

Verifying Procedural Programs via Constrained Rewriting Induction

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This paper aims to develop a verification method for procedural programs via a transformation into Logically Constrained Term Rewriting Systems (LCTRSs). To this end, we extend transformation methods based on integer TRSs to handle arbitrary data types, global variables, function calls and arrays, as well as encode safety checks. Then we adapt existing rewriting induction methods to LCTRSs and propose a simple yet effective method to generalize equations. We show that we can automatically verify memory safety and prove correctness of realistic functions. Our approach proves equivalence between two implementations, so in contrast to other works, we do not require an explicit specification in a separate specification language.

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1. INTRODUCTION

Ensuring with certainty that a program always behaves correctly is a hard problem. One approach to this is formal verification—proving with mathematical rigor that all executions of the program will have the expected outcome. Several methods for this have been investigated (see e.g., [Huth and Ryan 2000]). However, classically many of them require expert knowledge to manually prove relevant properties about the code.

Instead, we hope to raise the degree of automation, ideally creating a fully automatic verification / refutation process and tools to raise developer productivity. Indeed, over the last years automatic provers for program verification have flourished, as witnessed, e.g., by tool competitions like SV-COMP [SV-COMP] and the Termination Competition (http://termination-portal.org/wiki/Termination_Competition). Program verification is also recognized in industry, cf. e.g. Facebook's safety prover Infer [Calcagno et al. 2015] or Microsoft's temporal prover T2 [Brockschmidt et al. 2016]. However, these tools generally use specific reasoning techniques for imperative programs and benefit from

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the progress in automated theorem proving over the last decades only to a limited extent. This suggests likely avenues for improvement.

One such avenue is *inductive theorem proving*. This method is well investigated in functional programming [Bundy 2001] and term rewriting, the underlying core calculus of functional programming. To check a functional program f against a specification by a reference implementation f_{spec} , it suffices that $f(\vec{x}) \approx f_{spec}(\vec{x})$ is an *inductive theorem*. Thus, no explicit specification language is needed: giving a (possibly not optimized) reference implementation f_{spec} in the same programming language suffices.

To analyze imperative programs (in C, Java, etc.), recent works have applied transformations into term rewrite systems (e.g., [Otto et al. 2010]). In particular, *constrained rewriting systems* are popular as target language, since logical constraints to model the control flow can be separated from terms to model intermediate states [Furuichi et al. 2008; Falke and Kapur 2009; Sakata et al. 2009; Nakabayashi et al. 2010; Falke et al. 2011]. Unifying existing approaches, Kop and Nishida [2013] have proposed the framework of *logically constrained term rewriting systems (LCTRSs)*.

Aims. The aim of this paper is twofold. First, we propose a new transformation method from procedural programs into constrained term rewriting. This transformation makes it possible to use the many methods available to term rewriting also to analyze imperative programs. Unlike previous methods, we do not limit interest to integer functions.

Second, we develop a verification method for LCTRSs, based on rewriting induction [Reddy 1990]—a well-investigated method of inductive theorem proving—to prove (total) equivalence of two functions. We also supply two generalization techniques, the main one of which is specialized for transformed iterative functions.

The applications are many. First, checking equivalence between different implementations comes to mind. This allows the user to determine automatically if a modification in the program has changed its semantics (see e.g. [Godlin and Strichman 2013; Lahiri et al. 2012]). Proposing equivalent replacements may even be done automatically, via algorithm recognition (see e.g. [Alias and Barthou 2003]).

In compilation, automated equivalence checking can validate correctness of compiler optimizations on a per-instance basis [Necula 2000; Pnueli et al. 1998] or once-and-for-all for a given optimization template [Kundu et al. 2009; Lopes and Monteiro 2016]. Equivalence checking is also used in proofs of secure information flow [Terauchi and Aiken 2005] and can be used to prove safety properties, e.g., memory safety.

Why LCTRSs. Direct support of basic types like the integers, and of constraints to restrict evaluation—features absent in basic TRSs—is essential to handle realistic programs. Unlike earlier constrained rewriting systems, LCTRSs do not limit the underlying theory to (linear) integer arithmetic: we might use (combinations of) arbitrary first-order theories, including, e.g., n -dimensional integer arrays, floating point numbers, and bitvectors. This makes it possible to natively handle sophisticated programs.

Despite the generality, we get strong results on LCTRSs by reducing analysis problems like termination and equivalence to a sequence of satisfiability problems over the underlying theories. Automatic tools—like our tool Ctrl [Kop and Nishida 2015] for rewriting, termination, and inductive theorem proving—can defer such queries to an external *SAT Modulo Theories (SMT)* solver [Nieuwenhuis et al. 2006], as a black box. Future advances in the SMT world then directly transfer to analysis of LCTRSs.

Structure. We first recall the LCTRS formalism from [Kop and Nishida 2013] (§ 2) and show a way to translate procedural programs to LCTRSs (§ 3). Then we lift rewriting induction methods for constrained rewriting to LCTRSs (§ 4) and strengthen them with two dedicated generalization techniques (§ 5). Finally we discuss automation and experimental results (§ 6) as well as related and future work (§§ 7–8) and conclude.

Contributions over the conference version. The present paper provides several additional contributions over the conference version [Kop and Nishida 2014]: (1) We significantly extend our method to translate procedural programs to LCTRSs. (2) We extend our theory of constrained inductive theorem proving to *disproving* equivalence (following [Sakata et al. 2009; Falke and Kapur 2012]) and add several inference rules. (3) We provide an additional generalization technique and a detailed proof strategy to automate rewriting induction for translated procedural programs. (4) We have improved the implementation and added an automatic translation from C programs to LCTRSs.

1.1. Motivating Example

Aside from business applications, automatic equivalence proving can be used as an aid in grading student programming assignments. Combining a test run of the assignments on a set of sample inputs (which identifies many incorrect programs, but leaves false positives) with an automatic correctness check can save teachers a lot of time.

Example 1.1. Consider the following programming assignment.

Write a function sum which, given an integer array and its length as input, returns the sum of its elements. Do not modify the input array.

We consider four different C implementations of this exercise:

<pre>int sum1(int arr[],int n){ int ret=0; for(int i=0;i<n;i++) ret+=arr[i]; return ret; }</pre>	<pre>int sum2(int arr[], int n) { int ret, i; for (i = 0; i < n; i++) { ret += arr[i]; } return ret; }</pre>
<pre>int sum3(int arr[], int len) { int i; for (i = 0; i < len-1; i++) arr[i+1] += arr[i]; return arr[len-1]; }</pre>	<pre>int sum4(int *arr, int k) { if (k <= 0) return 0; return arr[k-1] + sum4(arr, k-1); }</pre>

The first solution is correct. The second is not, because `ret` is not initialized—which may be missed in standard tests depending on the compiler used. The third solution is incorrect because the array is modified against the instructions, and moreover, gives a random result or segmentation fault if `len = 0`. The fourth solution is correct.

These implementations can be transformed into the following LCTRSs:

- (1a) $\text{sum1}(arr, n) \rightarrow u(arr, n, 0, 0)$
- (1b) $u(arr, n, ret, i) \rightarrow \text{error} \quad [i < n \wedge (i < 0 \vee i \geq \text{size}(arr))]$
- (1c) $u(arr, n, ret, i) \rightarrow u(arr, n, ret + \text{select}(arr, i), i + 1) \quad [i < n \wedge 0 \leq i < \text{size}(arr)]$
- (1d) $u(arr, n, ret, i) \rightarrow \text{return}(arr, ret) \quad [i \geq n]$
- (2a) $\text{sum2}(arr, n) \rightarrow u(arr, n, ret, 0)$
u rules as copied from above
- (3a) $\text{sum3}(arr, len) \rightarrow v(arr, len, 0)$
- (3b) $v(arr, len, i) \rightarrow \text{error} \quad [i < len - 1 \wedge (i < 0 \vee i + 1 \geq \text{size}(arr))]$
- (3c) $v(arr, len, i) \rightarrow v(\text{store}(arr, i + 1, \text{select}(arr, i + 1) + \text{select}(arr, i)), len, i + 1)$
 $[i < len - 1 \wedge 0 \leq i \wedge i + 1 < \text{size}(arr)]$

- (3d) $v(arr, len, i) \rightarrow \text{return}(arr, \text{select}(arr, len - 1))$
 $[i \geq len - 1 \wedge 0 \leq len - 1 < \text{size}(arr)]$
- (3e) $v(arr, len, i) \rightarrow \text{error}$ $[i \geq len - 1 \wedge (len - 1 < 0 \vee len - 1 \geq \text{size}(arr))]$
- (4a) $\text{sum4}(arr, k) \rightarrow \text{return}(arr, 0)$ $[k \leq 0]$
- (4b) $\text{sum4}(arr, k) \rightarrow \text{error}$ $[k - 1 \geq \text{size}(arr)]$
- (4c) $\text{sum4}(arr, k) \rightarrow w(\text{select}(arr, k - 1), \text{sum4}(arr, k - 1))$ $[0 \leq k - 1 < \text{size}(arr)]$
- (4d) $w(n, \text{error}) \rightarrow \text{error}$
- (4e) $w(n, \text{return}(a, r)) \rightarrow \text{return}(a, n + r)$

Note that arrays carry an implicit size (their allocated memory) which is queried to model the runtime behavior of the C program and test for out-of-bound errors. The fresh variable in the right-hand side of (2a) models that the third parameter of v is assigned an *arbitrary* integer. The details of this transformation are discussed in § 3.

Using inductive theorem proving, we can now prove that

- $\forall arr \in \text{array}(\text{int}). \forall len \in \text{int}. \text{sum1}(arr, len) \leftrightarrow^* \text{sum4}(arr, len)$ if $0 \leq len \leq \text{size}(arr)$
- $\exists arr \in \text{array}(\text{int}). \exists len \in \text{int}. \text{sum3}(arr, len) \not\leftrightarrow^* \text{sum4}(arr, len)$ with $0 \leq len \leq \text{size}(arr)$

So sum1 and sum4 return the same result on any input such that the given length does not cause out-of-bound errors, but sum3 and sum4 do not. (It seems likely that the disproof obtained from inductive theorem proving could be used to extract counterexample inputs, but at present we have not studied a systematic way of doing so.)

For sum2 , we *do* have $\text{sum2}(arr, len) \leftrightarrow^* \text{sum4}(arr, len)$, since we *can* always choose to instantiate ret with 0. The system is not *confluent*; we can also prove that there exist a, n such that $\text{sum2}(a, n) \rightarrow^* s \neq t \leftarrow^* \text{sum4}(a, n)$ for terms s, t in normal form. As explained in § 6, we use a proof strategy which typically proves only the “ \neq ” statement.

1.2. Practical Use

The primary application that we see for our technique is the following:

1.2.1. Comparing a function to a specification. As in Ex. 1.1, we can verify correctness of a C function f against a reference implementation g by translating both functions to LCTRS rules (§ 3) and proving that $f(x_1, \dots, x_n) \approx g(x_1, \dots, x_n)$ [true] is an inductive theorem. If we only need equivalence under given preconditions on the input variables—such as $0 \leq len \leq \text{size}(arr)$ in Ex. 1.1—we formulate this as a constraint φ and analyze whether $f(x_1, \dots, x_n) \approx g(x_1, \dots, x_n)$ [φ] is an inductive theorem.

Note that we do not require a separate specification language—although if desirable, it is of course possible to specify the reference implementation directly as an LCTRS.

Further possible applications of our technique include:

1.2.2. Code optimization (or other improvement). Sometimes the “reference implementation” g suggested above can simply be an existing—and inefficient, or inelegant—version of a function. Thus, inductive theorem proving can be used to prove that it is safe to replace a function in a large real-life program by an optimized alternative.

1.2.3. Error checking. As the transformation from C to LCTRSs includes error checking (as seen for memory safety violations in Ex. 1.1), we can use inductive theorem proving to verify the absence of such errors. This is done by adding error-checking rules, e.g.,

$$\text{errorfree}(\text{return}(a, n)) \rightarrow \text{true} \qquad \text{errorfree}(\text{error}) \rightarrow \text{false}$$

and proving that $\text{errorfree}(\text{sum4}(a, n)) \approx \text{true}$ [φ] is an inductive theorem, where φ is the precondition on the input. Aside from memory safety, this approach can be used to certify the absence of for instance divisions by zero or integer overflow. The key is in the transformation, where we can choose which constructions result in an error.

1.2.4. *Classical correctness checks.* Aside from comparisons to an example implementation, we can also specify a correctness property directly in SMT. For instance, given an implementation of the `strlen` function, its correctness could be verified by proving that $\text{strlen}(x) \approx \text{return}(n) [0 \leq n < \text{size}(x) \wedge \text{select}(x, n) = 0 \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0)]$ is an inductive theorem. Alternatively, we can use extra rules to test properties in SMT.

Example 1.2. To analyze correctness of an implementation of `strcpy`, we may use

$$\begin{aligned} \text{test}(x, n, \text{error}) &\rightarrow \text{false} \\ \text{test}(x, n, \text{return}(y)) &\rightarrow b [b \Leftrightarrow \forall i \in \{0, \dots, n\}(\text{select}(x, i) = \text{select}(y, i))] \end{aligned}$$

and prove that the following equation is an inductive theorem:

$$[0 \leq n < \text{size}(x) \wedge n < \text{size}(y) \wedge \text{select}(x, n) = 0 \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0)] \text{test}(x, n, \text{strcpy}(y, x)) \approx \text{true}$$

Note that this more sophisticated test is *needed* in this case, since correctness of `strcpy` does not require that $x = y$ if $\text{strcpy}(x) \rightarrow^* \text{return}(y)$ (the sizes of x and y may differ).

2. PRELIMINARIES

In this section, we briefly recall *Logically Constrained Term Rewriting Systems* (usually abbreviated as *LCTRSs*), following the definitions in [Kop and Nishida 2013].

2.1. Logically Constrained Term Rewriting Systems

Many-sorted terms. We introduce terms, typing, substitutions, contexts, and subterms (with corresponding terminology) in the usual way for many-sorted term rewriting.

Definition 2.1. We assume given a set S of *sorts* and an infinite set \mathcal{V} of *variables*, each variable equipped with a sort. A *signature* Σ is a set of *function symbols* f , disjoint from \mathcal{V} , each equipped with a *sort declaration* $[\iota_1 \times \dots \times \iota_n] \Rightarrow \kappa$, with all ι_i and κ sorts. For readability, we often write κ instead of $[\] \Rightarrow \kappa$. The set $\text{Terms}(\Sigma, \mathcal{V})$ of *terms* over Σ and \mathcal{V} contains any expression s such that $\vdash s : \iota$ can be derived for some sort ι , using:

$$\frac{}{\vdash x : \iota} (x : \iota \in \mathcal{V}) \quad \frac{\vdash s_1 : \iota_1 \quad \dots \quad \vdash s_n : \iota_n}{\vdash f(s_1, \dots, s_n) : \kappa} (f : [\iota_1 \times \dots \times \iota_n] \Rightarrow \kappa \in \Sigma)$$

We fix Σ and \mathcal{V} . Note that for every term s , there is a unique sort ι with $\vdash s : \iota$.

Definition 2.2. Let $\vdash s : \iota$. We call ι the *sort* of s . Let $\text{Var}(s)$ be the set of variables occurring in s ; we say that s is *ground* if $\text{Var}(s) = \emptyset$.

Definition 2.3. A *substitution* γ is a sort-preserving total mapping from \mathcal{V} to $\text{Terms}(\Sigma, \mathcal{V})$. The result $s\gamma$ of applying a substitution γ to a term s is s with all occurrences of a variable x replaced by $\gamma(x)$. The *domain* of γ , $\text{Dom}(\gamma)$, is the set of variables x with $\gamma(x) \neq x$. The notation $[x_1 := s_1, \dots, x_k := s_k]$ denotes a substitution γ with $\gamma(x_i) = s_i$ for $1 \leq i \leq k$, and $\gamma(y) = y$ for $y \notin \{x_1, \dots, x_k\}$. For two substitutions γ and δ , their composition $\gamma \circ \delta$ is given by $(\gamma \circ \delta)(x) = \gamma(\delta(x)) = (x\delta)\gamma$ for all variables x .

