and deliver at least $n-f$ proposals in each epoch. As in Narwhal and Tusk [15], the transmission and consensus processes in Star (as described in Fig. 6) are decoupled. However, Star uses several new techniques for improved performance over DAG-based protocols. First, we use the more efficient wQCs for the data transmission layer. The transmission process is fully parallelizable and works in asynchronous environments. It proceeds in epochs, where all replicas can propose transactions and output a queue of weak certificates numbered by epochs. The consensus process has only one BFT instance and does not carry bulk data. It takes as input weak certificates of the proposals and agrees on which proposals in each epoch should be delivered. Second, the transmission layer in Star can be effectively pipelined and provides more efficient communication and state transfer. Moreover, the transmission process and the consensus process are implicitly "correlated" with epoch numbers, and the consensus process only handles messages transmitted in the same epoch, which helps achieve effective censorship resilience and improve blockchain quality. Such a design leads to reduced complexity and improved performance overall.



(a) Transmission process

(b) Consensus process

Figure 6. The Star architecture.

The transmission process. The transmission process evolves in epochs. Each epoch consists of n parallel weak consistent broadcast (wCBC) instances, as shown in Fig. 6 (a). Each replica maintains a queue Q of pending transactions and outputs a growing set $W[e]$ containing weak certificates for each epoch e. In each wCBC instance, a designated replica broadcasts a proposal (a batch of transactions) from its queue of pending transactions. Upon completing $n-f$ wCBC instances, each replica starts the next epoch and continues to propose new transactions.

wCBC may be viewed as a weak version of consistent broadcast (CBC), i.e., CBC with weak certificates. A wCBC instance consists of three steps. First, a designated sender sends a proposal containing a set of transactions to all replicas. The sender waits for signed responses from $f+1$ replicas to form a wQC and sends it to all replicas. Upon receiving a valid wQC, each replica delivers the corresponding proposal. Note it is possible that for a particular wCBC instance, a correct replica delivers m and another correct replica delivers $m' \neq m$. While multiple conflicting wQCs might be provided by a faulty sender, at most one wQC will be delivered.

So why wCBC? wCBC ensures that if a wQC is formed, at least one correct replica has received and stored the corresponding proposal. The use of wQCs is *sufficient* to ensure liveness, because any replica p_j, once obtaining wQC, can ask for the corresponding proposal from correct replicas; any correct replica that stores the proposal can simply send it to p_j, which can then validate the correctness of the proposal via the wQC. The above procedure is needed only when a correct replica stored a wQC but had no corresponding proposal. Even if the scenario occurs, it would not incur higher message or communication complexity.

Star develops the above idea and offers a pipelined version for high performance. Concretely, each replica can directly put forward a new proposal in the third step of wCBC. We describe the code of the transmission process at ln 3-13 of Algorithm 4, where each replica p_i ($i \in [0..n-1]$) runs the *initepoch*(e) function to start a new epoch e. Replica p_i chooses a set of transactions from Q as a proposal (say, b) using the *select* function. (The *select* function is vital to liveness and we will discuss its specification shortly.) It then broadcasts a message $\langle \textsc{proposal}, e, b, wqc\rangle$, where wqc is the wQC formed in epoch $e-1$. (If we are working in the non-chaining mode, then wqc is simply $\perp$.) p_i waits for $f+1$ votes for b to form a wQC. Then after receiving $n-f$ proposals for epoch e, p_i enters the next epoch $e+1$. Upon receiving $\langle \textsc{proposal}, e, b_j, wqc_j\rangle$ from p_j, each replica first verifies wqc_j, sends a signed vote for b_j to p_j, adds b_j to *proposals*, and adds wqc_j to $W[e-1]$.

Algorithm 4: The code of Star for p_i

1 **initialization:** epoch number e and the epoch number of the current block le are initialized to 1. Queue Q of pending transactions, received proposals *proposals*, the latest weak certificate wqc, and queue W of weak certificates are initialized to $\perp$
2 ▷ **transmission process** in the chaining (pipelined) mode
3 **func** *initepoch*(e)
4 $b.tx \leftarrow select(Q)$, $b.epoch \leftarrow e$ //select a proposal b from Q
5 broadcast $\langle \textsc{proposal}, e, b, wqc\rangle$
6 **upon** receiving a set M of $f+1$ signed votes for b
7 $wqc \leftarrow \textsc{qcCreate}(M)$ //create a weak certificate
8 **wait until** $|proposals[e]| \geq n-f$ //enter the next epoch
9 $e \leftarrow e+1$, *initepoch*(e)
10 **upon** receiving $\langle \textsc{proposal}, e, b_j, wqc_j\rangle$ from p_j for the first time
11 send signed vote for b_j to p_j
12 $proposals[e] \leftarrow proposals[e] \cup b_j$
13 $W[e-1] \leftarrow W[e-1] \cup wqc_j$ //certificates in the output queue
14 ▷ **consensus process**
15 **upon** $|W[le]| \geq n-f$
16 *a-broadcast*$(W[le])$ //run the underlying atomic broadcast
17 **upon** *a-deliver*(le, m)
18 $O \leftarrow obtain(le, m)$
19 *deliver* O //deliver the transactions in O in deterministic order
20 $le \leftarrow le+1$
21 ▷ **state transfer**
22 **func** *obtain*(e, m)
23 $O \leftarrow \perp$
24 **for** wQC $qc \in m$
25 **if** $\textsc{qcProposal}(qc) \in proposals[e]$
26 $O \leftarrow O \cup \textsc{qcProposal}(qc)$
27 **else** broadcast $\langle \textsc{fetch}, e, qc\rangle$
28 **wait for** a $\textsc{proposal}$ containing $\textsc{qcProposal}(qc)$
29 $O \leftarrow O \cup \textsc{qcProposal}(qc)$
30 clear $W[e]$, remove transactions in O from Q
31 **upon** receiving message $\langle \textsc{fetch}, e, qc\rangle$ from replica p_j
32 **if** $\textsc{qcProposal}(qc) \in proposals[e]$ //fetch missing proposals
33 send $\langle \textsc{proposal}, \textsc{qcProposal}(qc)\rangle$ to p_j

The consensus process. The consensus process also proceeds in epochs, using only one BFT instance to agree on wQCs. We can use any BFT protocol for the consensus process. When describing the consensus process (Algorithm 4, ln 15-20), we use the *a-broadcast* and *a-deliver* primitives.

Each replica p_i maintains le, a local parameter tracking the current consensus epoch number. p_i monitors its queue W (obtained from the transmission process) and checks whether $W[le]$ has at least $n-f$ weak certificates. If so, replicas run *a-broadcast(*$W[le]$*)*. (If the underlying BFT is leader-based, then only the leader proposes $W[le]$). When the *a-deliver* primitive terminates, each replica waits for the transactions corresponding to the a-delivered wQCs (from the transmission process) and delivers the transactions in a deterministic order. If some proposals are missing, the replica may simply fetch the proposals from other replicas (via the state transfer process at ln 21-33 of Algorithm 4). During state transfer, for each wQC qc in epoch e, a replica p_i broadcasts a $\langle \text{FETCH}, e, qc\rangle$ message to all replicas. Upon receiving such a message, a replica sends the corresponding proposal to p_i.

Liveness and blockchain quality. Protocols allowing all replicas to propose different transactions should address transaction censorship (liveness) which prevents a particular transaction proposed by a replica from never being delivered. First, the use of wQC ensures that if the underlying atomic broadcast completes, then the corresponding proposal has been obtained by correct replicas, or can be obtained via the fetch operation by correct replicas. We should also ensure that adversaries cannot censor certain transactions. So we have to be careful in specifying the *select* function. HoneyBadgerBFT [32] invents a method where replicas randomly select transactions from their queue and use threshold encryption to achieve censorship resilience. EPIC [31] combines the conventional FIFO strategy used in [9] and the random selection strategy used in HoneyBadgerBFT to avoid threshold encryption. The asynchronous pattern in Star allows us to adopt the same approach as in EPIC: replicas select random transactions for most epochs and periodically switch to FIFO. Hence, Star achieves liveness under asynchrony.

Star has a strong form of *blockchain quality*, ensuring at least 1/2 of transactions contained in *any* committed block in an epoch are from correct replicas. Note the concurrent work of Dumbo-NG [21] sacrifices this desirable feature.

Instantiating Star using PBFT. In Star, we use a variant of PBFT with the following small differences. First, as the proposed transactions are already assigned with epoch number in the transmission process, we directly use the epoch numbers as the *sequence* number in the consensus process. We additionally require that the leader cannot skip any epoch number. Last, during a view change, the new leader is not allowed to propose a nil block for any epoch number. Namely, for any epoch e such that an agreement is not reached in a prior view, the new leader simply proposes $W[e]$.

Complexity analysis. Star has n parallel wCBC instances and one instance of the underlying BFT protocol, so Star has $O(n^2)$ messages (whether using PBFT or HotStuff). The communication is $O(Ln^2+\lambda n^2)$ for the transmission process and $O(\lambda n^2)$ for the consensus process. As a replica can directly obtain a proposal based on epoch number and each QC, the time for state transfer of multiple QCs is $O(1)$.

Instead, Narwhal has a complex state transfer process. In particular, replicas have to obtain sequentially the blocks for each epoch since there is no guarantee that at least one correct replica holds the entire history. Hence, if a replica performs state transfer for a transaction proposed k epochs ago, the time is $O(k)$. Moreover, Narwhal has $O(Ln^2+\lambda n^3)$ communication, as each block includes at least $2f+1$ certificates of the prior epoch.