Two terms s and t are *unifiable* if there exists a substitution γ such that $s\gamma = t\gamma$. Then γ is called a *unifier* for s and t . If moreover for all unifiers γ' for s and t there is a substitution δ such that $\gamma' = \delta \circ \gamma$, we call γ a *most general unifier (mgu)* for s and t .

Definition 2.4. Given a term s , a *position* in s is a sequence p of positive integers such that $s|_p$ is defined, where $s|_\epsilon = s$ and $f(s_1, \dots, s_n)|_{i.p} = (s_i)|_p$. We call $s|_p$ a *subterm* of s . If $\vdash s|_p : \iota$ and $\vdash t : \iota$, then $s[t]_p$ denotes s with the subterm at position p replaced by t . A *context* C is a term containing one or more typed *holes* $\square_i : \iota_i$. If $s_1 : \iota_1, \dots, s_n : \iota_n$, we define $C[s_1, \dots, s_n]$ as C with each \square_i replaced by s_i .

Logical terms. Specific to LCTRSs, we consider different kinds of symbols and terms.

Definition 2.5. We assume given:

- signatures Σ_{terms} and Σ_{theory} such that $\Sigma = \Sigma_{terms} \cup \Sigma_{theory}$;
- a mapping \mathcal{I} which assigns to each sort ι occurring in Σ_{theory} a set \mathcal{I}_ι ;
- a mapping \mathcal{J} which assigns to each $f : [\iota_1 \times \dots \times \iota_n] \Rightarrow \kappa \in \Sigma_{theory}$ a function in $\mathcal{I}_{\iota_1} \times \dots \times \mathcal{I}_{\iota_n} \Rightarrow \mathcal{I}_\kappa$;
- for all sorts ι occurring in Σ_{theory} a set $\mathcal{V}al_\iota \subseteq \Sigma_{theory}$ of *values*: function symbols $a : [] \Rightarrow \iota$ such that \mathcal{J} gives a bijective mapping from $\mathcal{V}al_\iota$ to \mathcal{I}_ι .

We require that $\Sigma_{terms} \cap \Sigma_{theory} \subseteq \mathcal{V}al = \bigcup_\iota \mathcal{V}al_\iota$. The sorts occurring in Σ_{theory} are called *theory sorts*, and the symbols *theory symbols*. Symbols in $\Sigma_{theory} \setminus \mathcal{V}al$ are *calculation symbols*. A term in $\mathcal{T}erms(\Sigma_{theory}, \mathcal{V})$ is called a *logical term*.

Definition 2.6. For ground logical terms, let $\llbracket f(s_1, \dots, s_n) \rrbracket := \mathcal{J}_f(\llbracket s_1 \rrbracket, \dots, \llbracket s_n \rrbracket)$. For every ground logical term s there is a unique value c such that $\llbracket s \rrbracket = \llbracket c \rrbracket$; we say that c is the value of s . A *constraint* is a logical term φ of some sort bool with $\mathcal{I}_{\text{bool}} = \mathbb{B} = \{\top, \perp\}$, the set of *booleans*. A constraint φ is *valid* if $\llbracket \varphi \gamma \rrbracket = \top$ for *all* substitutions γ which map $\text{Var}(\varphi)$ to values, and *satisfiable* if $\llbracket \varphi \gamma \rrbracket = \top$ for *some* such substitutions. A substitution γ *respects* φ if $\gamma(x)$ is a value for all $x \in \text{Var}(\varphi)$ and $\llbracket \varphi \gamma \rrbracket = \top$.

Terms in $\mathcal{T}erms(\Sigma_{terms}, \emptyset)$ can be thought of as the primary objects of rewriting: a reduction typically begins and ends with such terms, with elements of $\Sigma_{theory} \setminus \mathcal{V}al$ (also called *calculation symbols*) to perform calculations in the underlying theory.

We typically choose a theory signature with $\Sigma_{theory} \supseteq \Sigma_{theory}^{core}$, where Σ_{theory}^{core} contains $\text{true}, \text{false} : \text{bool}, \wedge, \vee, \Rightarrow : [\text{bool} \times \text{bool}] \Rightarrow \text{bool}, \neg : [\text{bool}] \Rightarrow \text{bool}$, and, for all theory sorts ι , symbols $=_\iota, \neq_\iota : [\iota \times \iota] \Rightarrow \text{bool}$, and an evaluation function \mathcal{J} that interprets these symbols as expected. We omit the sort subscripts from $=$ and \neq when clear from context.

Definition 2.7. The standard integer signature Σ_{theory}^{int} is $\Sigma_{theory}^{core} \cup \{+, -, *, \text{exp}, \text{div}, \text{mod} : [\text{int} \times \text{int}] \Rightarrow \text{int}; \leq, < : [\text{int} \times \text{int}] \Rightarrow \text{bool}\} \cup \{n : \text{int} \mid n \in \mathbb{Z}\}$ with values $\text{true}, \text{false}$ and n for all $n \in \mathbb{Z}$. Thus, we use n (in sans-serif font) as the function symbol for $n \in \mathbb{Z}$ (in *math* font). We define \mathcal{J} in the natural way, except: since all \mathcal{J}_f must be total functions, we set $\mathcal{J}_{\text{div}}(n, 0) = \mathcal{J}_{\text{mod}}(n, 0) = \mathcal{J}_{\text{exp}}(n, k) = 0$ for all n and all $k < 0$. Of course, when constructing LCTRSs, we normally add explicit error checks to prevent such calls.

Example 2.8. Let $\mathcal{S} = \{\text{int}, \text{bool}\}$, and $\Sigma = \Sigma_{terms} \cup \Sigma_{theory}^{int}$, where

$$\Sigma_{terms} = \{ \text{fact} : [\text{int}] \Rightarrow \text{int} \} \cup \{ n : \text{int} \mid n \in \mathbb{Z} \}$$

Then both int and bool are theory sorts. We also define set and function interpretations, i.e., $\mathcal{I}_{\text{int}} = \mathbb{Z}, \mathcal{I}_{\text{bool}} = \mathbb{B}$, and \mathcal{J} is defined as above. With $=$ for $=_{\text{int}}$ and infix notation, examples of logical terms are $0 = 0 + -1$ and $x + 3 \geq y + -42$. Both are constraints. $5 + 9$ is also a (ground) logical term, but not a constraint. Expected starting terms are, e.g., $\text{fact}(42)$ or $\text{fact}(\text{fact}(-4))$: ground terms fully built using symbols in Σ_{terms} .

Rules and rewriting. We adapt the standard notions of rewriting (see, e.g., [Baader and Nipkow 1998]) by including constraints and adding rules to perform calculations.

Definition 2.9. A *rule* is a triple $\ell \rightarrow r [\varphi]$ with ℓ and r terms of the same sort and φ a constraint. Here, ℓ has the form $f(\ell_1, \dots, \ell_n)$ and contains at least one symbol in $\Sigma_{terms} \setminus \Sigma_{theory}$ (so ℓ is not a logical term). If $\varphi = \text{true}$ with $\mathcal{J}(\text{true}) = \top$, we may write $\ell \rightarrow r$. We define $L\text{Var}(\ell \rightarrow r [\varphi])$ as $\text{Var}(\varphi) \cup (\text{Var}(r) \setminus \text{Var}(\ell))$. A substitution γ *respects* $\ell \rightarrow r [\varphi]$ if $\gamma(x) \in \mathcal{V}al$ for all $x \in L\text{Var}(\ell \rightarrow r [\varphi])$, and $\llbracket \varphi \gamma \rrbracket = \top$. The rule is *left-linear* if ℓ is linear, i.e., all variables occur at most once in ℓ , and *irregular* if $\text{Var}(\varphi) \setminus \text{Var}(\ell) \neq \emptyset$.

Note that it is allowed to have $Var(r) \not\subseteq Var(\ell)$, but fresh variables in the right-hand side may only be instantiated with *values*. This is done to model user input or random choice. Otherwise, variables outside the constraint may be instantiated by any term; we do not impose strategies like innermost or call-by-value reduction.

Definition 2.10. We assume given a set of rules \mathcal{R} and let $\mathcal{R}_{\text{calc}}$ be the set $\{f(x_1, \dots, x_n) \rightarrow y \mid [y = f(\vec{x})] \mid f : [\iota_1 \times \dots \times \iota_n] \Rightarrow \kappa \in \Sigma_{\text{theory}} \setminus \mathcal{V}al\}$ (writing \vec{x} for x_1, \dots, x_n). The *rewrite relation* $\rightarrow_{\mathcal{R}}$ is a binary relation on terms, defined by:

$$C[\ell\gamma] \rightarrow_{\mathcal{R}} C[r\gamma] \quad \text{if } \ell \rightarrow r \text{ } [\varphi] \in \mathcal{R} \cup \mathcal{R}_{\text{calc}} \text{ and } \gamma \text{ respects } \ell \rightarrow r \text{ } [\varphi]$$

Here, C is a context with exactly one hole. We say that the reduction occurs at position p if $C = C[\square]_p$. Let $s \leftrightarrow_{\mathcal{R}} t$ if $s \rightarrow_{\mathcal{R}} t$ or $t \rightarrow_{\mathcal{R}} s$. A reduction step with $\mathcal{R}_{\text{calc}}$ is called a *calculation*. A term is in *normal form* if it cannot be reduced with $\rightarrow_{\mathcal{R}}$. We say that t is a *normal form of* s if $s \rightarrow_{\mathcal{R}}^* t$ and t is a normal form. The relation $\rightarrow_{\mathcal{R}}$ is *confluent* if whenever $s \rightarrow_{\mathcal{R}}^* t$ and $s \rightarrow_{\mathcal{R}}^* t'$, there exists also some u with $t \rightarrow_{\mathcal{R}}^* u$ and $t' \rightarrow_{\mathcal{R}}^* u$.

We usually call the elements of $\mathcal{R}_{\text{calc}}$ rules—or *calculation rules*—even though their left-hand side is a logical term. Note that if $\rightarrow_{\mathcal{R}}$ is confluent, every term has at most one normal form (intuitively, then \mathcal{R} is deterministic with respect to big-step semantics).

Definition 2.11. For $f(\ell_1, \dots, \ell_n) \rightarrow r$ $[\varphi] \in \mathcal{R}$ we call f a *defined symbol*; non-defined elements of Σ_{terms} and all values are *constructors*. Let \mathcal{D} be the set of all defined symbols and $\mathcal{C}ons$ the set of constructors. A term in $\mathcal{T}erms(\mathcal{C}ons, \mathcal{V})$ is a *constructor term*.

Now we may define a *logically constrained term rewriting system* (LCTRS) as the abstract rewriting system $(\mathcal{T}erms(\Sigma, \mathcal{V}), \rightarrow_{\mathcal{R}})$. An LCTRS is usually given by supplying Σ , \mathcal{R} , and an informal description of \mathcal{I} and \mathcal{J} if these are not clear from context.

Example 2.12. To implement an LCTRS calculating the *factorial* function, we use the signature Σ from Ex. 2.8 and the following rules:

$$\mathcal{R}_{\text{fact}} = \{ \text{fact}(x) \rightarrow 1 \text{ } [x \leq 0] \text{ , } \text{fact}(x) \rightarrow x * \text{fact}(x - 1) \text{ } [\neg(x \leq 0)] \}$$

Using calculation steps, a term $3 - 1$ reduces to 2 in one step (using the calculation rule $x - y \rightarrow z$ $[z = x - y]$), and $3 * (2 * (1 * 1))$ reduces to 6 in three steps. Using also the rules in $\mathcal{R}_{\text{fact}}$, $\text{fact}(3)$ reduces in ten steps to 6.

Example 2.13. To implement an LCTRS calculating the sum of elements in an array, let $\mathcal{I}_{\text{bool}} = \mathbb{B}$, $\mathcal{I}_{\text{int}} = \mathbb{Z}$, $\mathcal{I}_{\text{array}(\text{int})} = \mathbb{Z}^*$, so $\text{array}(\text{int})$ is mapped to finite-length integer sequences. Let $\Sigma_{\text{theory}} = \Sigma_{\text{theory}}^{\text{int}} \cup \{\text{size} : [\text{array}(\text{int})] \Rightarrow \text{int}, \text{select} : [\text{array}(\text{int}) \times \text{int}] \Rightarrow \text{int}\} \cup \{a \mid a \in \mathbb{Z}^*\}$. (We do *not* encode arrays as lists: every “array”—integer sequence— a corresponds to a unique symbol a .) The interpretation function \mathcal{J} behaves on $\Sigma_{\text{theory}}^{\text{int}}$ as usual, maps the values a to the corresponding integer sequence, and has:

$$\mathcal{J}_{\text{size}}(a) = k \text{ if } a = \langle n_0, \dots, n_{k-1} \rangle \quad \mathcal{J}_{\text{select}}(a, i) = n_i \text{ if } a = \langle n_0, \dots, n_{k-1} \rangle \text{ and } 0 \leq i < k \\ 0 \text{ otherwise}$$

In addition, let $\Sigma_{\text{terms}} = \{\text{sum}, \text{sum0} : [\text{array}(\text{int})] \Rightarrow \text{int}\} \cup \{n : \text{int} \mid n \in \mathbb{Z}\} \cup \{a \mid a \in \mathbb{Z}^*\}$ and let \mathcal{R} consist of

$$\text{sum}(x) \rightarrow \text{sum0}(x, \text{size}(x) - 1) \quad \text{sum0}(x, k) \rightarrow \text{select}(x, k) + \text{sum0}(x, k - 1) \begin{array}{l} [k \geq 0] \\ [k < 0] \end{array} \\ \text{sum0}(x, k) \rightarrow 0$$

Note that this implementation differs from the ones in Ex. 1.1, because there we analyzed encodings of imperative programs; on C level there is no functionality for the programmer to explicitly query the size of an array. Here, we avoided boundary checks.

Values are new in LCTRSs compared to older styles of constrained rewriting. These representatives of the underlying theory are always *constants* (constructor symbols which do not take arguments), even if they represent complex structures, as seen in Ex. 2.13. Note that variables in a rule's constraint *must* be instantiated by values; for instance in Ex. 2.12, a term $\text{fact}(1 + 2)$ must be reduced by a calculation first. We also do not match modulo theories, e.g., we do not equate $0 + (x + y)$ with $y + x$ for matching.

Differences to [Kop and Nishida 2013]. In the original definition of LCTRSs, variables in \mathcal{V} are unsorted, and a separate *variable environment* is used for typing. Also, $\rightarrow_{\mathcal{R}}$ is there defined as the union of two relations $\rightarrow_{\text{rule}}$ and $\rightarrow_{\text{calc}}$ rather than including $\mathcal{R}_{\text{calc}}$. These changes give equivalent results, but the current definitions cause less bookkeeping. A larger difference is the restriction on rules: in [Kop and Nishida 2013] left-hand sides must have a root symbol in $\Sigma_{\text{terms}} \setminus \Sigma_{\text{theory}}$. We follow Kop [2013] and Kop and Nishida [2014] in weakening this (only asking that they are not logical terms).

2.2. Quantification

The definition of LCTRSs does not permit constraints with quantifiers (constraints are terms, and first-order rewriting does not allow quantifiers in terms). In, for instance, an LCTRS over integers and arrays, which has $\text{addtoend} : [\text{int} \times \text{array}(\text{int})] \Rightarrow \text{array}(\text{int}) \in \Sigma_{\text{theory}}$ and $\text{extend} : [\text{array}(\text{int}) \times \text{int}] \Rightarrow \text{array}(\text{int}) \in \Sigma_{\text{terms}}$, we cannot specify a rule like:

$$\text{extend}(\text{arr}, x) \rightarrow \text{addtoend}(x, \text{arr}) \ [\forall y \in \{0, \dots, \text{size}(\text{arr}) - 1\} (x \neq \text{select}(\text{arr}, y))]$$

However, one of the key features of LCTRSs is that theory symbols, including predicates, are not confined to a fixed list. Therefore, we *can* add a new symbol to Σ_{theory} (and \mathcal{J}). For the extend rule, we might introduce a symbol $\text{notin} : [\text{int} \times \text{array}(\text{int})] \Rightarrow \text{bool}$ with $\mathcal{J}_{\text{notin}}(u, \langle a_0, \dots, a_{n-1} \rangle) = \top$ iff for all i : $u \neq a_i$, and replace the constraint by $\text{notin}(x, \text{arr})$. This generates exactly the same reduction relation as the original rule.

Thus, we can permit quantifiers in the constraints of rules and also on right-hand sides of rules, as an intuitive notation for fresh predicates. However, an *unbounded* quantification would likely not be useful, as it would give an undecidable relation $\rightarrow_{\mathcal{R}}$.

Comment: One might argue that adding symbols like this is problematic in practice: no SMT solver will support new symbols like notin . However, for the *technique* this makes no difference. In an implementation, we might allow quantifiers as syntactic sugar (and pass the same sugar to the SMT solver), or add a layer on top of the SMT solver which translates the new symbol(s), replacing for instance $(\text{notin } u \ a)$ by $(\text{forall } ((x \ \text{Int})) (\text{distinct } u \ (\text{select } a \ x)))$.

2.3. Rewriting Constrained Terms

In LCTRSs, the objects of study are *terms*, with $\rightarrow_{\mathcal{R}}$ defining the relation between them. However, for analysis it is often useful to consider *constrained terms*:

Definition 2.14. A constrained term is a pair $s[\varphi]$ of a term s and a constraint φ . We say $s[\varphi]$ and $t[\psi]$ are *equivalent*, notation $s[\varphi] \sim t[\psi]$, if for all substitutions γ which respect φ there is a substitution δ which respects ψ such that $s\gamma = t\delta$, and vice versa.

Intuitively, a constrained term $s[\varphi]$ represents all terms $s\gamma$ where γ respects φ , and can be used to reason about such terms. Equivalent constrained terms represent the same set of terms; for example $f(0) [\text{true}] \sim f(x) [x = 0]$, and $g(x, y) [x > y] \sim g(z, u) [u \leq z - 1]$. Note that $s[\varphi] \sim s[\psi]$ if and only if $\forall \vec{x} (\exists \vec{y} (\varphi) \leftrightarrow \exists \vec{z} (\psi))$ holds, where $\text{Var}(s) = \{\vec{x}\}$, $\text{Var}(\varphi) \setminus \text{Var}(s) = \{\vec{y}\}$ and $\text{Var}(\psi) \setminus \text{Var}(s) = \{\vec{z}\}$.

Definition 2.15. For a rule $\rho := \ell \rightarrow r[\psi] \in \mathcal{R} \cup \mathcal{R}_{\text{calc}}$ and position q , we let $s[\varphi] \rightarrow_{\rho, q} t[\psi]$ if there exists a substitution γ such that $s|_q = \ell\gamma$, $t = s[r\gamma]_q$, $\gamma(x)$ is a value or

variable in $Var(\varphi)$ for all $x \in LVar(\ell \rightarrow r [\psi])$, and $\varphi \Rightarrow (\psi\gamma)$ is valid. Let $s [\varphi] \rightarrow_{\text{base}} t [\varphi]$ if $s [\varphi] \rightarrow_{\rho,q} t [\varphi]$ for some ρ, q . The relation $\rightarrow_{\mathcal{R}}$ on constrained terms is defined as $\sim \cdot \rightarrow_{\text{base}} \cdot \sim$. We say that $s [\varphi] \rightarrow_{\mathcal{R}} t [\psi]$ at position q by rule ρ if $s [\varphi] \sim \cdot \rightarrow_{\rho,q} \cdot \sim t [\psi]$.

Example 2.16. In the LCTRS from Ex. 2.12, we have $\text{fact}(x) [x > 3] \rightarrow_{\mathcal{R}} x * \text{fact}(x - 1) [x > 3]$. Now we can use a calculation rule $x - y \rightarrow z [z = x - y]$, with a non-empty \sim -step, as follows: $x * \text{fact}(x - 1) [x > 3] \sim x * \text{fact}(x - 1) [x > 3 \wedge z = x - 1] \rightarrow_{\text{base}} x * \text{fact}(z) [x > 3 \wedge z = x - 1]$. The \sim -relation holds because indeed $\forall x(x > 3 \leftrightarrow \exists z(x > 3 \wedge z = x - 1))$.

Example 2.17. The \sim -relation also allows us to reformulate the constraint after a reduction. For example, with the rule $f(x) \rightarrow g(y) [y > x]$, we have: $f(x) [x > 3] \sim f(x) [x > 3 \wedge y > x] \rightarrow_{\text{base}} g(y) [x > 3 \wedge y > x] \sim g(y) [y > 4]$. We do *not* have that $f(x) [\text{true}] \rightarrow_{\mathcal{R}} g(x + 1) [\text{true}]$, as $x + 1$ cannot be instantiated to a value.

Example 2.18. A constrained term does not always need to be reduced in the most general way. With the rule $f(x) \rightarrow g(y) [y > x]$, we have $f(0) [\text{true}] \sim f(0) [y > 0] \rightarrow_{\text{base}} g(y) [y > 0]$, but we also have $f(0) [\text{true}] \sim f(0) [1 > 0] \rightarrow_{\text{base}} g(1) [1 > 0] \sim g(1) [\text{true}]$.

As intended, constrained reductions give information about usual reductions:

THEOREM 2.19. *If $s [\varphi] \rightarrow_{\mathcal{R}} t [\psi]$, then for all substitutions γ which respect φ there exists δ which respects ψ such that $s\gamma \rightarrow_{\mathcal{R}} t\delta$. Both steps use the same rule and position.*

PROOF. We first observe (**): *If $u [\xi] \rightarrow_{\text{base}} q [\xi]$, then for any substitution γ which respects ξ also $u\gamma \rightarrow_{\mathcal{R}} q\gamma$.* Proof: if $u [\xi] \rightarrow_{\text{base}} q [\xi]$, then there are $p, \ell \rightarrow r [c]$ and δ such that $u|_p = \ell\delta$, $q = u[r\delta]_p$, $\delta(x) \in Var(\xi) \cup Val$ for all $x \in LVar(\ell \rightarrow r [c])$ and $\xi \Rightarrow (c\delta)$ is valid. With $\eta = \gamma \circ \delta$, we have $(u\gamma)|_p = u|_p\gamma = \ell\delta\gamma = \ell\eta$ and $q\gamma = u[r\delta]_p\gamma = (u\gamma)[r\delta\gamma]_p = (u\gamma)[r\eta]_p$. We also have $\eta(x) = \delta(x)\gamma \in Val$ for $x \in LVar(\ell \rightarrow r [c])$ because γ respects ξ and, since $\llbracket \xi\gamma \rrbracket = \top$ and $\xi \Rightarrow (c\delta)$ is valid, also $\llbracket (c\delta)\gamma \rrbracket = \llbracket c\eta \rrbracket = \top$. So indeed $u\gamma \rightarrow_{\mathcal{R}} q\gamma$.

Now, suppose $s [\varphi] \rightarrow_{\mathcal{R}} t [\psi]$, so $s [\varphi] \sim s' [\xi] \rightarrow_{\text{base}} t' [\xi] \sim t [\psi]$, and let γ respect φ . By definition of \sim , there is some substitution η which respects ξ such that $s\gamma = s'\eta$. By (**)
 $s'\eta \rightarrow_{\mathcal{R}} t'\eta$. Again by definition of \sim , we find δ which respects ψ such that $t'\eta = t\delta$. \square

THEOREM 2.20. *If $s [\varphi] \rightarrow_{\mathcal{R}} t [\psi]$, then for all substitutions δ which respect ψ there exists γ which respects φ such that $s\gamma \rightarrow_{\mathcal{R}} t\delta$. Both steps use the same rule and position.*

PROOF. Parallel to the proof of Thm. 2.19: if $s [\varphi] \sim s' [\xi] \rightarrow_{\text{base}} t' [\xi] \sim t [\psi]$, then by definition of \sim there are suitable η, γ such that $t\delta = t'\eta \leftarrow_{\mathcal{R}} s'\eta = s\gamma$. \square

Comment: The relation $\rightarrow_{\mathcal{R}}$ on constrained terms is *not* stable: for instance, in the system from Ex. 2.18, we can derive $f(x) [\text{true}] \rightarrow_{\mathcal{R}} g(x) [\text{true}]$ even though $f(0) [\text{true}] \not\rightarrow_{\mathcal{R}} g(0) [\text{true}]$. This is because the variables in a constrained term $s [\varphi]$ are fully changeable; one can see variables in $Var(s)$ as universal and the others as existential. This is not problematic, as we do not instantiate constrained terms; to reason with constrained reduction we only use Theorems 2.19 and 2.20.

3. TRANSFORMING IMPERATIVE PROGRAMS INTO LCTRSS

Equivalence-preserving transformations of imperative programs into constrained rewriting systems operating on integers have been investigated in e.g. [Falke and Kapur 2009; Falke et al. 2011; Furuichi et al. 2008]; more generally, such translations from imperative to functional programs have been investigated at least since [McCarthy 1960]. Although these papers use different definitions of constrained rewriting, the proposed transformations can be adapted to produce LCTRSs that operate on integers, i.e., use Σ_{theory} as in Ex. 2.12. What is more, we can extend the ideas to also handle more advanced programming structures, such as arrays and exceptions.

In this section, we will discuss a number of ideas towards a translation from C to LCTRS. A more detailed and formal treatment of the limitation to integers and one-dimensional integer arrays is available online along with an implementation, at:

<http://www.trs.css.i.nagoya-u.ac.jp/c2lctrs/>

Given the extensiveness of the C specification, we will not attempt to prove that the result of our transformation corresponds to the origin. Instead, we shall rely on an appeal to intuition. An advantage is that the same ideas apply to other programming languages; we should be able to use similar translations for, e.g., Python or Java.

3.1. Transforming Simple Integer Functions

The base form of the transformation—limited to integer functions with no global variables or function calls—is very similar to the transformations for integer TRSs in [Falke and Kapur 2009; Falke et al. 2011; Furuichi et al. 2008]. Each function is transformed separately. We introduce a function symbol for every statement (including declarations), which operates on the variables in scope. The transition from one statement to another is encoded as a rule, with assignments reflected by argument updates in the right-hand side, and conditions by the constraint. Return statements are encoded by reducing to an expression $\text{return}_f(e)$, where $\text{return}_f : [\text{int}] \Rightarrow \text{result}_f$ is a constructor.

Example 3.1. Consider the following C function and its translation:

<pre>int fact(int x) { int z = 1; for (int i = 1; i <= x; i++) z *= i; return z; }</pre>	$\begin{aligned} \text{fact}(x) &\rightarrow u_1(x, 1) \\ u_1(x, z) &\rightarrow u_2(x, z, 1) \\ u_2(x, z, i) &\rightarrow u_3(x, z, i) && [i \leq x] \\ u_2(x, z, i) &\rightarrow u_5(x, z) && [\neg(i \leq x)] \\ u_3(x, z, i) &\rightarrow u_4(x, z * i, i) \\ u_4(x, z, i) &\rightarrow u_2(x, z, i + 1) \\ u_5(x, z) &\rightarrow \text{return}_{\text{fact}}(z) \end{aligned}$
---	--

For Σ_{theory} we assume the standard integer signature; Σ_{terms} contains fact , all u_i and the constructor return_f , all of which have output sort result_f and argument sorts int .

A realistic translation of C code must also handle the absence of a boolean data type, operator precedence, and expressions with side effects (e.g., a loop condition $--x$). All this is easily doable¹ (and included in our implementation), but for the sake of brevity we will not go into detail here.

Finally, the generated system is optimized to make it more amenable to analysis:²

- rules are combined where possible, e.g., replacing a pair of rules $\ell \rightarrow u(r_1, \dots, r_n) [\varphi]$ and $u(x_1, \dots, x_n) \rightarrow s [\text{true}]$ by $\ell \rightarrow s[x_1 := r_1, \dots, x_n := r_n]$ if u is not used elsewhere;
- unused arguments of function symbols are removed, such as the second (but not the first!) argument of u in an LCTRS with rules $u(x, y, z) \rightarrow u(x - 1, y + 1, z * 2) [x > 0]$ and $u(x, y, z) \rightarrow \text{return}(z) [\neg(x > 0)]$;
- constraints are simplified, for instance replacing $\neg(x > 0)$ by $x \leq 0$ in the rules above.

We will use these optimizations also for the extended transformations of §§ 3.2–3.6.

Comment: When *time complexity*—defined as, e.g., the number of certain calculation steps—is considered, the argument removal step is dangerous, as it may remove calculations. In such cases we would use a different simplification method.

¹This is discussed in the formal treatment at <http://www.trs.css.i.nagoya-u.ac.jp/c2lctrs/formal.pdf>

²Variations of such preprocessing steps preserving the properties of interest to simplify the output of an automatic translation are fairly standard in program analysis, see e.g. [Albert et al. 2008; Alpuente et al. 2007; Beyer et al. 2009; Falke et al. 2011; Giesl et al. 2017; Spoto et al. 2009].

Example 3.2. Optimizing the LCTRS from Ex. 3.1, we obtain:

$$\begin{array}{ll} \text{fact}(x) \rightarrow u_2(x, 1, 1) & u_2(x, z, i) \rightarrow u_2(x, z * i, i + 1) \quad [i \leq x] \\ & u_2(x, z, i) \rightarrow \text{return}_{\text{fact}}(z) \quad [i > x] \end{array}$$

Differences to older work. In contrast to existing transformations to integer TRSs (e.g. [Falke and Kapur 2009; Falke et al. 2011; Furuichi et al. 2008]), we do not consider *basic blocks*, but simply create rules for every statement; this gives no substantial difference after optimization. Additionally, return_f is new here: in the work by Falke et al, the return statement is omitted, as they focus on *termination*, while in [Furuichi et al. 2008] the final term reduces directly to the return-value, e.g. $u_4(x, z) \rightarrow z + x \ [x \leq 0]$.

3.2. Non-Integer Data Types

Integers are not special: as the definition of LCTRSs permits arbitrary theories, we can handle any data type in C. We might for instance interpret `double` as either real numbers or double-precision floating point numbers; this choice is left to the user and may vary by application. The only requirement is that a suitable theory signature—with corresponding SMT solver if the system is to be analyzed automatically—is available. The translation is straightforward, with the only difficulty that type casts must be made explicit, and we need to use separate symbols such as `+` for double addition.

Example 3.3. Consider the following C function and its translation.

```
double halfsum(double thold) {
  double ret = 0.0;
  for (int d = 2; d < 100; d *= 2) {
    ret += 1.0 / d;
    if (ret > thold) return ret;
  }
}
```

$$\begin{array}{ll} \text{halfsum}(t) \rightarrow u_2(t, 0.0, 2) & \\ u_2(t, r, d) \rightarrow u_4(t, r + 1.0/\text{todouble}(d), d) \quad [d < 100] & \\ u_2(t, r, d) \rightarrow \text{return}_{\text{halfsum}}(\text{rnd}) \quad [d \geq 100] & \\ u_4(t, r, d) \rightarrow \text{return}_{\text{halfsum}}(r) \quad [r > .t] & \\ u_4(t, r, d) \rightarrow u_2(t, r, d * 2) \quad [r \leq .t] & \end{array}$$

This demonstrates both an explicit cast and one possible way to handle an undefined return value (by a fresh variable, which may be instantiated with a random value).

3.3. Error Handling

The transformation of § 3.1 does not fully reflect the original C program: as computers have limited memory, integers are internally represented as *bitvectors*. To address this, we could change the theory. Rather than using \mathbb{Z} , we let $\text{Val}_{\text{int}} = \{\text{MININT}, \dots, \text{MAXINT}\}$ and make \mathcal{J}_+ , \mathcal{J}_- , and \mathcal{J}_* wrap around (e.g., $\mathcal{J}_-(\text{MININT}, 1) = \text{MAXINT}$). The resulting LCTRS has the same rules, but acts more closely to the real program behavior.