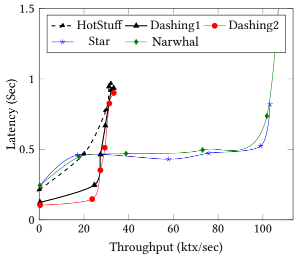
Discussion. An attempting approach relevant to Star is to allow clients to distribute the full transactions to all replicas and perform the agreement only on hashes. First, directly broadcasting transactions may not be safe or live. For instance, if Byzantine clients fail to send transactions consistently, then not all correct replicas can receive transactions. Intuitively, the transmission phase would use reliable broadcast, but Sonic allows the use of more efficient weak consistent broadcast (wCBC) and wCBC can be pipelined for better performance. Moreover, the attempting approach only allows replicas to deliver transactions proposed from the leader, while Star allows delivering transactions from $n-f$ different replicas at the same time. Our approach thus fully utilizes the network bandwidth.

6 Implementation and Evaluation

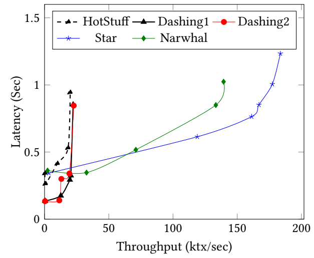
We implement all our protocols introduced in this work and HotStuff in Golang using around 12,000 LOC, including 1,500 LOC for evaluation. We implement the chaining (pipelining) mode for the Dashing protocols and HotStuff. For all the protocols, we implement the checkpoint protocol for garbage collection, where replicas run the checkpoint protocol every 5000 blocks. Following prior works [2, 23, 33, 39], we use a set of digital signatures as quorum certificates. In particular, we use SM2 signature (ISO standard) which has a similar performance as ECDSA. We also evaluate the performance of Narwhal using its open-source code [1].

We deploy the protocols in Amazon EC2 with up to 100 instances in both LAN and WAN. We use *m5.xlarge* instance which has four virtual CPUs and 16 GB memory. In the LAN setting, all the instances are located in the same region. In the WAN setting, the servers are evenly distributed over four different regions: us-west-1 (California, US), us-east-2 (Ohio, US), ap-southeast-1 (Singapore), and eu-west-1 (Ireland).

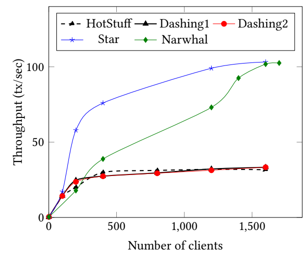
For each experiment, we use $3f+1$ replicas and use f to denote the network size. We ask the clients to submit requests to the system in an open loop, i.e., a client does not have to wait for the reply before sending the next request. We set the size for transactions and replies as 512 bytes. We



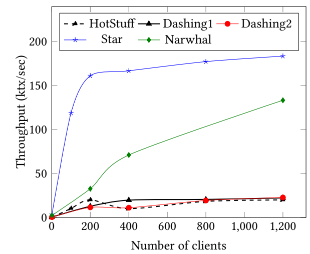
(a) Latency vs. throughput in LAN for $f = 1$.



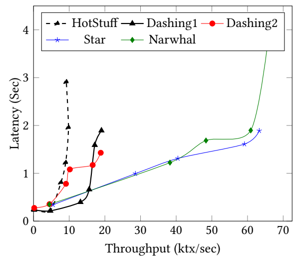
(b) Latency vs. throughput in LAN for $f = 10$.



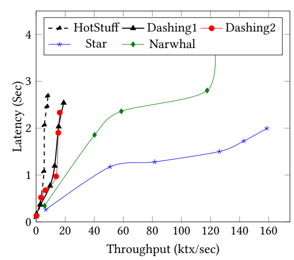
(c) Throughput in LAN for $f = 1$.



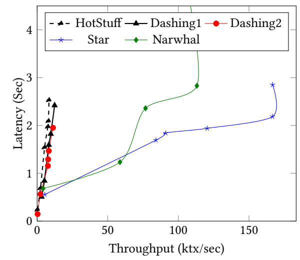
(d) Throughput in LAN for $f = 10$.



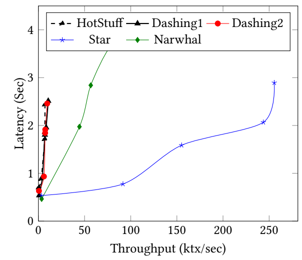
(e) Latency vs. throughput in WAN for $f = 1$.



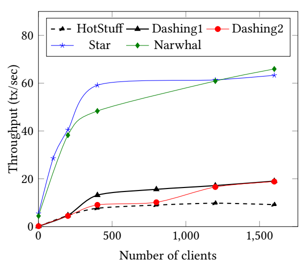
(f) Latency vs. throughput in WAN for $f = 10$.



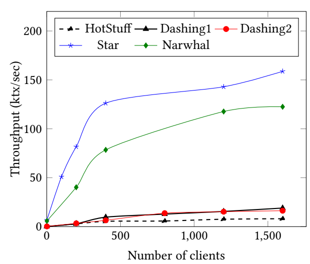
(g) Latency vs. throughput in WAN for $f = 20$.



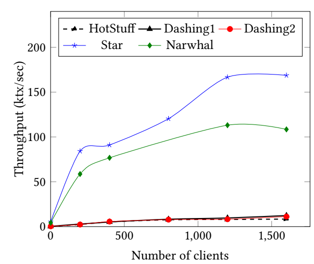
(h) Latency vs. throughput in WAN for $f = 30$.



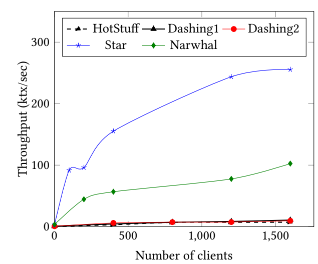
(i) Throughput in WAN for $f = 1$.



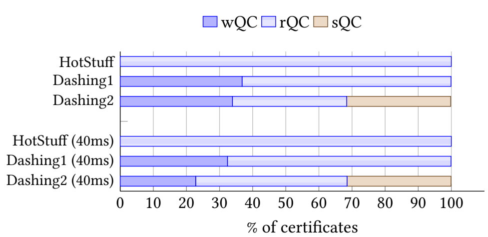
(j) Throughput in WAN for $f = 10$.



(k) Throughput in WAN for $f = 20$.



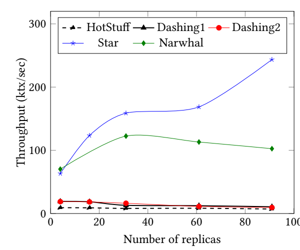
(l) Throughput in WAN for $f = 30$.



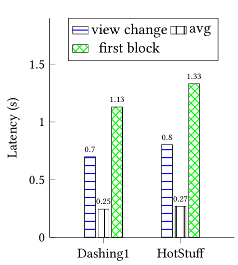
(m) Fractions of different certificates.

	ktx/sec	CPU	bandwidth
Narwhal	19.4	130%	8.8MB/s
	76.19	260%	33MB/s
	104.5 (peak)	330%	40MB/s
Star	12.2	150%	8MB/s
	34.0	190%	18MB/s
	153.8 (peak)	380%	100MB/s

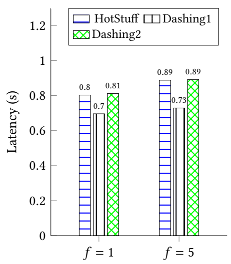
(n) CPU and bandwidth usage of Star and Narwhal. Maximum CPU usage is 400%.



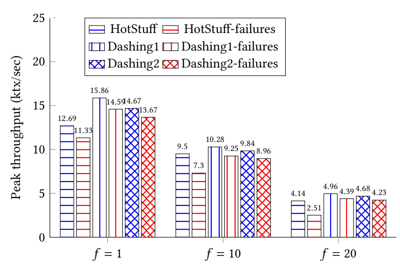
(o) Throughput of the protocols in WAN as f grows.



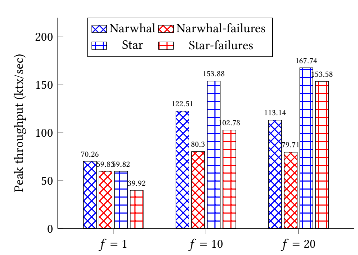
(p) Finality latency during view changes.



(q) View change latency under failures.



(r) Peak throughput of Dashing1, Dashing2, and HotStuff under f failures.



(s) Peak throughput of Star and Narwhal under f failures.

Figure 7. Performance of the protocols.

set Δ_2 as the average duration for one replica (e.g., the leader) from the point of time obtaining $f+1$ votes to the point of time obtaining $2f+1$ votes. (The range spans from 20ms to 800ms as f increases in our experiments.) We evaluate the throughput and latency of the protocols, where throughput is the number of transactions that can be delivered in a second and latency is the consensus time for each proposed block to be committed. We repeat each experiment five times and report the average results.

Performance (latency vs. throughput; throughput). We report the performance of protocols in both LAN and WAN settings. In LANs, we report latency vs. throughput for $f = 1$ and $f = 10$ in Fig. 7a and Fig. 7b and throughput as the number of clients increases in Fig. 7c and Fig. 7d. In the WAN setting, we report the performance of the protocols in Fig. 7e-7l. In all our experiments, Dashing1 and Dashing2 consistently outperform HotStuff. For instance, in WANs, the peak throughput of Dashing1 is 107.36% higher and 49.8% higher than that of HotStuff for $f = 1$ and $f = 30$, respectively. Indeed, the leader in Dashing only needs to collect wQCs rather than rQCs to proceed to the next phase, and thus Dashing protocols are more communication-efficient. To help understand Dashing vs. HotStuff, we additionally evaluate the fraction of QCs in Dashing1 and Dashing2. Our experiments are conducted in LANs for $f = 1$ in two settings: a setting with no network delay; and a setting with a 40ms network delay injected using the *tc* command. As shown in Fig. 7m, for Dashing1, the fraction of wQCs is 36.8% for the experiment with no delay and 32.4% for the one with 40ms delay. Similar results apply to Dashing2. The finding explains why Dashing improves HotStuff.

Star significantly and consistently outperforms other protocols. Meanwhile, when $f = 30$, Star achieves 2.38x the throughput of Narwhal. This improvement is due to the lower communication and a (much) simpler data structure used in Star, as well as the use of wQCs and pipelining.

To further understand the performance bottlenecks of the protocols, we also assess the CPU and bandwidth usage of the protocols using the *htop* and *nethogs* commands, respectively. We summarize the results for Narwhal vs. Star in Fig. 7n for $f = 10$ in WAN. Our results show that the bottleneck of both protocols is CPU (the maximum usage is 400% as each instance has 4 vCPU). When the CPU is fully utilized, Star in general consumes higher network bandwidth and processes more transactions than Narwhal, which explains why Star outperforms Narwhal. We observe a similar result for Dashing protocols vs. HotStuff.

Scalability. We report in Fig. 7o the peak throughput of Dashing1, Dashing2, Star, and HotStuff in WAN as f grows. All the Dashing protocols outperform HotStuff consistently. The peak throughput of Dashing1 is 47%-107% higher than that of HotStuff. For the Dashing protocols and HotStuff, the throughput degrades as f grows, echoing other protocols in the HotStuff family. The throughput of Narwhal first increases as f grows and then decreases as f grows further, matching the evaluation result reported in Narwhal [15].

bandwidth	HotStuff	Dashing1	improvement
10 Mbps	575	942	63.8%
20 Mbps	1,134	1,616	42.5%
50 Mbps	1,668	2,483	48.8%
200 Mbps	1,896	3,034	60.02%

Table 2. Peak throughput of HotStuff and Dashing1 in operation.

In comparison, the peak throughput of Star keeps growing as f increases (to 30). Meanwhile, the peak throughput of Star consistently outperforms other protocols. When $f = 30$, the peak throughput of Star is 243 ktx/sec, in contrast to 7 ktx/sec for HotStuff, 10 ktx/sec for Dashing1, and 102 ktx/sec for Narwhal. The performance difference is due to the separation from (pipelined) transmission from agreement as well as the parallel processing of transactions. For Sonic and Narwhal, the total peak throughput (roughly) equals the batch size multiplied by the network size $(n - f)$. While, for instance, the throughput of Sonic for $f = 10$ and that of $f = 20$ are rather similar, they reached their peak throughput under difference batch sizes.

Performance under failures. We assess the performance under failures for Dashing1, Dashing2, and HotStuff. We use 1,200 clients in all these experiments.

We first assess the average latency of view changes due to the leader failures caused by halting the leader in the middle of each experiment. We report the view change latency for $f = 1$ and $f = 5$ in Fig. 7q. We find the view change latency for Dashing2 is higher than Dashing1 and HotStuff, because each NEW-VIEW message consists of $n - f$ messages and replicas need to verify them accordingly.

We also report the peak throughput of the protocols for $f = 1$, 10, and 20, where we crash f replicas in each experiment. The throughput of Dashing1, Dashing2, and HotStuff degrades slightly under failures as shown in Fig. 7r. The throughput of HotStuff under failures is 10.71%-39.37% lower than that in the failure-free case. Meanwhile, the throughput degradation is 8.00%-11.39% and 6.80%-9.61% for Dashing1 and Dashing2, respectively. The lower performance degradation of Dashing protocols is again due to the use of wQCs.

We report the performance of Star and Narwhal under failures in Fig. 7s. Except for $f = 1$, the performance degradation of Star in the failure case is lower than that of Narwhal. For instance, when $f = 10$, the peak throughput of Star during failures is 33.21% lower than in the failure-free case, while the throughput of Narwhal during failures is 34.45% lower. When $f = 20$, the throughput degradation for Star and Narwhal is 8.44% and 29.5%, respectively.

Dashing1 in operation. Dashing1 has been deployed in a major cross-border payment system with nearly 20 commercial banks involved. The system uses dedicated bank networking channels (called Direct Connect) for communication. The average bandwidth between the sites is 25.7 Mbps.

(In contrast, the bandwidth in our Amazon EC2 experiments is significantly higher—around 10 Gbps.) Here we report the peak throughput with four sites (replicas) for Dashing1 and HotStuff with the following bandwidth settings: 10 Mbps, 20 Mbps, 50 Mbps, and 200 Mbps. Moreover, the machines used have 16-core CPU and 64 GB memory, and the transaction size is 218 bytes. As shown in Table 2, while both Dashing1 and HotStuff achieve lower performance when compared to those conducted on EC2, Dashing1 consistently outperforms HotStuff, showing weak certificates indeed lead to better performance.

In our production system, some application-level transaction validation may have high overhead due to the complex business logic and the protocol may thus experience unexpected view changes (even with a correct leader). Thus, we have to adjust the transaction processing programs to smooth the execution time and carefully tune view change timers. Also, for the 4-replica deployment, the system indeed can have more than 1 failure on rare occasions, calling for deployment on a larger scale.

7 Conclusion

We design and implement efficient BFT protocols using weak certificates, including Dashing offering improved efficiency and robustness compared to HotStuff, and a new BFT framework Star allowing processing parallel transactions using a single BFT instance. Via a deployment in both the LAN and WAN environments, we show that our protocols outperform existing ones of the same kind.

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A Correctness of Dashing1

We first introduce some notation we use in this section. Let b', b denote two blocks such that $b.parent = b'$. According to Algorithm 2 and Algorithm 3, after receiving a GENERIC message $\langle$GENERIC$, b, qc\rangle$, a correct replica votes for b only if (1) $b.stable = b'$ and qc is an rQC for b' (ln 18-19 of Algorithm 2 and ln 17-18 of Algorithm 3); or (2) $b.stable = b'.stable$ and qc is a wQC for b' (ln 17 of Algorithm 2 and ln 16 of Algorithm 3). In both cases, we say that qc and b are *matching*.

Let b, b' and b'' denote three consecutive blocks. In Algorithm 1, we have that a replica p_i commits b only after receiving an rQC qc for b'' such that $b''.stable = b'$, $b'.stable = b$, and $b.view = b'.view = b''.view = v$. In this case, we call qc a *commitQC* for b.

Lemma A.1. *If b and d are two conflicting blocks and $rank(b) = rank(d)$, then an rQC cannot be formed for both b and d.*

Proof. Let v denote $b.view$. As $rank(b) = rank(d)$, we have $d.view = v$. Suppose, towards a contradiction, an rQC is formed for both b and d. As a valid rQC consists of $2f + 1$ votes, a correct replica has voted for both b and d in view v. This causes a contradiction, because in the same view and for any height, a correct replica votes for at most one block. □

Lemma A.2. *Suppose that there exists an rQC or a wQC qc for b; if block d and d_c are on the branch led by b such that $d_c.parent = d$, then we have that*

(1) $d.height < d_c.height$ and at least one correct replica has received a certificate qc_d for d, where qc_d and d_c are matching;

(2) and if the view of the parent block of d is lower than $d.view$, then at least one correct replica has received an rQC qc_d for d and $d_c.stable = d$.

Proof. (1) We prove the claim (1) by induction for d. If $d = b.parent$, then d_c equals b. Since qc is an rQC or a wQC for b, at least one correct replica has voted for d_c. Then we have that $d.height < d_c.height$ and p_i has received a qc_d before voting for d_c, where qc_d and d_c are matching.

If $d \neq b.parent$, then there exists an rQC or a wQC for any block higher than d on the branch led by b. In this situation, there exists a block d_c on the branch led by b such that $d_c.parent = d$; an rQC or a wQC qc_c for d_c is received by at least one correct replica. Since qc_c consists of at least $f + 1$ votes, at least one correct replica p_i has voted for d_c in view $d_c.view$. Then we have that $d.height < d_c.height$ and p_i has received a qc_d before voting for d_c, where qc_d and d_c are matching. This completes the proof of claim (1).

(2) Based on claim (1), we know that at least one correct replica p_i has voted for d_c in view $d_c.view$. Let d' denote the parent block of b. Then $d'.view < d.view$. According to ln 16-18 of Algorithm 2, p_i votes for d_c only if p_i has received a rQC qc_d for d and $d_c.stable = d$. □

Lemma A.3. *If there exists a wQC qc_d for block d, then d extends $d.stable$ and at least one correct replica has received a rQC for $d.stable$.*

Proof. Let d_0 denote $d.parent$. As there exists a wQC for d, at least one correct replica p_i has received a certificate qc and voted for d in view $d.view$, where qc and d are matching. We distinguish two cases:

(1) qc is an rQC for d_0 and $d.stable = d_0$. Then we know that d extends $d.stable$, because d_0 is the parent block of d. Accordingly, at least one correct replica p_i has received a rQC qc for $d.stable$ before voting for d.

(2) qc is a wQC for d_0 and $d.stable = d_0.stable$. Let d_v denote the block of highest height on the branch led by d such that $d_v.stable \neq d.stable$. Let d'_v denote the block on the branch such that $d'_v.parent = d_v$. We have $d'_v.height > d_v$ and $d'_v.stable = d.stable$. Therefore, it follows from Lemma A.2 that at least one correct replica p_i has voted for d'_v. Thus, we have $d'_v.stable = d_v.stable$ or $d'_v.stable = d_v$ according to Algorithm 2 (ln 17-19). Since $d_v.stable \neq d.stable$, we have that $d'_v.stable \neq d_v.stable$. Then we know that $d'_v.stable = d.stabe = d_v$ and d extends $d.stable$. Meanwhile, p_i has received an rQC for d_v before voting for d'_v.

In both cases, d extends $d.stable$ and a correct replica has received an rQC for $d.stable$. □

Lemma A.4. *If there exists at least one rQC formed in view v, then there exists only one rQC qc of lowest rank in view v, and we have that*

(1) The view of $b.parent$ is lower than v, where b equals QCBLOCK(qc);

(2) If there exists an rQC for b_1 and $b_1.parent.view < v$, then b_1 equals b.