However, integer overflow is often indicative of an *error*. Indeed, in C an overflow for the type `int` leads to undefined behavior (which also surfaces in optimizing compilers such as `gcc` or `clang`). In order to model this (or other instances of undefined behavior in C, such as a missing return statement), we will reduce to a special error state.

Thus, for every rule $u_i(x_1, \dots, x_n) \rightarrow r \ [\varphi]$: if this rule represents a transition where an error may occur under condition τ , then we split it in two:

$$u_i(x_1, \dots, x_n) \rightarrow r \ [\varphi \wedge \neg\tau] \quad u_i(x_1, \dots, x_n) \rightarrow \text{error}_f \ [\varphi \wedge \tau]$$

As usual, we simplify the resulting constraint (writing, e.g., $x < 0$ instead of $\neg(x \geq 0)$).

Example 3.4. Continuing Ex. 3.2, we generate the following rewrite rules:

$$\begin{array}{ll} \text{fact}(x) \rightarrow u_2(x, 1, 1) & \\ u_2(x, z, i) \rightarrow u_2(x, z * i, i + 1) \quad [i \leq x \wedge z * i \leq \text{MAXINT} \wedge z * i \geq \text{MININT} \wedge i + 1 \leq \text{MAXINT}] & \\ u_2(x, z, i) \rightarrow \text{error}_{\text{fact}} \quad [i \leq x \wedge (z * i > \text{MAXINT} \vee z * i < \text{MININT} \vee i + 1 > \text{MAXINT})] & \\ u_2(x, z, i) \rightarrow \text{return}_{\text{fact}}(z) \quad [i > x] & \end{array}$$

Note that we could easily model assertions and throw statements for exceptions in the same way. Division by zero is handled in a similar way.

We can choose whether to add error transitions before or after the simplification step. The distinction is important: when simplifying, calculations which do not contribute to the final result are thrown away. In the case of overflow errors, it may seem reasonable to consider the post-simplification rules, as we did in Ex. 3.4. In the case of for instance division by zero, we should add the errors to the pre-simplification rules.

Comment: When transforming a function into an LCTRS, we can *choose* what errors to model. For instance, we could ignore overflows (effectively assuming unbounded integers), but still test for division by zero. We could also let error_f be a constructor which takes an argument, i.e., $\text{error}_f : [\text{Errors}] \Rightarrow \text{result}_f \in \Sigma_{\text{terms}}$, where Errors is a sort with constructors IntegerOverflow, DivisionByZero, and so on.

3.4. Global Variables

Thus far, we have considered very *local* code: a function never calls other functions or modifies global variables. By altering the return constructors, we easily change the latter: we assume that a function symbol is given all global variables that it uses as input, and that it returns those global variables it alters as output, along with its return value. This change also allows for non-redundant void functions.

Example 3.5. Consider the following short program and its (simplified) translation:

```
int best;
int up(int x) {
  if (x > best) { best = x; return 1; }
  return 0;
}

```

$$\begin{array}{l} \text{up}(b, x) \rightarrow \text{return}_{\text{up}}(x, 1) \quad [x > b] \\ \text{up}(b, x) \rightarrow \text{return}_{\text{up}}(b, 0) \quad [x \leq b] \end{array}$$

3.5. Function Calls

Next, let us consider *function calls*. A difficulty is that they may occur in an expression, e.g., $\text{fact}(3) + 5$, which is not well sorted in the corresponding LCTRS: $\text{fact}(3)$ has sort $\text{result}_{\text{fact}}$, not int . To avoid this issue, and to propagate errors, we split off function calls occurring inside expressions other than $\text{var} = \text{func}(\text{arg}_1, \dots, \text{arg}_n)$ and store their return value into a temporary variable. For example:

```
int ncr(int x, int y) {
  int a = fact(x);
  int b = fact(y) * fact(x - y);
  return a / b;
}

```

$$\Longrightarrow$$

```
int ncr(int x, int y) {
  int a = fact(x);
  int tmp1 = fact(y);
  int tmp2 = fact(x - y);
  int b = tmp1 * tmp2;
  return a / b;
}

```

This change may cause declarations at places in the function where a C compiler would not accept them, but for the translation, this is no issue. We translate the resulting function by executing function calls in a separate parameter and using a separate step to examine the outcome of a function call and assign it to the relevant variable(s).

Example 3.6. The ncr program above is transformed to the following optimized LCTRS (where we test for division by zero but not integer overflow for simplicity):

$$\begin{array}{l} \text{ncr}(x, y) \rightarrow \text{u}_2(x, y, \text{fact}(x)) \quad \text{u}_2(x, y, \text{error}_{\text{fact}}) \rightarrow \text{error}_{\text{ncr}} \\ \text{u}_2(x, y, \text{return}_{\text{fact}}(k)) \rightarrow \text{u}_3(x, y, k, \text{fact}(y)) \quad \text{u}_3(x, y, a, \text{error}_{\text{fact}}) \rightarrow \text{error}_{\text{ncr}} \\ \text{u}_3(x, y, a, \text{return}_{\text{fact}}(k)) \rightarrow \text{u}_4(x, y, a, k, \text{fact}(x - y)) \quad \text{u}_4(x, y, a, t_1, \text{error}_{\text{fact}}) \rightarrow \text{error}_{\text{ncr}} \\ \text{u}_4(x, y, a, t_1, \text{return}_{\text{fact}}(k)) \rightarrow \text{error}_{\text{ncr}} \quad [t_1 * k = 0] \\ \text{u}_4(x, y, a, t_1, \text{return}_{\text{fact}}(k)) \rightarrow \text{return}_{\text{ncr}}(a \text{ div } (t_1 * k)) \quad [t_1 * k \neq 0] \end{array}$$

3.6. Statically Allocated Arrays

Finally, let us consider arrays. After we have seen Ex. 1.1 and the way side effects were handled in § 3.4, this is largely as expected. For now, we will not consider aliasing.

To start, we must fix a theory signature and corresponding interpretations. For a given theory sort ι which admits at least one value, say 0_ι , let $\text{array}(\iota)$ be a new sort and $\mathcal{I}_{\text{array}(\iota)} = \mathcal{I}_\iota^*$ —so each value corresponds to a finite sequence. We introduce the following theory symbols (in addition to $\Sigma_{\text{theory}}^{\text{int}}$ and other desired theories):

- $\text{size}_\iota : [\text{array}(\iota)] \Rightarrow \text{int}$: we define $\mathcal{J}_{\text{size}_\iota}(a)$ as the length of the sequence a .
- $\text{select}_\iota : [\text{array}(\iota) \times \text{int}] \Rightarrow \iota$: if $a = \langle a_0, \dots, a_{n-1} \rangle$, we define $\mathcal{J}_{\text{select}_\iota}(a, k) = a_k$ if $0 \leq k < n$ and $\mathcal{J}_{\text{select}_\iota}(a, k) = 0_\iota$ otherwise.
- $\text{store}_\iota : [\text{array}(\iota) \times \text{int} \times \iota] \Rightarrow \text{array}(\iota)$: if $a = \langle a_0, \dots, a_{n-1} \rangle$, we define $\mathcal{J}_{\text{store}_\iota}(a, k, v) = \langle a_0, \dots, a_{k-1}, v, a_{k+1}, \dots, a_{n-1} \rangle$ if $0 \leq k < n$ and $\mathcal{J}_{\text{store}_\iota}(a, k, v) = a$ otherwise.

We will usually omit the subscript ι when the sort is clear from context.

Our arrays are different from SMT-LIB (cf. <http://www.smt-lib.org/>), where arrays are functions from one (possibly infinite) domain to another. For program analysis, finite-length sequences seem practical instead. SMT problems on our arrays can be translated to SMT-LIB format using an additional integer variable a_{size} for the size of an array a and universal quantification to set entries outside the array to a fixed value.

We encode *lookups* $a[i]$ as $\text{select}(a, i)$; for *assignments* $a[i] = e$, we replace a by $\text{store}(a, i, e)$. To ensure correctness here, we add boundary checks to the constraint and reduce to error_f if such a check is not satisfied. After an assignment, the updated variable is included in the return value since the underlying memory of the array was altered.

Example 3.7. Consider the following C implementation of the `strcpy` function, which copies the contents of `original` into the array `goal`, until a 0 is reached.

```
void strcpy(char goal[], char original[]) {
  int i = 0;
  for (; original[i] != 0; i++) goal[i] = original[i];
  goal[i] = 0;
}
```

For simplicity, we think of strings as integer arrays (although alternative choices for $\mathcal{I}_{\text{char}}$ make little difference). The function never updates `original`, but may update `goal`, so the return value must include the latter. We obtain the following LCTRS:

$$\begin{aligned}
\text{strcpy}(gl, org) &\rightarrow v(gl, org, 0) \\
v(gl, org, i) &\rightarrow \text{error}_{\text{strcpy}} \quad [i < 0 \vee i \geq \text{size}(org)] \\
v(gl, org, i) &\rightarrow w(gl, org, i) \quad [0 \leq i < \text{size}(org) \wedge \text{select}(org, i) = 0] \\
v(gl, org, i) &\rightarrow \text{error}_{\text{strcpy}} \quad [0 \leq i < \text{size}(org) \wedge \text{select}(org, i) \neq 0 \wedge i \geq \text{size}(gl)] \\
v(gl, org, i) &\rightarrow v(\text{store}(gl, i, \text{select}(org, i)), org, i + 1) \\
&\quad [0 \leq i < \text{size}(org) \wedge \text{select}(org, i) \neq 0 \wedge i < \text{size}(gl)] \\
w(gl, org, i) &\rightarrow \text{error}_{\text{strcpy}} \quad [i < 0 \vee i \geq \text{size}(gl)] \\
w(gl, org, i) &\rightarrow \text{return}_{\text{strcpy}}(\text{store}(gl, i, 0)) \quad [0 \leq i < \text{size}(gl)]
\end{aligned}$$

Here, the notation $0 \leq i < \text{size}(org)$ is shorthand for $0 \leq i \wedge i < \text{size}(org)$. Note that this LCTRS could be further simplified by combining the third rule with the last two rules.

Comment: It should now be clear how the systems from § 1.1 have been translated from C code to LCTRSs. The only deviation is that there we have included the array `arr` in the return value of `sum1`, `sum2`, and `sum4`, which is not necessary as it is not modified in these cases. This was done to allow for a direct comparison with `sum3`, where the array *is* modified. In addition, the return and error symbols in these examples are not indexed, for the same reason.

3.7. Dynamically Allocated Arrays and Aliasing

The transformation in § 3.6 allows us to abstract from the underlying memory model when encoding arrays. This makes analysis easier, but does not allow for aliasing or pointer arithmetic beyond accessing an array element. As a result, properties we prove about `strcpy` from Ex. 3.7 might fail to hold for a call like `strcpy(a, a)`.

As we seek to handle only *part* of the language, this does not need to be an issue; in practice, a fair number of programs are written without explicit pointer use and with easily removable aliasing only. For example, we might replace `strcpy(a, a)` by `strcpy'(a)`, and create new rules for `strcpy'` by collapsing the variables in the rules for `strcpy`. To handle programs with more sophisticated pointer use, including dynamically allocated arrays, we can encode the memory as a list of arrays and pass this along as a variable. This is somewhat beyond the scope of this paper, but is explored in Appendix A.2.

3.8. Remarks

The treatment in this section is both informal and incomplete: we have discussed only a fraction of the C language—albeit an important fraction for verification. We believe that these ideas easily extend further, with for instance the `switch` statement, user-defined data structures, or standard library functions, as well as compiler-specific choices. Important to note is that the translation gives several *choices*. Most pertinently, we saw the choices what sort of interpretations to use (e.g., whether `int` should be mapped to the set of integers or bitvectors) and what errors to consider.

In this paper, and in line with our automatic translation at <http://www.trs.css.i.nagoya-u.ac.jp/c2lctrs/>, we have chosen to work with real integers and not test for overflows. We also do not permit aliasing. By avoiding the more sophisticated translation steps, we obtain LCTRSs which are correspondingly easier to analyze.

The LCTRSs from this transformation are well behaved: all rules are left-linear and non-overlapping,³ and have the property that all ground terms can be reduced or are constructor terms. Rules $\ell \rightarrow r [\varphi]$ can have variables in r or φ which do not occur in ℓ : this is mostly due to unspecified values in the C code. Where such variables do not occur—or are removed in the optimization step—the resulting LCTRSs are confluent.

4. REWRITING INDUCTION FOR LCTRSS

In this section, we adapt the inference rules from [Reddy 1990; Falke and Kapur 2012; Sakata et al. 2009] to inductive theorem proving with LCTRSs. This provides the core theory for rewriting induction, strengthened with two generalization techniques in § 5.

We start by listing some restrictions we need to impose on LCTRSs for the method to work (§ 4.1). Then, we provide the theory for the technique (§ 4.2) and some illustrative examples (§ 4.3). Compared to older definitions of rewriting induction, we make several changes to best handle the new formalism. We complete by proving correctness (§ 4.4).

4.1. Restrictions

In order for rewriting induction to be successful, we need to impose certain restrictions.

Definition 4.1. In the following, we limit interest to LCTRSs which satisfy (1)–(4):

- (1) all core theory symbols are present in Σ_{theory} : $\Sigma_{theory} \supseteq \Sigma_{theory}^{core}$;
- (2) the LCTRS is *terminating*: there is no infinite reduction $s_1 \rightarrow_{\mathcal{R}} s_2 \rightarrow_{\mathcal{R}} \dots$;

³Non-overlappingness means that for every term s and rule $\rho: \ell \rightarrow r [\varphi]$ such that s reduces with ρ at the root position: (a) there are no other rules ρ' such that s reduces with ρ' at the root position, and (b) if s reduces with any rule at a non-root position q , then q is not a position of ℓ . For our translations, this holds because (a) rules with the same defined symbol have either incompatible constraints or non-unifiable arguments, and (b) in a rule $f(\ell_1, \dots, \ell_n) \rightarrow r [\varphi]$, the terms ℓ_i do not contain defined or calculation symbols.

- (3) the system is *quasi-reductive*: i.e., for every ground term s either $s \in \text{Terms}(\text{Cons}, \emptyset)$ (we say s is a *ground constructor term*), or there is some t such that $s \rightarrow_{\mathcal{R}} t$;
- (4) there are ground terms of every sort occurring in Σ .

Property 1 is the standard assumption from § 2. We will need symbols such as $=$, \wedge and \Rightarrow to add new information to a constraint. Termination (property 2) essentially indicates that a program cannot run indefinitely; this is crucial for our inductive reasoning, as the method uses induction on an extension of $\rightarrow_{\mathcal{R}}$ on terms.

Property 3 indicates that an evaluation cannot get “stuck”; roughly, that pattern matching and case analysis are exhaustive. Termination and quasi-reductivity together ensure that every ground term reduces to a constructor term. This makes it possible to do an exhaustive case analysis on the rules applicable to an equation, and lets us assume that variables are always instantiated by ground constructor terms.

The last property is natural, since inductive theorem proving makes a statement on *ground* terms; there is no point in regarding empty sorts. Together with quasi-reductivity and termination, this implies that all sorts admit ground *constructor* terms.

Methods to prove both quasi-reductivity and termination have previously been published for different styles of constrained rewriting; see e.g. [Falke and Kapur 2012] for quasi-reductivity and [Falke 2009; Sakata et al. 2011] for termination. These methods are easily adapted to LCTRSs. Quasi-reductivity is handled in [Kop 2017] and is moreover always satisfied by systems obtained from the transformations in § 3. Some basics of termination analysis for LCTRSs are discussed in [Kop 2013].

Example 4.2. As a running example in this section, we will consider $\mathcal{R}_{\text{fact}}$, which combines the factorial function from Ex. 3.2 with a recursive variant obtained from `int fact(int x) { if (x <= 1) return 1; else return x * fact(x - 1); }`.