Proof. If an rQC is formed in view v, then there exists only one rQC qc of lowest rank in view v (according to Lemma A.1).

(1) Let b denote QCBLOCK(qc) and b_v denote the block of lowest height such that $b_v.view = v$ on the branch led by b. Therefore, $b_v.height \leq b.height$ and the view of $b_v.parent$ is

lower than v. According to Lemma A.2, there must exist a rQC for b_v. Since qc is the lowest rQC formed in view v, we have that $b_v = b$ and the view of $b.parent$ is lower than v.

(2) If there exists an rQC for b_1, then at least one correct replica has voted for b_1 and b in view v. Note that in view v, a correct replica only votes for one block that extends a block proposed in a lower view according to Algorithm 3. Therefore, it must hold that $b_1 = b$. □

Lemma A.5. *If rQC qc for b is the rQC of lowest height formed in view v and there exists an rQC for block d such that $d.view = v$, then d equals b or d is an extension of b.*

Proof. Let d_0 denote the block of lowest height on the branch led by d such that $d_0.view = v$. Then the view of the parent block of d_0 is lower than v. According to Lemma A.2, at least one correct replica has received an rQC for d_0. By Lemma A.4, it holds that d_0 equals b. As d_0 is a block on the branch led by d, d equals b or d is an extension of b. □

Lemma A.6. *Suppose qc_1 and qc_2 are two rQCs, and each is received by at least one correct replica. Let b_1 and b_2 be $\textsc{qcBlock}(qc_1)$ and $\textsc{qcBlock}(qc_2)$, respectively. If b_1 is conflicting with b_2, then $b_1.view \neq b_2.view$.*

Proof. Assume towards contradiction that $b_1.view = b_2.view = v$. According to Lemma A.5, we know that there exists a block b which is the block of lowest height for which an rQC was formed in view v, b_1 and b_2 are blocks and either b_1 or b_2 is equals b or is an extensions of b. Then $b_1.height \geq b.height$ and $b_2.height \geq b.height$. We consider three cases:

(1) If $b_1.height = b.height$ or $b_2.height = b.height$, then b_1 equals b or b_2 equals b. Therefore, b_1 and b_2 are the same block or they are on the same branch.

(2) If $b.height < b_1.height$, $b.height < b_2.height$, and $b_1.height = b_2.height$, then according to Lemma A.1, b_1 and b_2 must be the same block.

(3) If $b.height < b_1.height$, $b.height < b_2.height$, and $b_1.height \neq b_2.height$, then b_1 and b_2 are extensions of b. W.l.o.g., we assume that $b_1.height < b_2.height$. Let b_2' denote a block on the branch led by b_2 such that $b_2'.height = b_1.height$. Then b_2' is an extension of b. If b_2' is conflicting with b_1, then according to Lemma A.1, we have that no rQC for b_2' can be formed in view v and at most f correct replicas voted for b_2'. Thus, an rQC for any extensions of b_2' cannot be formed by Algorithm 2. Therefore, we have that b_2' must be equal to b_1.

In all cases, b_1 and b_2 must be blocks on the same branch, contradicting the condition that they are conflicting blocks. Therefore, we have that $b_1.view \neq b_2.view$. □

Lemma A.7. *If there exists a commitQC qc for b and an rQC qc_d for d, each is received by at least correct replica, and $rank(b) < rank(d)$, then d must be an extension of b.*

Proof. Let v be $b.view$, v_d be $d.view$, b'' be $\textsc{qcBlock}(qc)$, and b' be $b''.parent$. As qc is a *commitQC* for b, we have that $b'.stable = b'.parent = b$, $b''.stable = b'$, and $b.view = b'.view = b''.view = v$. According to Lemma A.2, there exist rQCs for b, b', and b'' such that all these rQCs are received by at least one correct replica. Note that an rQC for b' is also a *lockedQC* for b. Let S denote the set of correct replicas that have voted for b''. Since qc consists of $2f+1$ votes, we know that $|S| \geq f+1$.

Since $rank(d) > rank(b)$, $v_d \geq v$. Then we prove the lemma by induction over the view v_d, starting from view v.
Base case: Suppose $v_d = v$. According to Lemma A.6, d must be an extension of b.
Inductive case: Assume this property holds for view v_d from v to $v+k-1$ for some $k \geq 1$. We now prove that it holds for $v_d = v+k$. Let b_0 denote the block of lowest height for which an rQC qc_0 was formed in view v_d and b_0' denote $b_0.parent$. Let m denote the GENERIC message for b_0. According to Lemma A.4, $b_0'.view < v_d$ and b_0 is proposed during view change. Since qc_0 consists of $2f+1$ votes, at least one replica $p_i \in S$ has voted for b_0 in view v_d. Let b_{lock} denote the locked block lb of p_i when voting for b_0. Note that p_i updates its lb only after receiving a *lockedQC* for a block of a higher rank than its locked block. Then we know that $rank(b_{lock}) \geq rank(b)$. Note that $b_{lock}.view < v_d$. According to Lemma A.6 and the inductive hypothesis, b_{lock} must be either equal to b or an extension of b. Then p_i votes for b_0 only if one of the following conditions is satisfied:

1) $b_0.stable = b_0'.stable$, $m.justfy$ is a wQC for b_0', $b_0'.view < v_d$ and $rank(b_0'.stable) \geq rank(b_{lock})$ (ln 16 in Algorithm 3).
2) $b_0.stable = b_0'$, $m.justfy$ is an rQC for b_0', $b_0'.view < v_d$, and $rank(b_0') \geq rank(b_{lock})$ (ln 17 in Algorithm 3).

If condition 1) is satisfied, then according to Lemma A.3, b_0' is an extension of $b_0'.stable$ and at least one correct replica has received an rQC for $b_0'.stable$. Note that $rank(b_0'.stable) \geq rank(b_{lock})$. According to Lemma A.1 and the inductive hypothesis, $b_0'.stable$ is equal to b or an extension of b. Hence, b_0 must be an extension of b.

If condition 2) is satisfied, then $rank(b_0') \geq rank(b_{lock}) \geq rank(b)$ and $m.justify$ is an rQC for b_0'. According to Lemma A.1 and the inductive hypothesis, b_0' is either equal to b or an extension of b.

Either way, b_0 must be an extension of b. Note that an rQC for d is formed in view v_d. According to Lemma A.5, we know that d is equal to b_0 or an extension of b_0. Therefore, d must be an extension of b and the property holds in view $v+k$. This completes the proof of the lemma. □

Theorem A.8. *(safety) If b and d are conflicting blocks, then they cannot be committed each by at least one correct replica.*

Proof. Suppose that a *commitQC* is formed for both b and d. According to Lemma A.2, there must exist rQCs for both b and d, each received by at least one correct replica. If $b.view = d.view$, then according to Lemma A.6, rQCs for both b and d cannot be formed. If $b.view \neq d.view$, w.l.o.g., we assume that $rank(b) < rank(d)$. According to Lemma A.7, an rQC for d cannot be formed in view $d.view$. Hence, no *commitQC* for d can be formed in view $d.view$. In both cases, *commitQC* for both b and d cannot be formed. □

Theorem A.9. *(liveness) After GST, there exists a bounded time period T_f such that if the leader of view v is correct and all correct replicas remain in view v during T_f, then a decision is reached.*

Proof. Suppose after GST, in a new view v, the leader p_i is correct. Then p_i can collect a set M of $2f + 1$ VIEW-CHANGE messages from correct replicas and broadcast a new block b in a message $m = \langle \text{GENERIC}, b, qc \rangle$.

Let b' denote $b.parent$. Let b_{high} denote the block of highest rank locked by at least one correct replica. Note that a correct replica locks b_{high} only after receiving a *lockedQC* qc for it. Let b_1 denote QCBLOCK(qc). Then we know that $b_1.parent = b_1.stable = b_{high}$ and a set S of at least $f + 1$ correct replicas have voted for b_1. Therefore, at least one message in M is sent by a replica $p_j \in S$. According to Algorithm 2 and Algorithm 3, a correct replica votes for block b_1 only after receiving an rQC for b_{high} and QC_r of the replica is the rQC of highest rank received by the replica. Thus, the rank of the rQC qc_j sent in VIEW-CHANGE message by p_j is no less than that of b_{high}. From Algorithm 3, there are two cases for b: (1) $b.stable = b'$, qc is an rQC for b' and $rank(qc) \geq rank(qc_j)$; (2) $b.stable = b'.stable$, qc is a wQC for b' and $rank(b'.stable) \geq rank(qc_j)$. In case (1), b will be voted for by all the correct replicas as conditions on ln 17 of Algorithm 3 are satisfied. In case (2), b will be voted for by all the correct replicas as conditions on ln 16 of Algorithm 3 are satisfied.

If all correct replicas are synchronized in their view, p_i is able to form a QC for b and generate new blocks. All correct replicas will vote for the new blocks proposed by p_i. Therefore a *commitQC* for b can be formed by p_i, leading to a new decision. Hence, after GST, the duration T_f for these phases to complete is of bounded length. This completes the proof of the theorem. □

B Dashing2

We present the pseudocode and the correctness proof of Dashing2 in this section.

Dashing2 Details. Compared with Dashing1, an sQC is used as a certificate for a fast path in Dashing2. We present in Algorithm 6 and Algorithm 7 the normal case operation and view change protocol of Dashing2, respectively. The utility functions are presented in Algorithm 5. Dashing2 follows the notation of Dashing1. rQCs and sQCs are collectively called *qualified* QCs in this section.

Normal case protocol (Algorithm 6). Similar to Dashing1, in each phase, the leader broadcasts a block b in message $\langle \text{GENERIC}, b, qc_{high} \rangle$ to all replicas and waits for signed responses from the replicas. qc_{high} is the last QC the leader receives (either a wQC, an rQC, or an sQC). After collecting $f + 1$ matching votes, the leader starts a timer Δ_2 (ln 6). The timer is used to determine if the leader can form an rQC or an sQC in time. After Δ_2 expires, the leader combines the signatures in the votes into qc_{high} for the next phase.