- | | |
|--|--|
| (1) $\text{factiter}(x) \rightarrow \text{iter}(x, 1, 1)$ | (4) $\text{factrec}(x) \rightarrow \text{return}(1)$ $[x \leq 1]$ |
| (2) $\text{iter}(x, z, i) \rightarrow \text{iter}(x, z * i, i + 1)$ $[i \leq x]$ | (5) $\text{factrec}(x) \rightarrow$ |
| (3) $\text{iter}(x, z, i) \rightarrow \text{return}(z)$ $[i > x]$ | $\text{mul}(x, \text{factrec}(x - 1))$ $[x > 1]$ |
| | (6) $\text{mul}(x, \text{return}(y)) \rightarrow \text{return}(x * y)$ |

(Function symbols were renamed for readability.) We can choose a signature which includes $\Sigma_{\text{theory}}^{\text{core}}$, and each of the sorts—`int`, `bool`, `result`—clearly admits ground terms (e.g., `0`, `false`, `return(0)`). The system was obtained using § 3, so is quasi-reductive. Termination follows because in the recursive rule (2), the value $x - i$ is decreased, while bounded from below by 0, and in the recursion in rule (5), x decreases against the bound 1. This could be proved using, e.g., interpretations with support for built-in integers and non-theory symbols [Fuhs et al. 2009], and is automatically handled by our tool Ctrl.

4.2. Rewriting Induction

We now introduce the notions of *constrained equations* and *inductive theorems*.

Definition 4.3. A (*constrained*) *equation* is a triple $s \approx t [\varphi]$ with s and t terms and φ a constraint. We write $s \simeq t [\varphi]$ to denote either $s \approx t [\varphi]$ or $t \approx s [\varphi]$. A substitution γ *respects* $s \approx t [\varphi]$ if γ respects φ and $\text{Var}(s) \cup \text{Var}(t) \subseteq \text{Dom}(\gamma)$; it is called a *ground constructor substitution* if all $\gamma(x)$ with $x \in \text{Dom}(\gamma)$ are ground constructor terms.

An equation $s \approx t [\varphi]$ is an *inductive theorem* of an LCTRS \mathcal{R} if $s\gamma \leftrightarrow_{\mathcal{R}}^* t\gamma$ for any ground constructor substitution γ that respects this equation.

Intuitively, if an equation $f(\vec{x}) \approx g(\vec{x}) [\varphi]$ is an inductive theorem, then f and g define the same function (conditional on φ , and assuming confluence). As we require termination, we thus consider *total equivalence* in the categorization of Godlin and Strichman [2008]: on all inputs, both programs terminate and return the same values.

To prove that an equation is an inductive theorem, we consider nine inference rules, in §§ 4.2.1–4.2.9. Four originate in [Reddy 1990]; three are based on extensions [Bouhoula 1997; Falke and Kapur 2012; Sakata et al. 2009]; two are new. All these rules modify a triple $(\mathcal{E}, \mathcal{H}, b)$, called a *proof state*. Here, \mathcal{E} is a set of equations, \mathcal{H} a set of rules with $\rightarrow_{\mathcal{R} \cup \mathcal{H}}$ terminating, and $b \in \{\text{COMPLETE}, \text{INCOMPLETE}\}$. A rule in \mathcal{H} plays the role of an *induction hypothesis* for “proving” the equations in \mathcal{E} and is called an *induction rule*. The flag b indicates whether we can use the current proof state to *refute* that the initial equation is an inductive theorem; we can do so if $b = \text{COMPLETE}$.

The definition of these rules is used in the following result, proved in § 4.4.

THEOREM 4.4. *Let an LCTRS with rules \mathcal{R} and signature Σ , satisfying the restrictions from Def. 4.1, be given. Let \mathcal{E} be a finite set of equations and let $\text{flag} = \text{COMPLETE}$ if we can confirm that \mathcal{R} is confluent and $\text{flag} = \text{INCOMPLETE}$ otherwise. If $(\mathcal{E}, \emptyset, \text{flag}) \vdash_{\text{ri}}^* (\emptyset, \mathcal{H}, \text{flag}')$ for some $\mathcal{H}, \text{flag}'$, then every equation in \mathcal{E} is an inductive theorem of \mathcal{R} . If $(\mathcal{E}, \emptyset, \text{flag}) \vdash_{\text{ri}}^* \perp$, then there is some equation in \mathcal{E} that is not an inductive theorem of \mathcal{R} .*

Example 4.5. We will illustrate the various rules by proving that *factrec* and *factiter* are equivalent on positive input,⁴ by showing that (FCT.A) is an inductive theorem:

$$\text{(FCT.A) } \text{factrec}(n) \approx \text{factiter}(n) [n \geq 1]$$

$\mathcal{R}_{\text{fact}}$ is confluent: as seen in § 3.8, it is left-linear and non-overlapping, and the right-hand sides do not introduce fresh variables, so confluence is given by [Kop and Nishida 2013, Thm. 4]. Thus, we will start with the proof state $(\{ \text{(FCT.A)} \}, \emptyset, \text{COMPLETE})$.

Let us now define the nine inference rules to reduce proof states.

4.2.1. SIMPLIFICATION. Our first inference rule originates in [Reddy 1990] and can be considered one of the core rules of rewriting induction.

Definition 4.6. If $s \approx t[\varphi] \rightarrow_{\mathcal{R} \cup \mathcal{H}} u \approx t[\psi]$, where \approx is seen as a fresh constructor for the purpose of constrained term reduction,⁵ then we may derive:

$$(\mathcal{E} \uplus \{(s \simeq t[\varphi])\}, \mathcal{H}, b) \vdash_{\text{ri}} (\mathcal{E} \cup \{(u \approx t[\psi])\}, \mathcal{H}, b)$$

This inference rule allows us to reduce one side of an equation. This is altered from Reddy’s definition by using constrained rather than normal reduction.

Example 4.7. Following Ex. 4.5, we observe that *factiter*(n) can be reduced by the unconstrained rule (1). Thus, using SIMPLIFICATION we obtain the proof state:

$$(\{ \text{(FCT.B)} : \text{iter}(n, 1, 1) \approx \text{factrec}(n) [n \geq 1] \}, \emptyset, \text{COMPLETE})$$

Here we reduce the right-hand side of the equation (recall that $s \simeq t$ in the rule means $s \approx t$ or $t \approx s$); the reduced term moves to the left-hand side of the new equation. Next, observe that *iter*($n, 1, 1$) can be reduced by rule (2) if $n \geq 1$; SIMPLIFICATION then gives:

$$(\{ \text{(FCT.C)} : \text{iter}(n, 1 * 1, 1 + 1) \approx \text{factrec}(n) [n \geq 1] \}, \emptyset, \text{COMPLETE})$$

Recall that constrained reduction also allows for steps with calculation rules; see, e.g., Ex. 2.16. The added complexity is that we must decide how to handle the fresh variable these rules introduce. In this paper we will use the following strategy:

⁴We limit interest to positive input for demonstration purposes only: these functions give the same result on *all* input, but considering only $n \geq 1$ allows us to apply the inference rules in a convenient order.

⁵It does not suffice if $s[\varphi] \rightarrow_{\mathcal{R}} u[\psi]$: when reducing constrained terms, unused variables may be manipulated at will, which causes problems if they are used in t . For example,

$$f(x+0) [x > y] \sim f(x+0) [z = x+0] \rightarrow_{\text{base}} f(z) [z = x+0] \sim f(x) [x < y]$$

but we should certainly not replace an equation $f(x+0) \approx g(y) [x > y]$ by $f(x) \approx g(y) [x < y]$.

- if $s \rightarrow_{\text{calc}} u$ then $s \approx t[\varphi]$ is simplified to $u \approx t[\varphi]$, e.g. $f(0 + 1) \approx r[\varphi]$ reduces to $f(1) \approx r[\varphi]$;
- a calculation containing variables can be replaced by a fresh variable, which is defined in the (updated) constraint, e.g. $f(x + 1) \approx r[\varphi]$ reduces to $f(y) \approx r[\varphi \wedge y = x + 1]$; if such a definition already occurs in the constraint, the relevant variable is used instead, e.g. $f(x + 1) \approx r[\varphi \wedge y = x + 1]$ reduces to $f(y) \approx r[\varphi \wedge y = x + 1]$.

Example 4.8. The proof state from Ex. 4.7 is further simplified to:

$$(\{(\text{FCT.D}) : \text{iter}(n, 1, 2) \approx \text{factrec}(n) [n \geq 1]\}, \emptyset, \text{COMPLETE})$$

4.2.2. EXPANSION. Our second core rule also originates from [Reddy 1990], but has been more heavily adapted to support irregular rules.

Definition 4.9. Let s, t be terms and φ a constraint, all with variables distinct from those in \mathcal{R} (we can always rename the variables in the rules to support this), and p a position of s . Let $\text{Expd}(s \approx t[\varphi], p)$ be a set of equations containing, for all rules $\ell \rightarrow r[\psi] \in \mathcal{R}$ such that ℓ is unifiable with $s|_p$ with most general unifier γ , an equation $s' \approx t'[\varphi']$ where $s\gamma \approx t\gamma [(\varphi\gamma) \wedge (\psi\gamma)] \rightarrow_{\mathcal{R}} s' \approx t'[\varphi']$ with rule $\ell \rightarrow r[\psi]$ at position $1 \cdot p$. Here, as in SIMPLIFICATION, \approx is seen as a fresh constructor for the reduction. If $s|_p$ is basic (i.e., $s|_p = f(s_1, \dots, s_n)$ with $f \in \mathcal{D}$ and all s_i constructor terms), we may derive:

$$(\mathcal{E} \uplus \{s \simeq t[\varphi]\}, \mathcal{H}, b) \vdash_{\text{ri}} (\mathcal{E} \cup \text{Expd}(s \approx t[\varphi], p), \mathcal{H}, b)$$

If, moreover, $\mathcal{R} \cup \mathcal{H} \cup \{s \rightarrow t[\varphi]\}$ is terminating, we may even derive:

$$(\mathcal{E} \uplus \{s \simeq t[\varphi]\}, \mathcal{H}, b) \vdash_{\text{ri}} (\mathcal{E} \cup \text{Expd}(s \approx t[\varphi], p), \mathcal{H} \cup \{s \rightarrow t[\varphi]\}, b)$$

Intuitively, this inference rule uses narrowing for a *case analysis*: *Expd* generates all resulting equations if a ground constructor instance of $s \approx t[\varphi]$ is reduced at position p of s . In addition, we save the current equation as a rule to take an induction step.

Example 4.10. Following Ex. 4.8, we consider which rules may apply to an instance of $\text{factrec}(n)$ with $n \geq 1$. For $\text{Expd}(\text{factrec}(n) \approx \text{iter}(n, 1, 2) [n \geq 1], \epsilon)$, we choose:

$$\left\{ \begin{array}{l} (\text{FCT.E}): \quad \text{return}(1) \approx \text{iter}(n, 1, 2) [n \geq 1 \wedge n \leq 1], \\ (\text{FCT.F}): \quad \text{mul}(n, \text{factrec}(n-1)) \approx \text{iter}(n, 1, 2) [n \geq 1 \wedge n > 1] \end{array} \right\}$$

In both cases we used the unifier $\gamma = [x := n]$. If we write (FCT.D^{-1}) for the rule generated from the inverse of (FCT.D) —so $\text{factrec}(n) \rightarrow \text{iter}(n, 1, 2) [n \geq 1] - \mathcal{R} \cup \{(\text{FCT.D}^{-1})\}$ is terminating as the new rule does not cause mutual recursion between *iter* and *factrec*. We continue with $(\{(\text{FCT.E}), (\text{FCT.F})\}, \{(\text{FCT.D}^{-1})\}, \text{COMPLETE})$. Now we can show the second kind of calculation step, using SIMPLIFICATION on (FCT.F) , which gives:

$$\left(\left\{ \begin{array}{l} (\text{FCT.E}): \quad \text{return}(1) \approx \text{iter}(n, 1, 2) [n \geq 1 \wedge n \leq 1], \\ (\text{FCT.G}): \quad \text{mul}(n, \text{factrec}(m)) \approx \text{iter}(n, 1, 2) [n > 1 \wedge m = n - 1] \end{array} \right\}, \{(\text{FCT.D}^{-1})\}, \text{COMPLETE} \right)$$

Here, we also removed the redundant clause $n \geq 1$, which is allowed by definition of $\rightarrow_{\mathcal{R}}$ on constrained terms. As $n \geq 1 \wedge n \leq 1$ implies $n = 1$, we may use SIMPLIFICATION with rule (3) on (FCT.E) , and with rule (2) followed by calculations on (FCT.G) , to get:

$$\left(\left\{ \begin{array}{l} (\text{FCT.H}): \quad \text{return}(1) \approx \text{return}(1) [n = 1], \\ (\text{FCT.I}): \quad \text{iter}(n, 2, 3) \approx \text{mul}(n, \text{factrec}(m)) [n > 1 \wedge m = n - 1] \end{array} \right\}, \{(\text{FCT.D}^{-1})\}, \text{COMPLETE} \right)$$

Now we can use “induction”: we eliminate the occurrence of *factrec* with a SIMPLIFICATION step using the induction rule (FCT.D^{-1}) and substitution $[n := m]$. This gives:

$$\left(\left\{ \begin{array}{l} (\text{FCT.H}): \quad \text{return}(1) \approx \text{return}(1) [n = 1], \\ (\text{FCT.J}): \quad \text{mul}(n, \text{iter}(m, 1, 2)) \approx \text{iter}(n, 2, 3) [n > 1 \wedge m = n - 1] \end{array} \right\}, \{(\text{FCT.D}^{-1})\}, \text{COMPLETE} \right)$$

Note that the choice of $Expd$ is non-deterministic, as it uses reduction of constrained terms. The most natural choice for $Expd(s \approx t [\varphi], p)$ —which we use in examples—is

$$\{ s[r]_p \gamma \approx t\gamma [(\varphi\gamma) \wedge (\psi\gamma)] \mid \ell \rightarrow r [\psi] \in \mathcal{R}, s|_p \text{ unifies with } \ell \text{ with mgu } \gamma \}$$

However, for *irregular* rules in particular, it may be strategic to choose a different set. Consider for example a (non-confluent) LCTRS with rules $f(x) \rightarrow g(y) [x > 0 \wedge x > y]$ and $f(x) \rightarrow g(y) [x \leq 0 \wedge x \leq y]$. With the choice for $Expd(s \approx t [\varphi], p)$ above, an equation $f(x) \approx g(0) [\text{true}]$ results in $\{ g(y) \approx g(0) [x > 0 \wedge x > y], g(y) \approx g(0) [x \leq 0 \wedge x \leq y] \}$. If g is a constructor, neither of these equations can be handled. Using the full definition of EXPANSION, we can choose $g(0) \approx g(0) [\text{true}]$ for both equations.

Also note that there is no choice in the orientation of the rule added to \mathcal{H} : this is determined by the side of the equation on which the expansion was applied. Thus, in Ex. 4.10 we were not allowed to add (FCT.D) instead of (FCT.D⁻¹).

Our definition of EXPANSION differs from both its original and existing work on constrained rewriting induction. To start, those works define $Expd(s \approx t [\varphi], p)$ simply as the “natural choice” given above. Second, we included a case where no rule is added, to allow for progress when adding the rule might cause non-termination. Forms of this case appear as a separate rule in other work, e.g., CASE ANALYSIS in [Bouhoula 1997] and REWRITE/PARTIAL SPLITTING in [Bouhoula and Jacquemard 2008b; 2008a]. A weaker form with constraints is given in [Falke and Kapur 2012] (CASE-SIMPLIFY).

4.2.3. DELETION. The last of the core rules serves to remove solved equations from \mathcal{E} .

Definition 4.11. If $s = t$ or φ is not satisfiable, we can delete $s \approx t [\varphi]$ from \mathcal{E} :

$$(\mathcal{E} \uplus \{s \approx t [\varphi]\}, \mathcal{H}, b) \vdash_{\text{ri}} (\mathcal{E}, \mathcal{H}, b)$$

Compared to the corresponding rule in [Reddy 1990], the unsatisfiability case is new; it is similar to the corresponding rules in [Sakata et al. 2009; Falke and Kapur 2012].

Example 4.12. Following Ex. 4.10, the left- and right-hand side of (FCT.H) are the same, so we may remove the equation with DELETION, obtaining $(\{ \text{(FCT.J)} \}, \{ \text{(FCT.D}^{-1}) \}, \text{COMPLETE})$. We will see the other form of DELETION in Ex. 4.18.