Upon receiving a $\langle \text{GENERIC}, b, \pi \rangle$ message from the leader, each replica p_i first verifies whether b is well-formed and proposed during normal operation (ln 16-17), i.e., b has a higher rank than its parent block b', $b.height = b'.height + 1$, and b' and b are proposed in the same view. Let b'' denote the parent of b'. We distinguish two cases:

- If the π field is a wQC for b' (ln 19-22), p_i verifies if the stable block of b and b' are the same block such that b indeed extends b'. p_i also verifies if b, b', b'', and $b.stable$ are all proposed in the same view and p_i has previously voted for b'. If so, p_i updates its local parameter QC_w to π and creates a signature for b (Algorithm 5, ln 13).
- If π is an rQC or an sQC for b' (ln 23-25), p_i verifies if the stable block of b is b', b' does not have a lower rank than vb, and b' does not have a lower rank than the QC_r of p_i. If so, p_i updates its local parameter QC_r to π and generates a signature (Algorithm 5, ln 10 and ln 15). If π is an rQC, b'' has a qualified QC, and b'' and b are proposed in the same view, then p_i commits block b'' and delivers transactions in b'' (Algorithm 5, ln 11-12). If π is an sQC, b'' has a qualified QC, and b'' and b are proposed in the same view, then p_i commits block b' and delivers the transactions in b' (Algorithm 5, ln 14-15).

In both cases, the replica updates its vb to b, and sends its signature to the leader.

View change protocol (Algorithm 7). Every replica starts timer Δ_1 for the first transaction in its queue. If the transaction is not processed before Δ_1 expires, the replica triggers view change. In particular, the replica sends a $\langle \text{VIEW-CHANGE}, vb, (QC_r, QC_w) \rangle$ message to the leader (Algorithm 6, ln 30). Upon receiving $n - f$ VIEW-CHANGE messages (denoted as M), the leader chooses a block to extend based on the output of SAFEBLOCK(M) in Algorithm 5.

We now describe the procedure in more detail. Below, all number of lines is referred to as that in Algorithm 5. First, the leader obtains a block b_1 with a QC that has the highest rank (ln 17-18). The leader then obtains a block b_0 with a wQC vc such that b_0, $b_0.parent$ and $b_0.stable$ are proposed in the same view, and among all the blocks with weak QCs, b_0 has the highest stable block (ln 19-24). The leader also obtains block b_2 such that b_2 is contained in more than $f + 1$ VIEW-CHANGE messages in M. If no such block exists, b_2 is set to $\perp$ (ln 18 and ln 25-26). Then the leader checks if the rank of the stable block of b_2 is no less than that of b_1 (ln 27). If so, the leader selects b_0 to extend. Otherwise, the leader checks if the rank of the stable block of b_0 is no less than that of b_1 (ln 28). If so, the leader will extend b_0. If neither is satisfied, the leader chooses b_1 to extend (ln 29).

Then the leader extends the selected block with a block b and broadcasts b to the replicas (ln 4-5 of Algorithm 7).

Upon receiving a $\langle$NEW-VIEW$, b, M\rangle$ message from a new leader, each replica p_i verifies b basing on the output of SAFEBLOCK(M) (ln 14-18). If b is a block extending the output block of SAFEBLOCK(M), then p_i votes for b (ln 16 and ln 18).

Algorithm 5: Utilities for Dashing2

1 **procedure** CREATEBLOCK(b', v, req, qc)
2     $b.pl \leftarrow hash(b'), b.parent \leftarrow b', b.req \leftarrow req,$
3     $b.height \leftarrow b'.height{+}1, b.view \leftarrow v$
4   **if** qc is a wQC or $\perp$ **then** $b.sl \leftarrow b'.sl, b.stable \leftarrow b'.stable$
5   **then return** b
6   **if** qc is an rQC or an sQC **then** $b.sl \leftarrow hash(b')$ **return** b
7 **procedure** STATEUPDATE(QC_w, QC_r, qc)
8     $b' \leftarrow$ QCBLOCK$(qc), b'' \leftarrow b'.parent,$
9     $b_0 \leftarrow$ QCBLOCK$(QC_w), b_{high} \leftarrow$ QCBLOCK(QC_r)
10   **if** qc is an rQC **then**
11         $QC_r \leftarrow qc$
12   **if** $b'.stable = b''$ **and** $b''.view = b'.view$ **then**
13   deliver the transactions in b''
14   **if** qc is a wQC **then** $QC_w \leftarrow qc$
15   **if** qc is an sQC **and** $b'.stable = b''$ **and** $b'.view = b''.view$
16   **then** $QC_r \leftarrow qc$, deliver the transactions in b'
17 **procedure** SAFEBLOCK(M)
18     $b_0 \leftarrow \perp, b_1 \leftarrow \perp$
19     $qc_{high} \leftarrow$ the qualified QC of highest rank contained in M
20     $b_1 \leftarrow$QCBLOCK$(qc_{high}), b \leftarrow$CREATEBLOCK$(b_1, cview, req, qc_{high})$
21   **for** a wQC $qc \in M.justify$
22         $d \leftarrow$ QCBLOCK$(qc), d' \leftarrow d.parent, d_s \leftarrow d.stable$
23   **if** $d_s.view = d'.view = d.view$ **then**
24   **if** $rank(d_s) > rank(b_0.stable)$ **then** $vc \leftarrow qc, b_0 \leftarrow d$
25   **if** $rank(d_s) = rank(b_0.stable)$ **and** $rank(d) > rank(b_0)$
26   **then** $vc \leftarrow qc, b_0 \leftarrow d$
27   **for** $d \in M.block$
28   **if** $num(d, M.block) \geq f+1$ **then** $b_1 \leftarrow d$
29   **if** $rank(b_1.stable) \geq rank(b_2)$ **then return** $(b_1, \perp)$
30   **else if** $rank(b_0.stable) \geq rank(b_1)$ **then return** (b_0, vc)
31   **return** (b_2, qc_{high})

B.1 Correctness of Dashing2

We first introduce some notation we use for the proof. Let b' and b denote two blocks such that $b.parent = b'$ and $b'.view = b.view$. According to Algorithm 6, after receiving a GENERIC message $\langle$GENERIC$, b, qc\rangle$, a correct replica votes for b only if (1) $b.stable = b'$ and qc is an rQC or an sQC for b' (ln 23-25); or (2) $b.stable = b'.stable$ and qc is a wQC for b' (ln 19-22). In both cases, we say that qc and b are *matching*.

Let b' and b denote two consecutive blocks. In Algorithm 5, a replica p_i commits b only after receiving a certificate qc and one of the following conditions is satisfied:

(1) qc is an rQC for b' such that $b'.stable = b'.parent = b$ and $b.view = b'.view$ (ln 9-12);

(2) qc is an sQC for b, $b.stable = b.parent$ and $b.parent.view = b.view$ (ln 14-15).

Algorithm 6: Normal case protocol of Dashing2 for p_i

1 **initialization:** $cview \leftarrow 1$, vb, QC_w, QC_r are initialized to $\perp$
2 Start a timer Δ_1 for the first request in the queue of pending transactions
3 ▷ **GENERIC phase:**
4 **as** a leader
5   wait for votes for b: $M \leftarrow \{\sigma | \sigma$ is a signature for $\langle$GENERIC$, b, \perp\rangle\}$
6   **upon** $|M| = f+1$ **then** set a start timer Δ_2
7   **upon** Δ_2 *timeout* **then** $qc_{high} \leftarrow$ QCCREATE(M)
8         $b \leftarrow$CREATEBLOCK$(b, cview, req, qc_{high})$
9   broadcast $m = \langle$GENERIC$, b, qc_{high}\rangle$
10   **if** qc_{high} is a wQC **then** $QC_w \leftarrow qc_{high}$
11   **if** qc_{high} is an rQC or an sQC **then** $QC_r \leftarrow qc_{high}$
12 **as** a replica
13   **wait for** $m = \langle$GENERIC$, b, \pi\rangle$ from LEADER$(cview)$
14         $b' \leftarrow b.parent, b'' \leftarrow b'.parent, b_s \leftarrow b.stable,$
15         $b_{high} \leftarrow$QCBLOCK$(QC_r), m \leftarrow \langle$GENERIC$, b, \perp\rangle$
16   **if** $rank(b') \geq rank(b)$ **or** $b.height \neq b'.height + 1$ **or**
17           $b'.view \neq cview$ **then discard the message**
18   **if** π is a wQC for b' **and** $b.sl = b'.sl$ **and** $rank(b_s) \geq$
19           $rank(b_{high})$ **and** $b_s.view = b''.view = b'.view = cview$
20   **and** $b' = vb$ **then** $vb \leftarrow b$, STATEUPDATE(QC_w, QC_r, π)
21   **if** π is an rQC or an sQC for b' **and** $b.stable = b'$
22   **and** $rank(b') \geq rank(vb)$ **and** $rank(b') \geq rank(b_{high})$
23   **then** $vb \leftarrow b$, STATEUPDATE(QC_w, QC_r, π)
24   **if** $vb = b$ **then** send QCVOTE(m) to LEADER$(cview)$
25 ▷ **NEW-VIEW phase:** switch to this line if Δ_1 *timeout* occurs
26 **as** a replica
27     $cview \leftarrow cview + 1$
28   send $\langle$VIEW-CHANGE$, vb, (QC_r, QC_w)\rangle$ to LEADER$(cview)$

In both cases, qc is a *commitQC* for b.

Lemma B.1. *Suppose a block b has been voted for by a correct replica, then*

(1) any block d on the branch led by b has been voted for by at least one correct replica and $d.parent.height + 1 = d.height$;

(2) if d and d_c are two blocks on the branch led by b such that $d_c.parent = d$ and $d_c.view = d.view = v$, then we have that (i) at least one correct replica has received a certificate (wQC, rQC, or sQC) qc_d for d, where qc_d and d_c are matching; (ii) if the view of the parent block of d is lower than v, then at least one correct replica has received a qualified QC for d and $d_c.stable = d$.

Proof. Let d denote a block on the branch led by b.

(1) We prove claim (1) by induction for d. If $d = b$, then d has been voted for by at least one correct replica.

If $d \neq b$ and any block higher than d on the branch led by b has been voted for by at least one correct replica, then we need to prove that d is voted for by at least one correct replica. In this situation, there exists a block d_c on the branch led by b such that $d_c.parent = d$ and d_c has been voted for by at least one correct replica p_i. According to Algorithm 2 and

Algorithm 7: View change protocol of Dashing2 for p_i

▷ **VIEW-CHANGE phase**
as a new leader
  //M is a set of $n-f$ VIEW-CHANGE messages collected by the new leader
  $(b', qc) \leftarrow$ SAFEBLOCK(M), $b \leftarrow$ CREATEBLOCK$(b', cview, req, qc)$
  broadcast $m = \langle$NEW-VIEW$, b, M\rangle$
  //switch to normal case protocol
as a replica
  wait for $m = \langle$NEW-VIEW$, b, \pi\rangle$ from LEADER$(cview)$
  $b' \leftarrow b.parent$, $b_s \leftarrow b.stable$, $b_{high} \leftarrow$ QCBLOCK(QC_r),
  $m \leftarrow \langle$GENERIC$, b, \perp\rangle$
  if $b'.view \geq cview$ **or** $rank(b') \geq rank(b)$ **or** $b.height \neq$
   $b'.height + 1$ **then discard the message**
  if $M \in \pi$ **then**
    $(b_p, qc) \leftarrow$ SAFEBLOCK(M), $m \leftarrow \langle$GENERIC$, b, \perp\rangle$
    if $b_p = b'$ **and** qc is a wQC or $\perp$ **and** $b.stable = b'.stable$
    then send QCVOTE(m) to LEADER$(cview)$
    if $b_p = b'$ **and** qc is an rQC or sQC **and** $b.stable = b'$
    then send QCVOTE(m) to LEADER$(cview)$
//switch to normal case protocol. Three consecutive rQCs are required for the first block proposed during the view change.
▷ **NEW-VIEW phase:** switch to NEW-VIEW phase if Δ_1 times out

Algorithm 3, $rank(d) < rank(d_c)$ and $d_c.height = d.height{+}1$. Therefore, $d.view \leq d_c.view$.

We now distinguish two cases: $d.view = d_c.view$ and $d.view < d_c.view$.

If $d.view = d_c.view$, then p_i has received a qc_d for d, where qc_d and d_c are matching according to Algorithm 6. As qc_d consists of at least $f+1$ votes, at least one correct replica has voted for d and $d.parent.height + 1 = d.height$.

If $d.view < d_c.view$, then from Algorithm 7 we know that d_c is proposed in a NEW-VIEW message m in view $d_c.view$ and $m.justify$ contains a set M of $2f+1$ VIEW-CHANGE messages for view $d_c.view$. Then p_i votes for d_c if (i) a wQC, an rQC or an sQC for d is provided by a replica in M, or (ii) for $f+1$ messages in M, the *block* fields are all set to d. In either case, d has been voted for by at least one correct replica. This completes the proof of claim (1).

(2) Based on claim (1), at least one correct replica p_i has voted for d_c. (i) If $d_c.view = d.view = v$, then d_c is proposed during normal case operation. According to ln 19 and ln 23 of Algorithm 6, p_i has received a certificate (wQC, rQC, or sQC) qc_d for d before voting for d_c, where d and d_c are matching. (ii) Meanwhile, according to ln 16-25 of Algorithm 6, if $d.parent.view < v$, then p_i votes for d_c only if p_i has received an rQC or an sQC for d and $d_c.stable = d$. □

Lemma B.2. *Suppose that qc_b and qc_d are two qualified QCs, and each is received by at least one correct replica. Let b and d be QCBLOCK(qc_b) and QCBLOCK(qc_d), respectively. If b and d are two conflicting blocks, then $rank(b) \neq rank(d)$.*

Proof. Assume, on the contrary, that $rank(b) = rank(d)$. Let v denote the view of b and d. As each qualified QC consists of at least $2f+1$ votes, at least one correct replica has voted for both b and d. Let b' and d' denote the parent block of b and d, respectively. Since a correct replica votes for at most one block of each height during normal case operation, at least one of b and d is proposed during view change. Therefore, $b'.view < v$ or $d'.view < v$. Now we consider two cases:

(1) $b'.view < v$ and $d'.view < v$. According to Algorithm 7, a correct replica p_i votes for at most one block that extends a block proposed in a lower view. Hence, b equals d.

(2) ($b'.view < v$ and $d'.view = v$) or ($b'.view = v$ and $d'.view < v$). If $b'.view < v$ and $d'.view = v$, then there exists a block d_0 of lowest height on the branch led by d such that $d_0.view = v$. Hence, the view of $d_0.parent$ is lower than v. Let d_0' denote a block on the branch led by d such that $d_0'.parent = d_0$. By Lemma B.1, at least one correct replica p_i has voted for d_0'. According to ln 16-25 in Algorithm 6, p_i has received an rQC or an sQC for d_0. Note that the view of $d_0.parent$ is lower than v. Then d_0 and b must be the same block according to case (1). Therefore, d is an extension of b. The proof is similar in the case where $b'.view = v$ and $d'.view < v$.

In both cases, d and b are either the same block or on the same branch, contradicting the condition that they are conflicting blocks. Therefore, $rank(b) \neq rank(d)$. □

Lemma B.3. *If a correct replica has voted for d and set its vb to d, then d must be an extension of $d.stable$ and at least one correct replica has received a qualified QC for $d.stable$.*

Proof. Let d_0 denote $d.parent$. Let p_i denote a correct replica that has voted for d and set its vb to d. According to ln 19-25 of Algorithm 6, p_i has received a certificate qc for d_0, where qc and d are matching. We distinguish two cases.

(1) qc is an rQC or an sQC for d_0 and $d.stable = d_0$ (ln 23-25 in Algorithm 6). In this case, d is an extension of $d.stable$ and p_i received a qualified QC for $d.stable$.

(2) qc is a wQC for d_0 and $d.stable = d_0.stable$ (ln 19-22 in Algorithm 6). Let d_v denote the block of lowest height on the branch led by d such that $d_v.stable = d.stable$. Let d_v' denote $d_v.parent$. Then $d_v.stable \neq d_v'.stable$. According to Lemma B.1, at least one correct replica p_j has voted for d_v since $d_v.stable \neq d_v'.stable$. Note p_j votes for d_v only if one of the following conditions holds: i) $d_v.stable = d_v'.stable$; ii) $d_v.stable = d_v'$ and p_i receives a qualified QC for d_v'. In this case, $d_v.stable = d.stable = d_v'$, d is an extension of $d.stable$, and p_j has received a qualified QC for $d.stable$.

Either way, d is an extension of $d.stable$ and at least one correct replica has received a qualified QC for $d.stable$. □

Lemma B.4. *If a qualified QC is formed in view v, then there exists only one block b of lowest rank for which a qualified QC is formed in view v, and we have that:*

(1) the view of $b.parent$ is lower than v;

(2) if there exists a qualified QC for b_1, $b_1.view = v$, and the view of $b_1.parent$ is lower than v, then b_1 equals b;

(3) if there exists a qualified QC for d and $d.view = v$, then d equals b or d is an extension of b.

Proof. If a qualified QC is formed in view v, then there exists only one block b of lowest rank for which a qualified QC is formed in view v (according to Lemma B.2).

(1) Let b_v denote the block of lowest height such that $b_v.view = v$ on the branch led by b. We have $b_v.height \leq b.height$ and the view of $b_v.parent$ is lower than v. If $b_v \neq b$, then there exists a block b'_v on the branch led by b such that $b'_v.parent = b_v$ and $b'_v.view = b_v.view = v$. From Lemma B.1, at least one correct replica p_i has received an rQC or an sQC for b_v. Thus, b_v is a block of lower rank than b and a qualified QC for b_v is formed in view v, contradicting the definition of b. Hence, we have $b_v = b$ and the view of $b.parent$ is lower than v.

(2) If there exists a qualified QC for b_1, at least one correct replica has voted for both b_1 and b in view v. According to Algorithm 7, in view v, a correct replica only votes for one block that extends a block proposed in a lower view than v. Therefore, it must hold that $b_1 = b$.

(3) There exists a qualified QC for d and $d.view = v$. Let d_0 denote the block of lowest height on the branch led by d such that $d_0.view = v$. Then the view of the parent block of d_0 is lower than v. From Lemma B.1, a correct replica has received a qualified QC for d_0. According to claim (2), we know d_0 equals b. Therefore, d equals b or d is an extension of b. □

Lemma B.5. *For any qualified QC qc, if* QCBLOCK(qc) $= b$ *and $b.view = v$, then any block proposed in view v on the branch led by b has been voted for by at least $f + 1$ correct replicas.*

Proof. Assume that block d is on the branch led by b such that $d.view = v$ and fewer than $f + 1$ correct replicas have voted for d. We immediately know that a qualified QC for d cannot be formed. Let d' denote a block such that $d'.parent = d$. So, a correct replica p_i votes for d' only if a wQC for d is received and p_i has voted for d. Since fewer than $f+1$ correct replicas have voted for d, a qualified QC for d or any extensions of d (including b) cannot be formed (a contradiction). □

Lemma B.6. *For any two qualified QCs qc_1 and qc_2, let b_1 and b_2 be* QCBLOCK(qc_1) *and* QCBLOCK(qc_2)*, respectively. If b_1 is conflicting with b_2, then $b_1.view \neq b_2.view$.*

Proof. Assume, on the contrary, that $b_1.view = b_2.view = v$. Let b be the block of lowest height for which a qualified QC was formed in view v. Then according to Lemma B.4, either b_1 or b_2 equals b or is an extension of b. Hence, $b_1.height \geq b.height$ and $b_2.height \geq b.height$. We consider three cases:

(1) If $b_1.height = b.height$ or $b_2.height = b.height$, then b_1 equals b or b_2 equals b. Therefore, b_1 and b_2 are the same block or they are on the same branch.