4.2.4. POSTULATE. Sometimes it is useful to make the problem *seemingly* harder. To this end, we consider the last inference rule from [Reddy 1990].

Definition 4.13. For any set of equations \mathcal{E}' , we can derive:

$$(\mathcal{E}, \mathcal{H}, b) \vdash_{\text{ri}} (\mathcal{E} \cup \mathcal{E}', \mathcal{H}, \text{INCOMPLETE})$$

The POSTULATE rule allows us to add additional equations to \mathcal{E} (although at a price: we cannot conclude non-equivalence after adding a potentially unsound equation). The reason to do so is that in proving the equations in \mathcal{E}' to be inductive theorems, we may derive new induction rules. These can then be used to simplify the elements of \mathcal{E} .

Example 4.14. Following Ex. 4.12, EXPANSION followed by SIMPLIFICATION gives:

$$\text{(FCT.K): } \text{mul}(n, \text{iter}(m, 2, 3)) \approx \text{iter}(n, 6, 4) [n \geq 3 \wedge m = n - 1]$$

But now a pattern starts to arise. Expanding and fully simplifying again, we obtain:

$$\text{(FCT.L): } \text{mul}(n, \text{iter}(m, 6, 4)) \approx \text{iter}(n, 24, 5) [n \geq 4 \wedge m = n - 1]$$

And so on. Here, (FCT.K) cannot be handled by the induction rule (FCT.J⁻¹), nor can (FCT.L) be handled by (FCT.K⁻¹). We have a *divergence*: a sequence of increasingly complex equations, each generated from the same leg in an EXPANSION (see also the *divergence critic* in [Walsh 1996]). Yet the previous induction rules never apply to the new equation. This suggests we need a lemma equation. We use POSTULATE to get:

$$\left(\left(\begin{array}{l} \text{(FCT.J):} \\ \text{(FCT.M):} \end{array} \begin{array}{l} \text{mul}(n, \text{iter}(m, 1, 2)) \approx \text{iter}(n, 2, 3) \\ \text{mul}(n, \text{iter}(m, x, y)) \approx \text{iter}(n, x', y') \end{array} \right) \left. \begin{array}{l} [n > 1 \wedge m = n - 1] \\ [n \geq y \wedge m = n - 1 \wedge y' = y + 1 \wedge x' = x * y] \end{array} \right\}, \left. \begin{array}{l} \{(\text{FCT.D}^{-1})\}, \\ \text{INCOMPLETE} \end{array} \right)$$

Using EXPANSION on the right-hand of (FCT.M), we have:

$$\left(\left(\begin{array}{l} \text{(FCT.J):} \\ \text{(FCT.N):} \\ \text{(FCT.O):} \end{array} \begin{array}{l} \text{mul}(n, \text{iter}(m, 1, 2)) \approx \text{iter}(n, 2, 3) \\ \text{iter}(n, x' * y', y' + 1) \approx \text{mul}(n, \text{iter}(m, x, y)) \\ \text{return}(x') \approx \text{mul}(n, \text{iter}(m, x, y)) \end{array} \right) \left. \begin{array}{l} [n > 1 \wedge m = n - 1] \\ [n \geq y \wedge m = n - 1 \wedge y' = y + 1 \wedge x' = x * y \wedge y' \leq n] \\ [n \geq y \wedge m = n - 1 \wedge y' = y + 1 \wedge x' = x * y \wedge y' > n] \end{array} \right\}, \left. \begin{array}{l} \{(\text{FCT.D}^{-1})\} \\ \{(\text{FCT.M}^{-1})\} \\ \text{INCOMPLETE} \end{array} \right)$$

But now we have added (FCT.M⁻¹) as an induction rule. As a result—since $n > 1$ clearly implies $n \geq 2$ —we can use SIMPLIFICATION with a substitution $[n := n, x := 1, y := 2, x' := 2, y' := 3]$ to reduce (FCT.J) to the equation $\text{mul}(n, \text{iter}(m, 1, 2)) \approx \text{mul}(n, \text{iter}(m, 1, 2))$ [...], which we may immediately remove by DELETION. We continue with the proof state $(\{(\text{FCT.N}), (\text{FCT.O})\}, \{(\text{FCT.D}^{-1}), (\text{FCT.M}^{-1})\}, \text{INCOMPLETE})$.

Although the need to choose arbitrary new equations for use in POSTULATE may seem somewhat problematic, this is actually a key step. Complex theorems typically require more than straight induction, both in our setting and in mathematical proofs in general. Thus, generation of suitable *lemma equations* \mathcal{E}' is not only part, but even at the heart, of inductive theorem proving. Hence, this subject has been extensively investigated [Bundy et al. 2005; Kapur and Sakhanenko 2003; Kapur and Subramaniam 1996; Nakabayashi et al. 2010; Urso and Kounalis 2004; Walsh 1996], and a large variety of lemma generation techniques exist, at least in the setting without constraints.

4.2.5. GENERALIZATION. A very typical use of POSTULATE is to *generalize* a problematic equation. For simplicity, we add a shortcut to do this in one step.

Definition 4.15. If for all substitutions γ which respect φ there is a substitution δ which respects ψ with $s\gamma = s'\delta$ and $t\gamma = t'\delta$, then we can derive:

$$(\mathcal{E} \uplus \{s \approx t [\varphi]\}, \mathcal{H}, b) \vdash_{\text{ri}} (\mathcal{E} \cup \{s' \approx t' [\psi]\}, \mathcal{H}, \text{INCOMPLETE})$$

This inference rule is rarely *necessary*: we could usually add $s' \approx t' [\psi]$ using POSTULATE, and use the resulting induction rules to eliminate $s \approx t [\varphi]$, as we did in Ex. 4.14. By generalizing instead, we avoid extra steps, and intuitively, we strengthen an induction statement rather than add a separate lemma. Without constraints, GENERALIZATION can be seen as a combination of POSTULATE and the SUBSUMPTION rule in [Bouhoula 1997]. As there are several results for generalizing equations in the literature [Bundy et al. 1993; Bundy et al. 2005; Basin and Walsh 1992; Walsh 1996; Urso and Kounalis 2004], the combination is useful beyond just this paper.

Example 4.16. In Ex. 4.14, we could have used GENERALIZATION immediately to move from the proof state $(\{(\text{FCT.J})\}, \{(\text{FCT.D}^{-1})\}, \text{INCOMPLETE})$ to $(\{(\text{FCT.M})\}, \{(\text{FCT.D}^{-1})\}, \text{INCOMPLETE})$.

4.2.6. EQ-DELETION. The following rule, which was adapted from [Sakata et al. 2009], provides a link between the equation part $s \approx t$ and the constraint.

Definition 4.17. Let C be an arbitrary context with n holes (C may contain symbols in Σ_{theory}). If all $s_i, t_i \in \text{Terms}(\Sigma_{\text{theory}}, \text{Var}(\varphi))$, then we can derive:

$$\begin{aligned} & (\mathcal{E} \uplus \{C[s_1, \dots, s_n] \simeq C[t_1, \dots, t_n] [\varphi]\}, \mathcal{H}, b) \vdash_{\text{ri}} \\ & (\mathcal{E} \cup \{C[s_1, \dots, s_n] \approx C[t_1, \dots, t_n] [\varphi \wedge \neg(\bigwedge_{i=1}^n s_i = t_i)]\}, \mathcal{H}, b) \end{aligned}$$

Intuitively, if $\bigwedge_{i=1}^n s_i = t_i$ holds, then $C[s_1, \dots, s_n]\gamma \leftrightarrow_{\mathcal{R}_{\text{calc}}}^* C[t_1, \dots, t_n]\gamma$, so we are done. EQ-DELETION excludes this case from the equation. In combination with DELETION, this rule gives a more general variation of THEORY $_{\top}$ in [Falke and Kapur 2012].

Example 4.18. Continuing from Ex. 4.14 (or Ex. 4.16), we observe that $n \geq y$, $y' = y + 1$ and $y' > n$ together imply $n = y$, and with $m = n - 1$ we thus have $y > m$ as well. Therefore, SIMPLIFICATION on (FCT.O) by rule (3) followed by (6) gives:

$$\text{(FCT.P): } \text{return}(n * x) \approx \text{return}(x') [n = y \wedge m = n - 1 \wedge y' = y + 1 \wedge x' = x * y]$$

We can use EQ-DELETION with the context $C[\square] = \text{return}(\square)$ to replace (FCT.P) by:

$$\text{(FCT.Q): } \text{return}(n * x) \approx \text{return}(x') [n = y \wedge m = n - 1 \wedge y' = y + 1 \wedge x' = x * y \wedge \neg(n * x = x')]$$

As $n = y$ and $x' = x * y$ together imply that $n * x = x'$, the constraint of this equation is not satisfiable. We may remove it using DELETION, giving the proof state $(\{ \text{(FCT.N)} \}, \{ \text{(FCT.D}^{-1}), \text{(FCT.M}^{-1}) \}, \text{INCOMPLETE})$.

EQ-DELETION is among the core rules for constrained rewriting induction: almost all inductive proofs use it, in contrast to the remaining three inference rules.

Example 4.19. To complete our example, consider (FCT.N). As $y + 1 = y' \leq n \wedge m = n - 1$ implies $y \leq m$, we may apply SIMPLIFICATION with rule (2) to replace it by:

$$\text{(FCT.R): } \text{mul}(n, \text{iter}(m, x * y, y + 1)) \approx \text{iter}(n, x' * y', y' + 1) \\ [n \geq y \wedge m = n - 1 \wedge y' = y + 1 \wedge x' = x * y \wedge y' \leq n]$$

Then, using SIMPLIFICATION with calculations (and observing that both $x * y$ and $y + 1$ are “defined” in the constraint, as discussed in § 4.2.1), we get:

$$\text{(FCT.S): } \text{mul}(n, \text{iter}(m, x', y')) \approx \text{iter}(n, x'', y'') \\ [n \geq y' \wedge m = n - 1 \wedge x' = x * y \wedge x'' = x' * y' \wedge y'' = y' + 1]$$

(We removed the clauses with y from the constraint, as y does not occur in the equation part.) But now the induction rule (FCT.M $^{-1}$) applies! As this rule is irregular, we must be careful. We use the substitution $\gamma = [n := n, m := m, x'' := x', y' := y'', x := x', y := y']$, which also affects variables not occurring in the left-hand side. The substituted constraint for the rule is $n \geq y' \wedge m = n - 1 \wedge y'' = y' + 1 \wedge x'' = x' * y'$, which is indeed implied by the constraint of (FCT.S). Using SIMPLIFICATION, we thus obtain:

$$\left(\left\{ \begin{array}{l} \text{(FCT.T): } \text{mul}(n, \text{iter}(m, x', y')) \approx \text{mul}(n, \text{iter}(m, x', y')) \\ [n \geq y' \wedge m = n - 1 \wedge x' = x * y \wedge x'' = x' * y' \wedge y'' = y' + 1] \end{array} \right\}, \{ \dots \}, \text{INCOMPLETE} \right)$$

As the left- and right-hand side of the remaining equation are the same, we may remove it using DELETION. This leaves a proof state of the form $(\emptyset, \mathcal{H}, \text{INCOMPLETE})$, so by Thm. 4.4, the equation $\text{factrec}(n) \approx \text{factiter}(n) [n \geq 1]$ is an inductive theorem.

4.2.7. CONSTRUCTOR. Where Falke and Kapur [2012] and Sakata et al. [2009] focus on systems with only theory symbols and defined symbols, here we are also interested in non-theory constructors, such as error_f and return_f . To support this, we add:

Definition 4.20. If f is a constructor, we can derive:

$$(\mathcal{E} \uplus \{f(s_1, \dots, s_n) \approx f(t_1, \dots, t_n)[\varphi]\}, \mathcal{H}, b) \vdash_{\text{ri}} (\mathcal{E} \cup \{s_i \approx t_i [\varphi] \mid 1 \leq i \leq n\}, \mathcal{H}, b)$$

The CONSTRUCTOR rule originates in [Bouhoula 1997], where it is called POSITIVE DECOMPOSITION, although variations occur in earlier work on implicit induction, e.g., [Huet and Hullot 1982]. It is used to split up a large equation into smaller problems. This inference rule is particularly useful in applications where a recursive structure, such as a list, is inductively built up, but will also be invaluable as part of a disproof.

Example 4.21. Suppose that, in Ex. 4.5, we had started with (BAD.A): $\text{factiter}(x) \approx \text{factrec}(x - 1)$ [true]. Following some expansions and simplifications, we arrive at

$$\left(\left\{ \begin{array}{l} \text{(BAD.B): } \text{return}(2) \approx \text{return}(1) \ [x = 2] \\ \text{(BAD.C): } \text{iter}(x, 1, 1) \approx \text{factrec}(y) \ [y = x - 1 \wedge y > 1] \end{array} \right\}, \mathcal{H}, \text{COMPLETE} \right)$$

(for some \mathcal{H}). We can use CONSTRUCTOR to replace (BAD.B) by (BAD.D): $2 \approx 1$ [$x = 2$].

4.2.8. DISPROVE. Recall that, to show that an equation is *not* an inductive theorem, we must derive \perp from a COMPLETE proof state. For this, we use DISPROVE.

Definition 4.22. Suppose $\vdash s : \iota$ and one of the following holds:

- $s, t \in \text{Terms}(\Sigma_{\text{theory}}, \mathcal{V})$, ι is a theory sort, and $\varphi \wedge s \neq t$ is satisfiable;
- $s = f(\vec{s})$ and $t = g(\vec{t})$ with f, g distinct constructors and φ satisfiable;
- $s \in \mathcal{V} \setminus \text{Var}(\varphi)$, φ is satisfiable, at least two different constructors have output sort ι , and either t is a variable distinct from s or t has the form $g(\vec{t})$ with $g \in \text{Cons}$;

Then we may derive:

$$(\mathcal{E} \uplus \{s \simeq t [\varphi]\}, \mathcal{H}, \text{COMPLETE}) \vdash_{\text{ri}} \perp$$

The first case of this rule corresponds to THEORY $_{\top}$ in [Falke and Kapur 2012] and Thm. 7.2 in [Sakata et al. 2009]; note that the restriction to theory sorts only excludes the case where s and t are non-logical variables. The second case corresponds to POSITIVE CLASH in [Bouhoula 1997]. The third case is new in rewriting induction, but appears in [Huet and Hullot 1982], an implicit induction method based on completion.

Example 4.23. Following Ex. 4.21, we observe that $x = 2 \wedge 2 \neq 1$ is satisfiable. Thus, by DISPROVE we reduce $(\{ \text{(BAD.D)}, \text{(BAD.C)} \}, \mathcal{H}, \text{COMPLETE})$ to \perp . By confluence of $\mathcal{R}_{\text{fact}}$, we see that $\text{factiter}(x)$ and $\text{factrec}(x - 1)$ have different normal forms for some x .

4.2.9. COMPLETENESS. A downside of POSTULATE and GENERALIZATION is the potential loss of the completeness flag. To weaken this problem—and empower automatic tools to combine the search for a proof and a disproof—we add our final inference rule.

Definition 4.24. For any set of equations \mathcal{E} and $\mathcal{E}' \subseteq \mathcal{E}$ we can derive:

$$\begin{array}{l} \text{If } (\mathcal{E}, \mathcal{H}, \text{COMPLETE}) \vdash_{\text{ri}}^* (\mathcal{E}', \mathcal{H}', \text{INCOMPLETE}) \\ \text{then } (\mathcal{E}, \mathcal{H}, \text{COMPLETE}) \vdash_{\text{ri}} (\mathcal{E}', \mathcal{H}', \text{COMPLETE}) \end{array}$$

Essentially, COMPLETENESS allows us to return the completeness flag that was lost due to a POSTULATE or GENERALIZATION step, once we have managed to remove all the added / generalized lemma equations. In practice, a tool or human prover might have a derivation that could be denoted $(\mathcal{E}, \mathcal{H}, \text{COMPLETE}) \vdash_{\text{ri}} (\text{POSTULATE}) (\mathcal{E} \cup \mathcal{E}', \mathcal{H}, \text{INCOMPLETE}) \vdash_{\text{ri}} \cdots \vdash_{\text{ri}} (\mathcal{E}, \mathcal{H} \cup \mathcal{H}', \text{INCOMPLETE}) \vdash_{\text{ri}} (\text{COMPLETENESS}) (\mathcal{E}, \mathcal{H} \cup \mathcal{H}', \text{COMPLETE})$ by remembering the set \mathcal{E} where the completeness flag was lost.