(2) If $b.height < b_1.height$, $b.height < b_2.height$, and $b_1.height = b_2.height$, then according to Lemma B.2, b_1 and b_2 must be the same block.

(3) If $b.height < b_1.height$, $b.height < b_2.height$, and $b_1.height \neq b_2.height$, then b_1 and b_2 are extensions of b. W.l.o.g., we assume that $b_1.height < b_2.height$. Let b'_2 denote a block on the branch led by b_2 such that $b'_2.height = b_1.height$. Then b'_2 is an extension of b and b'_2 and b_1 are blocks proposed during the normal case operation in view v. According to Lemma B.5, at least $f + 1$ correct replicas have voted for b'_2. Since each rQC consists of at least $2f + 1$ votes, at least one correct replica has voted for both b'_2 and b_1. Note that during the normal case operation, a correct replica votes for at most one block of each height. Therefore, it holds that b'_2 and b_1 must be either the same block or on the same branch.

In all cases, b_1 and b_2 are the same block or are blocks on the same branch, contradicting the condition that they are conflicting blocks. Therefore, $b_1.view \neq b_2.view$. □

Lemma B.7. *Suppose that all the correct replicas have voted for b in view v, $b.parent = b.stable$ and $b.parent$ is proposed in view v. If a correct replica has received a wQC qc for d such that $rank(d.stable) \geq rank(b.parent)$, and d, $d.parent$, and $d.stable$ are blocks proposed in view v, then d equals b or d is an extension of b.*

Proof. As b, $b.parent$, d, and $d.parent$ are all blocks proposed in view v, b and d are blocks proposed during normal case operation in view v. According to Algorithm 6, we know that if a correct replica has voted for d, the replica will set its vb to d at the same time. Since qc consists of $f+1$ votes, at least one correct replica has voted for d. From Lemma B.3, d is an extension of $d.stable$ and at least one correct replica has received a qualified QC for $d.stable$. Now we consider two cases:

(1) $rank(d.stable) = rank(b.parent)$. Since $b.parent = b.stable$, any correct replica votes for b only after receiving a qualified QC for $b.parent$. Then $d.stable = b.parent$ and $d.height \geq b.height$ (according to Lemma B.2). Let d' denote the block on the branch led by d such that $d'.height = b.height$. Then at least one correct replica has voted for d' in view v according to Lemma B.1. Since correct replicas vote for at most one block of each height during normal operation in a view, d' must be equal to b. Therefore, d equals b or d is an extension of b.

(2) $rank(d.stable) > rank(b.parent)$. It is straightforward to see that $rank(d.stable) \geq rank(b)$. According to Lemma B.6, $d.stable$ is equal to b or $d.stable$ is an extension of b. Hence, d is an extension of b. □

Lemma B.8. *For a commitQC qc for b and a qualified QC qc_d for d, if $rank(b) < rank(d)$, then d must be an extension of b.*

Proof. Let v denote $b.view$ and v_d denote $d.view$. As $rank(d) > rank(b)$, then $v_d \geq v$. Let b' denote QCBLOCK(qc). Since qc is a *commitQC* for b, there are two conditions: (1) qc is an rQC for b', $b'.stable = b'.parent = b$ and $b'.view = v$; (2) qc is an sQC for b, $b.parent = b.stable$ and the view of $b.parent$ equals v.

We prove the lemma by induction over the view v_d, starting from view v.

Base case: Suppose $v_d = v$. From Lemma B.6, for condition (1) or (2), d must be an extension of b.

Inductive case: Assume this property holds for view v_d from v to $v + k - 1$ for some $k \geq 1$. We now prove that it holds for $v_d = v + k$.

Let d_0 denote the block of lowest height on the branch led by d such that $d_0.view = v_d$. Then the view of the parent block of d_0 is lower than v_d, d_0 is proposed during view change in view v_d, and d_0 is voted for by at least one correct replica p_i (Lemma B.1).

Let m denote the NEW-VIEW message for d_0. According to Algorithm 7, $m.justify$ is a set M of $2f + 1$ VIEW-CHANGE messages for view v_d. Let qc_1 denote the qualified QC with the highest rank contained in $M.justify$ and let b_1 denote QCBLOCK(qc_1). For all the wQCs contained in $M.justify$, a correct replica chooses the wQC for a block with the highest stable block according to ln 19-24 in Algorithm 5 and sets the wQC as vc. Let b_0 denote QCBLOCK(vc). Note that b_0, $b_0.parent$ and $b_0.stable$ are proposed in the same view. Then b_0 is a block proposed during the normal case operation. Let b_2 denote the block which is included in more than $f + 1$ messages in M. If no such block exists, b_2 is set to $\perp$.

In view v_d, p_i votes for d_0 if $d_0' = d_0.parent$, $d_0'.view < v_d$, $d_0'.height + 1 = d_0.height$ and one of the following conditions are satisfied:

i) $d_0' = b_2$, $rank(b_2.stable) \geq rank(b_1)$ (ln 27 in Algorithm 5).
ii) $d_0' = b_0$, i) is not satisfied and $rank(b_0.stable) \geq rank(b_1)$ (ln 28 in Algorithm 5).
iii) $d_0' = b_1$, i) and ii) are not satisfied (ln 29 in Algorithm 5) .

Note that b_0 is a block proposed during the normal case operation in view $b_0.view$. Since a wQC consists of $f + 1$ votes, at least one is sent by a correct replica. Hence, at least one correct replica has voted for b_0 and sets its vb as b_0. According to Lemma B.3, b_0 is an extension of $b_0.stable$ and at least one correct replica has received a qualified QC for $b_0.stable$.

Next, we prove the property holds in view $v + k$ for the two situations for *commitQC*, respectively.

(1) qc is an rQC. Let S denote the set of correct replicas that have received a qualified QC for b in view v. Since in view v correct replicas vote for b' only after receiving a qualified QC for b, we have $|S| \geq f + 1$. Note that a correct replica updates its QC_r only with a qualified QC with a higher rank. Thus, for any VIEW-CHANGE message sent by a replica in S, the $justify$ field is set to a qualified QC with the same or a higher rank than b. Since M consists of $2f + 1$ messages, at least one message in M is sent by a replica in S. Therefore, $rank(b_1) \geq rank(b)$ and $b_1.view < v_d$.

According to the inductive hypothesis, b_1 must be equal to b or an extension of b. Therefore, if condition iii) is satisfied, d_0 must be an extension of b. If condition i) is satisfied, then $rank(b_2) > rank(b_1)$ and $rank(b_2.stable) \geq rank(b_1)$. Since at least one correct replica has set its vb to b_2, then b_2 is an extension of $b_2.stable$ and a qualified QC qc_2 for $b_2.stable$ has been received by a correct replica from Lemma B.3. According to the inductive hypothesis, b_2 is an extension of b. Hence, d_0' is an extension of b. If condition ii) is satisfied, then $rank(b_0.stable) \geq rank(b_1)$. Note that b_0 is an extension of $b_0.stable$ and at least one correct replica has received a qualified QC for $b_0.stable$. Thus, b_0 is an extension of b (according to the inductive hypothesis). Therefore, d_0' is an extension of b. No matter which condition is satisfied, both d_0 and d must be extensions of d_0' and extensions of b.

(2) qc is an sQC, the view of $b.parent$ equals v and $b.parent = b.stable$. Since qc consists of $3f + 1$ votes, all the correct replicas have received a qualified QC for $b.parent$, changed its QC_r to a qualified QC for $b.parent$, and voted for b in view v. Let V denote the set of correct senders of messages in M. It is clear that $|V| \geq f + 1$. Since correct replicas only change their QC_r to a qualified QC of a higher rank, we have $rank(b_1) \geq rank(b.parent)$.

(a) If $rank(b_1) \geq rank(b)$, then from Lemma B.2 and the induction hypothesis, b_1 is equal to b or b_1 is an extension of b. If condition iii) is satisfied, then d_0 and d are extensions of b. If condition i) or ii) is satisfied, at least one correct replica has voted for d_0' and set its vb to d_0', and $rank(d_0'.stable) \geq rank(b_1)$. According to Lemma B.3, d_0' is an extension of $d_0'.stable$ and at least one correct replica has received a qualified QC for $d_0'.stable$. Again, from the induction hypothesis, $d_0'.stable$ is equal to b or $d_0'.stable$ is an extension of b. Therefore, d_0 and d are extensions of b.

(b) If $rank(b_1) < rank(b)$, then $rank(b_1) = rank(b.parent)$. If $b_2 = b$, then condition i) is satisfied. Hence, d_0' equals b and d_0 and d are extensions of b.

If $b_2 \neq b$, then there exists a correct replica p_i in V such that when p_i sent a VIEW-CHANGE message for v_d, its last voted block vb is b_e and $b_e \neq b$. Let b_e' denote $b_e.parent$. According to ln 19-22 in Algorithm 6, p_i has received a wQC qc_e for b_e', $rank(b_e') \geq rank(b)$, and $rank(b_e') \geq rank(b.parent)$. If $b_e'.view = v$, then b_e' equals b or b_e' is an extension of b from Lemma B.7. If $b_e'.view > v$, then the view of $b_e'.stable$ is higher than v. From Lemma B.3, b_e' is an extension of $b_e'.stable$ and a correct replica has received a qualified QC for $b_e'.stable$. From the inductive hypothesis, as $rank(b_e'.stable) > rank(b)$, it must hold that $b_e'.stable$ is an extension of b. Therefore, b_e must be an extension of b, b_2 is set to $\perp$ or

b_2 is an extension of b. If condition i) is satisfied, d'_0 equals b_2. We know that p_i has sent qc_e in its VIEW-CHANGE message. Then $rank(b_1.stable) \geq rank(b.parent)$. If condition i) is not satisfied, condition ii) is satisfied and d'_0 equals b_1. Note that a wQC for b_1 is included in M and b_1 is proposed during normal case operation. Similar to b'_e, b_1 must be an extension of b. Either way, d'_0 is equal to or an extension of b. Thus, d_0 and d are extensions of b.

Therefore, d must be an extension of b and the property holds in view $v+k$ based on Case (1) and Case (2). This completes the proof of the lemma. □

Theorem B.9. *(safety) If b and d are conflicting blocks, then not both can be committed by at least one correct replica.*

Proof. Suppose that there exist *commitQC*'s for both b and d. According to Lemma B.1, a qualified QC must have been formed for both b and d. From Lemma B.2, if $rank(b) = rank(d)$, only one qualified QC for b and d can be formed in the same view. For the case where $rank(b) \neq rank(d)$, we assume w.l.o.g. that $rank(b) < rank(d)$. From Lemma A.7, we know that a qualified QC for d cannot be formed in view $d.view$. This completes the proof of the theorem. □

Theorem B.10. *(liveness) After GST, there exists a bounded time period T_f such that if the leader of view v is correct and all correct replicas remain in view v during T_f, then a decision is reached.*

Proof. Suppose after GST, in a new view v, the leader p_i is correct. Then p_i can collect a set M of $2f+1$ VIEW-CHANGE messages from correct replicas and broadcast a new block b_v in a NEW-VIEW message m. Since $m.justify$ contains M, every correct replica can verify the block b_v using a call SAFEBLOCK(M).

Under the assumption that all correct replicas are synchronized in their view, p_i is able to form a *QC* for b and generate new blocks. All correct replicas will vote for the new blocks from p_i. Therefore a *commitQC* for b can be formed by p_i and any correct replica will vote for b. After GST, the duration T_f for these phases to complete is of bounded length. □

C The Underlying BFT Protocol in Star

C.1 The Consensus Protocol Implemented in Star

We now describe the concrete atomic broadcast protocol that we implemented in Star. We use a variant of PBFT that differs from PBFT in two minor aspects. The protocol we will describe in the following is not presented in its general manner but instead takes as input the output from the transmission process.

Normal case operation. We first describe the normal case protocol.

Step 1: Pre-prepare. The leader checks whether $|W[le]| \geq n-f$. If so, it proposes a block B and broadcasts a $\langle$PRE-PREPARE$, v, B\rangle$ message to all replicas.

The block B is of the form $\langle v, cmd, height\rangle$, where v is the current view number, $B.cmd = W[le]$, and $B.height = le$. We directly use $B.height$ as the sequence number for B in the protocol.

Step 2: Prepare. Replica receives a valid PRE-PREPARE message for block B and broadcasts a PREPARE message.

After receiving a PRE-PREPARE message $\langle$PRE-PREPARE$, v, B\rangle$ from the leader, a replica p_j first verifies whether 1) its current view is v, 2) $B.cmd$ consists of at least $n-f$ wQCs or rQCs for epoch e, and 3) p_j has not voted for a block $B.height$ in the current view. Then p_j broadcasts a signed PREPARE message $\langle$PREPARE$, v, hash(B)\rangle$. The replica also updates its W queue if any QC included in $B.cmd$ is not in $W[B.height]$.

Step 3: Commit. Replica receives $n-f$ PREPARE messages for B and broadcasts a COMMIT message.

After receiving $n-f$ matching PREPARE messages with the same $hash(B)$, replica p_j combines the messages into a regular certificate for B, called a *prepare certificate.* Then p_j broadcasts a $\langle$COMMIT$, v, hash(B)\rangle$ message. After receiving $n-f$ COMMIT messages with the same $hash(B)$, p_j *a-delivers* B with sequence number le.

Note that the PRE-PREPARE step and the COMMIT step carry only $hash(B)$ as the message transmitted. The total communication for the normal case operation is thus $O(n^2\lambda)$ where λ is the security parameter.

Checkpointing. After a fixed number of blocks are a-delivered, replicas execute the checkpoint protocol for the garbage collection. Each replica broadcasts a checkpoint message that includes its current system state and the epoch number for the latest a-delivered block. Each replica waits for $n-f$ matching checkpoint messages which form a stable checkpoint. Then the system logs for epoch numbers lower than the stable checkpoint can be deleted.

View change. We now describe the view change protocol. After a correct replica times out, it sends a VIEW-CHANGE message to all replicas. Upon receiving $f+1$ VIEW-CHANGE messages, a replica also broadcasts a VIEW-CHANGE message. The new leader waits for $n-f$ VIEW-CHANGE messages, denoted as M, and then broadcasts a NEW-VIEW message to all replicas.

The VIEW-CHANGE message is of the form $\langle$VIEW-CHANGE$, C, \mathcal{P}\rangle$, where C a stable checkpoint and $\mathcal{P}$ is a set of prepare certificates. For $\mathcal{P}$, a prepare certificate certificate for each epoch number greater than C and lower than the replica's last vote is included.

The NEW-VIEW message is of the form $\langle$NEW-VIEW$, v+1, c, M, \mathcal{PP}\rangle$, where c is the latest stable checkpoint, M is the set of VIEW-CHANGE messages M, and $\mathcal{PP}$ is a set of PRE-PREPARE messages. The set $\mathcal{PP}$ is computed as follows: For each epoch number e between C and the epoch number of any replica's last vote, the new leader creates a new PRE-PREPARE message. If a prepare certificate is provided by any replica in the VIEW-CHANGE message, the PRE-PREPARE message is of the form $\langle$PRE-PREPARE$, v+1, h\rangle$, where h is the hash in the prepare certificate. If none of the replicas provides a prepare certificate, the new leader creates a message $\langle$PRE-PREPARE, $v+1, B\rangle$, where B is of the form $\langle v+1, W[e], e\rangle$.

Upon receiving a NEW-VIEW message, a replica verifies the PRE-PREPARE messages in the $\mathcal{PP}$ field by executing the same procedures as the leader based on M. Then the replicas resume normal operation.

D Correctness of Star

Based on the safety and liveness properties of the underlying atomic broadcast protocol in the consensus process, we now prove the correctness of Star.

According to the Star specification, a set V consisting of transactions in batches $\{\text{QCPROPOSAL}(qc_k)\}_{k\in[1..n-f]}$ delivered (in a deterministic order) by p_i must correspond to the set m (consisting of $n-f$ wQCs $\{qc_k\}_{k\in[1..n-f]}$) a-delivered by p_i from the underlying atomic broadcast protocol. In this case, we simply say V is *associated with m*.

We prove the safety of Star by showing that different sets of transactions cannot be committed together in the same epoch, each by a correct replica. We begin with the following lemma:

Lemma D.1. *If V_i associated with some m is delivered by p_i and V_j associated with the same m is delivered by p_j, then we have $V_i = V_j$.*

Proof. Assume, towards contradiction, that $V_i \neq V_j$. Let $\{qc_k\}_{k\in[1..n-f]}$ be the $n-f$ wQCs contained in m. Then we have that V_i is a union of transactions in proposals $\{b_k\}_{i\in[1..n-f]}$, where we have $b_k = \text{QCPROPOSAL}(qc_k)$. Similarly, V_j is a union of transactions in proposals $\{b'_k\}_{i\in[1..n-f]}$, where $b'_k = \text{QCPROPOSAL}(qc_k)$. Since $V_i \neq V_j$, we have that there exists $k \in [1..n-f]$ such that $b_k \neq b'_k$. Note that qc_k is a wQC for b_k and also a wQC for b'_k. Since $b_k \neq b'_k$, this violates the unforgeability of digital signatures (or threshold signatures). □

Now we are ready to prove safety.

Theorem D.2. *(safety) If a correct replica delivers a transaction tx before delivering tx', then no correct replica delivers a transaction tx' without first delivering tx.*

Proof. Suppose that a correct replica p_i delivers a transaction tx before delivering tx'. Let L_i denote the a-delivered messages log of p_i and TL_i denote the delivered transactions log of p_i. For any correct replica p_j, let L_j denote the a-delivered messages log and TL_j denote the delivered transactions log of p_j. According to the safety of the consensus protocol, either L_i equals L_j or one of L_i and L_j is a prefix of the other. Note that TL_i and TL_j contain transactions associated with messages in the a-delivered messages logs in a deterministic order. According to Lemma D.1, either TL_i equals TL_j or one of TL_i and TL_j is a prefix of the other. This completes the proof of the theorem. □

Theorem D.3. *(liveness) If a transaction tx is submitted to all correct replicas, then all correct replicas eventually deliver tx.*

Proof. If a transaction tx is submitted to all correct replicas, eventually in some epoch, tx is included in the proposal by at least one correct replica. Using the strategy in EPIC (following HoneyBadgerBFT), eventually the wQC wqc for the proposal containing the transaction tx will be sent to the consensus process.

At least $n-f$ wQCs will be a-delivered in the consensus process, and at least $f+1$ wQCs must be proposed by correct replicas. So there is some probability that wqc for tx will be delivered. If the corresponding transaction has been received by a correct replica, then we are done. Otherwise, a correct replica just needs to run the fetch operation to get the corresponding proposal containing tx. Recall that the use of wQC ensures that a correct replica must have stored the corresponding proposal. (If the underlying atomic broadcast only achieves consistency rather than agreement, then we can still use the standard state machine replication mechanism such as state transfer to ensure that all correct replicas deliver the transaction.) □