Example 4.25. Recall Ex. 4.14. Starting in $(\{ \text{(FCT.J)} \}, \{ \text{(FCT.D}^{-1}) \}, \text{COMPLETE})$, we lost completeness by adding a lemma equation. Then, after using EXPANSION, we arrived at $(\{ \text{(FCT.J)}, \text{(FCT.N)}, \text{(FCT.O)} \}, \{ \text{(FCT.D}^{-1}), \text{(FCT.M}^{-1}) \}, \text{INCOMPLETE})$. Applying the proof steps of Examples 4.18 and 4.19 without touching (FCT.J), we could reduce this state to $(\{ \text{(FCT.J)} \}, \{ \text{(FCT.D}^{-1}), \text{(FCT.M}^{-1}) \}, \text{INCOMPLETE})$. But the only equation (FCT.J) in this set is the one we started with. Thus, we may restore the completeness flag, resulting in $(\{ \text{(FCT.J)} \}, \{ \text{(FCT.D}^{-1}), \text{(FCT.M}^{-1}) \}, \text{COMPLETE})$.

There are many other potential inference rules we could consider, as various extensions of the base method have been studied in the literature (see e.g. [Bouhoula 1997]). For now, we stick to these nine rules and leave the remainder to future work.

4.3. Examples

The running example in § 4.2 gives a good general idea of the power of the method and the way it is applied. In this section we present some further examples. For brevity, we only list the equations \mathcal{E} in each step, not the completeness flag or induction rules \mathcal{H} . Unless stated otherwise, these induction rules are not applicable to new equations.

Example 4.26. Let us look at an assignment to implement `strlen`, a string function which operates on 0-terminated char arrays. As `char` is a numeric data type, we use integer arrays in the LCTRS translation (although another underlying sort $\mathcal{I}_{\text{char}}$ would make little difference). The example function and its LCTRS translation are as follows:

```

int strlen(char *s) {
  for(int i = 0; ; i++){
    if(s[i] == 0)
      return i;
  }
}

```

$$\begin{array}{ll}
(1) & \text{strlen}(x) \rightarrow u(x, 0) \\
(2) & u(x, i) \rightarrow \text{error} \quad [i < 0 \vee i \geq \text{size}(x)] \\
(3) & u(x, i) \rightarrow \text{return}(i) \quad [0 \leq i < \text{size}(x) \wedge \text{select}(x, i) = 0] \\
(4) & u(x, i) \rightarrow u(x, i + 1) \quad [0 \leq i < \text{size}(x) \wedge \text{select}(x, i) \neq 0]
\end{array}$$

Note that the bounds checks guarantee termination. To see that `strlen` does what we would expect it to do, we want to know that for *valid C strings*, `strlen(a)` returns the first integer i such that $a[i] = 0$. Following § 1.2.4, this corresponds to the equation:

$$\text{(LEN.A)} \quad \text{strlen}(x) \approx \text{return}(n) \quad [0 \leq n < \text{size}(x) \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0) \wedge \text{select}(x, n) = 0]$$

Here, we use bounded quantification, which, as described in § 2.2, can be seen as syntactic sugar for an additional predicate; the underlying LCTRS could, e.g., use a symbol `nonzero` and replace $\forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0)$ by `nonzero(x, n)` in the constraint.

We first use **SIMPLIFICATION** with rule (1), which gives **(LEN.B)**:

$$u(x, 0) \approx \text{return}(n) \quad [0 \leq n < \text{size}(x) \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0) \wedge \text{select}(x, n) = 0]$$

We continue with **EXPANSION**, again on the left-hand side. Since the constraint implies that $0 < \text{size}(x)$, the error case (2) is unsatisfiable, so we delete it, which leaves:

$$\text{(LEN.C)} \quad \text{return}(0) \approx \text{return}(n) \quad [0 \leq n < \text{size}(x) \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0) \wedge \text{select}(x, n) = 0 \wedge 0 \leq 0 < \text{size}(x) \wedge \text{select}(x, 0) = 0]$$

$$\text{(LEN.D)} \quad u(x, 0 + 1) \approx \text{return}(n) \quad [0 \leq n < \text{size}(x) \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0) \wedge \text{select}(x, n) = 0 \wedge 0 \leq 0 < \text{size}(x) \wedge \text{select}(x, 0) \neq 0]$$

As the constraint of **(LEN.C)** implies that $n = 0$, we can remove **(LEN.C)** using **EQ-DELETION** and **DELETION**. **(LEN.D)** is simplified with a calculation:

$$\text{(LEN.E)} \quad u(x, 1) \approx \text{return}(n) \quad [0 \leq n < \text{size}(x) \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0) \wedge \text{select}(x, n) = 0 \wedge 0 < \text{size}(x) \wedge \text{select}(x, 0) \neq 0]$$

Which we expand again (once more skipping the error case due to unsatisfiability):

$$\text{(LEN.F)} \quad \text{return}(1) \approx \text{return}(n) \quad [0 \leq n < \text{size}(x) \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0) \wedge \text{select}(x, n) = 0 \wedge 0 < \text{size}(x) \wedge \text{select}(x, 0) \neq 0 \wedge 0 \leq 1 < \text{size}(x) \wedge \text{select}(x, 1) = 0]$$

$$\text{(LEN.G)} \quad u(x, 1 + 1) \approx \text{return}(n) \quad [0 \leq n < \text{size}(x) \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0) \wedge \text{select}(x, n) = 0 \wedge 0 < \text{size}(x) \wedge \text{select}(x, 0) \neq 0 \wedge 0 \leq 1 < \text{size}(x) \wedge \text{select}(x, 1) \neq 0]$$

The constraint of **(LEN.F)** implies that $n = 1$, so we easily remove this equation. **(LEN.G)** is simplified using a calculation and then expanded again:

$$\text{(LEN.H)} \quad \text{return}(2) \approx \text{return}(n) \quad [\dots \wedge 2 < \text{size}(x) \wedge \text{select}(x, 2) = 0]$$

$$\text{(LEN.I)} \quad u(x, 2 + 1) \approx \text{return}(n) \quad [0 \leq n < \text{size}(x) \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0) \wedge \text{select}(x, n) = 0 \wedge 0 < \text{size}(x) \wedge \text{select}(x, 0) \neq 0 \wedge 1 < \text{size}(x) \wedge \text{select}(x, 1) \neq 0 \wedge 2 < \text{size}(x) \wedge \text{select}(x, 2) \neq 0]$$

We drop (LEN.H) easily. Simplifying (LEN.I) and reformulating its constraint gives:

$$\text{(LEN.J)} \quad u(x, 3) \approx \text{return}(n) \quad [0 \leq n < \text{size}(x) \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0) \wedge \text{select}(x, n) = 0 \wedge 0 \leq 2 < \text{size}(x) \wedge \forall j \in \{0, \dots, 2\}(\text{select}(x, j) \neq 0)]$$

Note that we grouped together the $\neq 0$ statements into a quantification, which looks a lot like the other quantification in the constraint. Now, let us generalize! We will use the generalized equation (LEN.K): $u(x, k) \approx \text{return}(n) [\varphi]$, where:

$$\varphi : k = m + 1 \wedge 0 \leq n < \text{size}(x) \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0) \wedge \text{select}(x, n) = 0 \wedge 0 \leq m < \text{size}(x) \wedge \forall j \in \{0, \dots, m\}(\text{select}(x, j) \neq 0)$$

Obviously, (LEN.J) is an instance of (LEN.K); we use EXPANSION to obtain:

$$\begin{aligned} \text{(LEN.L)} \quad \text{error} &\approx \text{return}(n) && [\varphi \wedge (k < 0 \vee k \geq \text{size}(x))] \\ \text{(LEN.M)} \quad \text{return}(k) &\approx \text{return}(n) && [\varphi \wedge 0 \leq k < \text{size}(x) \wedge \text{select}(x, k) = 0] \\ \text{(LEN.N)} \quad u(x, k+1) &\approx \text{return}(n) && [\varphi \wedge 0 \leq k < \text{size}(x) \wedge \text{select}(x, k) \neq 0] \end{aligned}$$

The two \forall statements in φ , together with $\text{select}(x, n) = 0$, imply that $m < n$, so $k \leq n$. Consequently, (LEN.L) has an unsatisfiable constraint and may be deleted: $k < 0$ cannot hold because $k = m + 1$ and $0 \leq m$, nor $k \geq \text{size}(x)$ because $k \leq n$ and $n < \text{size}(x)$.

For (LEN.M), the two \forall statements together with $\text{select}(x, k) = 0$ imply that $n - 1 < k$, so $n \leq k$. Thus, $n = k$. EQ-DELETION gives an equation with an unsatisfiable constraint, which we remove using DELETION. As for (LEN.N), we use SIMPLIFICATION with a calculation and reformulate the constraint to obtain:

$$\text{(LEN.O)} \quad u(x, p) \approx \text{return}(n) \quad [p = k + 1 \wedge \text{select}(x, n) = 0 \wedge 0 \leq n < \text{size}(x) \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0) \wedge 0 \leq k < \text{size}(x) \wedge \forall j \in \{0, \dots, k\}(\text{select}(x, j) \neq 0) \wedge \text{some constraints on } m]$$

This equation is simplified to an equation of the form $\text{return}(n) \approx \text{return}(n) [\dots]$ using the induction rule obtained from (LEN.K); we complete with DELETION.

Example 4.27. We consider \mathcal{R}_{sum} , the LCTRS with the two correct implementations of the motivating Ex. 1.1; that is, rules (1a)–(1d) and (4a)–(4e). The rules are terminating because in the recursive rule (1c), $n - i$ decreases in every step and is bounded from below by 0, and in rule (4c), the value k decreases against the bound 0.

To prove equivalence of these implementations when the given length is within the array bounds, we must show that (ARR.A) is an inductive theorem:

$$\text{(ARR.A)} \quad \text{sum1}(a, k) \approx \text{sum4}(a, k) \quad [0 \leq k \leq \text{size}(a)]$$

The derivation follows a similar pattern as with factorial: we first simplify the left hand using rule (1a), then expand on the right and use the induction rule, $\text{sum4}(a, k) \rightarrow u(a, k, 0, 0) [0 \leq k \leq \text{size}(a)]$, to eliminate the remaining occurrence of sum4 . This gives:

$$\begin{aligned} &w(n, u(a, k', 0, 0)) \approx u(a, k, r, 1) \\ &[k' = k - 1 \wedge 0 \leq k' < \text{size}(a) \wedge n = \text{select}(a, k') \wedge r = 0 + \text{select}(a, 0)] \end{aligned}$$

Continuing to expand and simplify, we easily remove the equations resulting from rules (1b) and (1d) in every step, but the recursive rule (1c) causes a divergence.

$$\begin{aligned} u(a, k, r_2, 3) &\approx w(n, u(a, k', r_1, 2)) \quad [k' = k - 1 \wedge 2 < k \leq \text{size}(a) \wedge r_2 = r_1 + \text{select}(a, 1) \wedge \dots] \\ u(a, k, r_3, 4) &\approx w(n, u(a, k', r_2, 3)) \quad [k' = k - 1 \wedge 3 < k \leq \text{size}(a) \wedge r_3 = r_2 + \text{select}(a, 2) \wedge \dots] \\ u(a, k, r_4, 5) &\approx w(n, u(a, k', r_3, 4)) \quad [k' = k - 1 \wedge 4 < k \leq \text{size}(a) \wedge r_4 = r_3 + \text{select}(a, 3) \wedge \dots] \end{aligned}$$

We can easily complete after generalizing any of these equations to:

$$\text{(ARR.GEN)}: \quad u(a, k, r, i) \approx w(n, u(a, k', r', i')) \quad [k' = k - 1 \wedge 0 \leq i' < k \leq \text{size}(a) \wedge i' = i - 1 \wedge r = r' + \text{select}(a, i') \wedge n = \text{select}(a, k')]$$

Example 4.28. Recall `strcpy` from Ex. 3.7 and the analysis rules and equation from Ex. 1.2. The inductive proof follows roughly the same lines as the one for `strlen` and is found automatically by our tool (see § 6). We reach a divergence in equations such as:

- $\text{test}(x, n, v(a, x, 1)) \approx \text{true} [0 \leq n < \text{size}(x) \wedge n < \text{size}(a) \wedge \text{select}(x, n) = 0 \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) = 0) \wedge \text{select}(x, 0) \neq 0 \wedge \text{select}(x, 0) = \text{select}(a, 0)]$
- $\text{test}(x, n, v(b, x, 2)) \approx \text{true} [0 \leq n < \text{size}(x) \wedge n < \text{size}(b) \wedge \text{select}(x, n) = 0 \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) = 0) \wedge \text{select}(x, 0) \neq 0 \wedge \text{select}(x, 0) = \text{select}(b, 0) \wedge \text{select}(x, 1) \neq 0 \wedge \text{select}(b, 1) = \text{select}(x, 1)]$
- $\text{test}(x, n, v(c, x, 3)) \approx \text{true} [\dots \wedge \text{select}(c, 2) \neq 0 \wedge \text{select}(c, 2) = \text{select}(x, 2)]$

To generalize, we abstract 1, 2, 3 by $k \geq 0$, collect similar statements into quantifications and remove the endpoint. We quickly complete after this GENERALIZATION to:

$$\text{test}(x, n, v(c, x, k)) \approx \text{true} [0 \leq n < \text{size}(x) \wedge n < \text{size}(c) \wedge \text{select}(x, n) = 0 \wedge 0 \leq k \wedge \forall i \in \{0, \dots, n-1\}(\text{select}(x, i) \neq 0) \wedge \forall i \in \{0, \dots, k-1\}(\text{select}(x, i) \neq 0) \wedge \forall i \in \{0, \dots, k-1\}(\text{select}(c, i) = \text{select}(x, i))]$$

Example 4.29. Let us compare two implementations of the Fibonacci function:

- | | | |
|-----|--|--------------|
| (1) | $\text{fibrec}(x) \rightarrow 0$ | $[x \leq 0]$ |
| (2) | $\text{fibrec}(1) \rightarrow 1$ | |
| (3) | $\text{fibrec}(x) \rightarrow \text{plus}(\text{fibrec}(x-1), \text{fibrec}(x-2))$ | $[x \geq 2]$ |
| (4) | $\text{plus}(\text{return}(x), \text{return}(y)) \rightarrow \text{return}(x+y)$ | |
| (5) | $\text{fibiter}(x) \rightarrow \text{iter}(x, 1, 0, 1)$ | |
| (6) | $\text{iter}(x, i, y, z) \rightarrow \text{iter}(x, i+1, z, y+z)$ | $[x \geq i]$ |
| (7) | $\text{iter}(x, i, y, z) \rightarrow \text{return}(y)$ | $[x < i]$ |

Starting with the equation $\text{fibrec}(x) \approx \text{fibiter}(x)$ [true] eventually results in a divergence:

$$\begin{aligned} \text{iter}(n, 3, 1, 2) &\approx \text{plus}(\text{iter}(m, \text{iter}(m, 2, 1, 1)), \text{iter}(k, \text{iter}(k, 1, 0, 1))) [m = n-1 \wedge k = n-2] \\ \text{iter}(n, 4, 2, 3) &\approx \text{plus}(\text{iter}(m, \text{iter}(m, 3, 1, 2)), \text{iter}(k, \text{iter}(k, 2, 1, 1))) [m = n-1 \wedge k = n-2] \\ \text{iter}(n, 5, 3, 5) &\approx \text{plus}(\text{iter}(m, \text{iter}(m, 4, 2, 3)), \text{iter}(k, \text{iter}(k, 3, 1, 2))) [m = n-1 \wedge k = n-2] \end{aligned}$$

The proof is easily finished by using the following generalization:

$$\begin{aligned} \text{iter}(n_3, i_3, z_3, z_4) &\approx \text{plus}(\text{iter}(n_2, i_2, z_2, z_3), \text{iter}(n_1, i_1, z_1, z_2)) \\ [n_2 = n_3 - 1 \wedge n_1 = n_2 - 2 \wedge i_3 = i_2 + 1 \wedge i_2 = i_1 + 1 \wedge z_3 = z_1 + z_2 \wedge z_4 = z_2 + z_3] \end{aligned}$$

Thus, we can show equivalence of functions with wildly different time complexities (`fibrec`'s running time is exponential in the input value, whereas that of `fibiter` is linear).

Example 4.30. Finally, we consider an example which Godlin and Strichman [2008, § 6, item 2] describe as beyond their method. Here two recursive imperative programs calculating $\sum_{i=1}^n i$ are compared. The methods from § 3 yield the following LCTRS.

- | | | | | | |
|-----|---|--------------|-----|---|--------------|
| (1) | $f(n) \rightarrow \text{return}(n)$ | $[n \leq 0]$ | (4) | $g(n) \rightarrow \text{return}(n)$ | $[n \leq 1]$ |
| (2) | $f(n) \rightarrow u(n, f(n-1))$ | $[n > 0]$ | (5) | $g(n) \rightarrow v(n, g(n-1))$ | $[n > 1]$ |
| (3) | $u(n, \text{return}(m)) \rightarrow \text{return}(n+m)$ | | (6) | $v(n, \text{return}(m)) \rightarrow \text{return}(n+m)$ | |

Starting with the equation $f(x) \approx g(x)$ [true] eventually results in a divergence:

- (CR.A): $u(x, u(y_1, g(y_2))) \approx v(x, u(z_1, g(z_2)))$
 $[x > 1 \wedge y_1 = x-1 \wedge z_1 = x-1 \wedge y_2 = y_1-1 \wedge z_2 = z_1-1]$
- (CR.B): $u(x, u(y_1, u(y_2, g(y_3)))) \approx v(x, u(z_1, u(z_2, g(z_3))))$
 $[x > 1 \wedge y_1 = x-1 \wedge z_1 = x-1 \wedge y_2 = y_1-1 \wedge z_2 = z_1-1 \wedge y_3 = y_2-1 \wedge z_3 = z_2-1]$
- (CR.C): $u(x, u(y_1, u(y_2, u(y_3, g(y_4)))))) \approx v(x, u(z_1, u(z_2, u(z_3, g(z_4)))))) [\dots]$

As the constraints imply that each $y_i = z_i$, these equations can all be generalized to $u(x, u(y, z)) \approx v(x, u(y, z)) [x > 1]$. Again, the proof is quickly completed.

4.4. Soundness and Completeness of Rewriting Induction

We now give an intuition on how to prove Thm. 4.4. The complete proof can be found in Appendix B. We follow the proof method of [Sakata et al. 2009], which builds on the original proof idea in [Reddy 1990]. This uses the relation $\leftrightarrow_{\mathcal{E}}$, defined by

$$C[s\gamma]_p \leftrightarrow_{\mathcal{E}} C[t\gamma]_p \quad \text{if } s \approx t \text{ } [\varphi] \in \mathcal{E} \text{ or } t \approx s \text{ } [\varphi] \in \mathcal{E}, \text{ and } \gamma \text{ respects } \varphi$$

for \mathcal{E} a set of equations. The proof is split up into several auxiliary lemmas. To start:

LEMMA 4.31. *All equations in \mathcal{E} are inductive theorems if and only if $\leftrightarrow_{\mathcal{E}} \subseteq \leftrightarrow_{\mathcal{R}}^*$ on ground terms (so if s, t are ground and $s \leftrightarrow_{\mathcal{E}} t$, then also $s \leftrightarrow_{\mathcal{R}}^* t$).*

This is obvious from the definitions. The next lemma originates in [Sakata et al. 2009], which is adapted from [Koike and Toyama 2000] and is key to our method.

LEMMA 4.32 ([SAKATA ET AL. 2009]). *Let \rightarrow_1 and \rightarrow_2 be binary relations. We have $\leftrightarrow_1^* = \leftrightarrow_2^*$ if (a) $\rightarrow_1 \subseteq \rightarrow_2$, (b) \rightarrow_2 is well founded, and (c) $\rightarrow_2 \subseteq (\rightarrow_1 \cdot \rightarrow_2^* \cdot \leftrightarrow_1^* \cdot \leftarrow_2^*)$.*

PROOF. It follows from $\rightarrow_1 \subseteq \rightarrow_2$ that $\leftrightarrow_1^* \subseteq \leftrightarrow_2^*$. To show that $\leftrightarrow_2^* \subseteq \leftrightarrow_1^*$, we prove $\rightarrow_2^* \subseteq \leftrightarrow_1^*$ by well-founded induction on \rightarrow_2 . Since the base case $s \rightarrow_2^* s$ is clear, we suppose $s \rightarrow_2 t \rightarrow_2^* u$. As $\rightarrow_2 \subseteq (\rightarrow_1 \cdot \rightarrow_2^* \cdot \leftrightarrow_1^* \cdot \leftarrow_2^*)$ there must be some a, b, c such that $s \rightarrow_1 a \rightarrow_2^* b \leftrightarrow_1^* c \leftarrow_2^* t$. Since $\rightarrow_1 \subseteq \rightarrow_2$ (i.e., $s \rightarrow_2 a$), we can apply the induction hypothesis both on a and on t , so $a \leftrightarrow_1^* b \leftrightarrow_1^* c \leftrightarrow_1^* t$ and $t \leftrightarrow_1^* u$. Therefore, $s \leftrightarrow_1^* u$. \square

We will use Lemma 4.32 with $\rightarrow_{\mathcal{R}}$ for \rightarrow_1 , and $\rightarrow_{\mathcal{R}\cup\mathcal{H}}$ for \rightarrow_2 . Soundness of the algorithm then follows if $\leftrightarrow_{\mathcal{E}}$ is included in $\leftrightarrow_{\mathcal{H}}^*$ whenever $(\mathcal{E}, \emptyset, \text{flag}) \vdash_{\text{ri}}^* (\emptyset, \mathcal{H}, \text{flag}')$.

Thm. 4.4 is the combination of Lemma 4.31 with Lemmas 4.33 and 4.34 below.

LEMMA 4.33. *If $(\mathcal{E}, \emptyset, \text{flag}) \vdash_{\text{ri}}^* (\emptyset, \mathcal{H}, \text{flag}')$, then $\leftrightarrow_{\mathcal{E}} \subseteq \leftrightarrow_{\mathcal{R}}$ holds on ground terms.*

PROOF IDEA. Let $\leftrightarrow_{\mathcal{E}}$ denote a parallel application of zero or more $\leftrightarrow_{\mathcal{E}}$ steps. We first show that $(\mathcal{E}, \mathcal{H}, \text{flag}) \vdash_{\text{ri}} (\mathcal{E}', \mathcal{H}', \text{flag}')$ by any rule other than COMPLETENESS implies both (a) $\leftrightarrow_{\mathcal{E}} \subseteq (\rightarrow_{\mathcal{R}\cup\mathcal{H}}^* \cdot \leftrightarrow_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R}\cup\mathcal{H}}^*)$ on ground terms, and (b) if $\rightarrow_{\mathcal{R}\cup\mathcal{H}} \subseteq (\rightarrow_{\mathcal{R}} \cdot \rightarrow_{\mathcal{R}\cup\mathcal{H}}^* \cdot \leftrightarrow_{\mathcal{E}} \cdot \leftarrow_{\mathcal{R}\cup\mathcal{H}}^*)$ on ground terms, then $\rightarrow_{\mathcal{R}\cup\mathcal{H}} \subseteq (\rightarrow_{\mathcal{R}} \cdot \rightarrow_{\mathcal{R}\cup\mathcal{H}}^* \cdot \leftrightarrow_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R}\cup\mathcal{H}}^*)$ on ground terms. We show this by considering how each step alters \mathcal{E} and \mathcal{H} , which we use to see that $(\mathcal{E}, \mathcal{H}, \text{flag}) \vdash_{\text{ri}}^* (\mathcal{E}', \mathcal{H}', \text{flag}')$ implies (a) and (b), by induction on the total number of \vdash_{ri} -steps in the derivation (counting also the hidden steps inside COMPLETENESS). Thus, if $(\mathcal{E}, \emptyset, \text{flag}) \vdash_{\text{ri}}^* (\emptyset, \mathcal{H}, \text{flag}')$ then $\rightarrow_{\mathcal{R}\cup\mathcal{H}} \subseteq \rightarrow_{\mathcal{R}} \cdot \rightarrow_{\mathcal{R}\cup\mathcal{H}}^* \cdot \leftarrow_{\mathcal{R}\cup\mathcal{H}}^*$, so we can apply Lemma 4.32 to conclude that $\rightarrow_{\mathcal{R}}$ and $\rightarrow_{\mathcal{R}\cup\mathcal{H}}$ are the same (on ground terms). Therefore, and by property (a), $\leftrightarrow_{\mathcal{E}} \subseteq \leftrightarrow_{\mathcal{R}}$. \square

LEMMA 4.34. *If \mathcal{R} is confluent and $(\mathcal{E}, \emptyset, \text{COMPLETE}) \vdash_{\text{ri}}^* \perp$, then $\leftrightarrow_{\mathcal{E}} \not\subseteq \leftrightarrow_{\mathcal{R}}$ holds on ground terms.*

PROOF IDEA. By confluence and termination together, we can speak of the normal form $u \downarrow_{\mathcal{R}}$ of any term u ; if u is ground, then by quasi-reductivity its normal form is a ground constructor term. A property of confluence is that if $w \leftrightarrow_{\mathcal{R}}^* q$, then $w \downarrow_{\mathcal{R}} = q \downarrow_{\mathcal{R}}$. So, it suffices to prove that for some $s \approx t \text{ } [\varphi] \in \mathcal{E}$ there is a ground constructor substitution γ which respects this equation, such that $s\gamma \neq t\gamma$. We first note that if $(\mathcal{E}, \mathcal{H}, \text{COMPLETE}) \vdash_{\text{ri}}^* \perp$, then this can only be a DISPROVE step; in all cases the equation that causes the disproof has this property. We also see, by examining the various inference rules, that if $(\mathcal{E}_1, \mathcal{H}_1, \text{COMPLETE}) \vdash_{\text{ri}} (\mathcal{E}_2, \mathcal{H}_2, \text{COMPLETE})$ and both (a) $\rightarrow_{\mathcal{R}\cup\mathcal{H}_1} \subseteq \rightarrow_{\mathcal{R}} \cdot \rightarrow_{\mathcal{R}\cup\mathcal{H}_1}^* \cdot \leftrightarrow_{\mathcal{E}_1} \cdot \leftarrow_{\mathcal{R}\cup\mathcal{H}_1}^*$ and (b) $\leftrightarrow_{\mathcal{E}_1} \cup \leftrightarrow_{\mathcal{H}_1} \subseteq \leftrightarrow_{\mathcal{R}}^*$ on ground terms, then also $\leftrightarrow_{\mathcal{E}_2} \cup \leftrightarrow_{\mathcal{H}_2} \subseteq \leftrightarrow_{\mathcal{R}}^*$ on ground terms. In a reduction $(\mathcal{E}, \emptyset, \text{COMPLETE}) = (\mathcal{E}_1, \mathcal{H}_1, \text{flag}_1) \vdash_{\text{ri}} \dots \vdash_{\text{ri}} (\mathcal{E}_n, \mathcal{H}_n, \text{flag}_n) \vdash_{\text{ri}}^* \perp$, we may assume (a) by the observations in the proof of Lemma 4.33, and (b) is inductively preserved. As $\leftrightarrow_{\mathcal{E}_n \cup \mathcal{H}_n}$ cannot be included in $\leftrightarrow_{\mathcal{R}}^*$, therefore neither can $\leftrightarrow_{\mathcal{E}} = \leftrightarrow_{\mathcal{E}_1 \cup \mathcal{H}_1}$. We complete by Lemma 4.31. \square

5. GENERALIZING EQUATIONS

Divergence, as encountered in all examples in § 4, is very common in inductive theorem proving: we often need a more general claim to obtain a stronger induction hypothesis. As it is not always easy to find a suitable generalization, the (automatic) generation of suitable generalizations, and lemma equations for POSTULATE, has been extensively investigated [Bundy et al. 2005; Kapur and Sakhanenko 2003; Kapur and Subramaniam 1996; Nakabayashi et al. 2010; Urso and Kounalis 2004; Walsh 1996].

Also for transformed procedural programs, we will certainly need a large variety of lemma generation techniques to handle most practical cases. We start the work by proposing two methods to generalize equations, specialized to deal with constraints.

5.1. Generalizing Initializations

Our first and most important technique fundamentally relies on the constrained setting. Although it may appear deceptively simple (at its core, the generalization just drops a part of the constraint), it is particularly effective for dealing with loops.

Example 5.1. Let us state the rules of $\mathcal{R}_{\text{fact}}$ from Ex. 4.2 in an alternative way: we replace rule (1) $\text{factiter}(x) \rightarrow \text{iterm}(x, 1, 1)$ by (1'): $\text{factiter}(x) \rightarrow \text{iter}(x, v_1, v_2) [v_1 = 1 \wedge v_2 = 1]$. That is, the values corresponding to *initializations* $\text{int } z = 1; \text{int } i = 1;$ are moved into the constraint. Evidently, this change does not alter the relation $\rightarrow_{\mathcal{R}}$.

Now consider what happens if we use the same steps as in Ex. 4.2–4.14. The resulting proof has the same shape, but with more complex equations. Some instances:

$$\begin{aligned} (\text{FCT.B}') : \text{iter}(n, v_1, v_2) &\approx \text{factrec}(n) [n \geq 1 \wedge v_1 = 1 \wedge v_2 = 1] \\ (\text{FCT.D}') : \text{iter}(n, z_1, i_1) &\approx \text{factrec}(n) [n \geq 1 \wedge v_1 = 1 \wedge v_2 = 1 \wedge z_1 = v_1 * v_2 \wedge i_1 = v_2 + 1] \\ (\text{FCT.J}') : \text{mul}(n, \text{iter}(m, z_1, i_1)) &\approx \text{iter}(n, z_2, i_2) [n > 1 \wedge v_1 = 1 \wedge v_2 = 1 \wedge m = n - 1 \wedge \\ & z_1 = v_1 * v_2 \wedge i_1 = v_2 + 1 \wedge z_2 = z_1 * i_1 \wedge i_2 = i_1 + 1] \\ (\text{FCT.K}') : \text{mul}(n, \text{iter}(m, z_2, i_2)) &\approx \text{iter}(n, z_3, i_3) [n > 1 \wedge v_1 = 1 \wedge v_2 = 1 \wedge m = n - 1 \wedge \\ & z_1 = v_1 * v_2 \wedge i_1 = v_2 + 1 \wedge z_2 = z_1 * i_1 \wedge i_2 = i_1 + 1 \wedge z_3 = z_2 * i_2 \wedge i_3 = i_2 + 1] \end{aligned}$$

Here the left- and right-hand side of the divergent equations (FCT.J') and (FCT.K') are the same modulo variable renaming, while the constraint grows. Essentially, we keep track of parts of the history of an equation in its constraint. We generalize (FCT.J') by dropping all clauses $v_i = q_i$ where v_i is an initialization variable and q_i a value. We rename the variables v_i (as they no longer play a special role) and obtain:

$$(\text{FCT.M}') : \text{mul}(n, \text{iter}(m, z_1, i_1)) \approx \text{iter}(n, z_2, i_2) [n > 1 \wedge m = n - 1 \wedge z_1 = x_1 * x_2 \wedge i_1 = x_2 + 1 \wedge z_2 = z_1 * i_1 \wedge i_2 = i_1 + 1]$$

We can complete the derivation with (FCT.M') as we did with (FCT.M) before.

Formally, what we do here is threefold. First, we alter the set of rules we work from.

Definition 5.2 (Initialization-free Rules). Given \mathcal{R} , fix a set $\mathcal{V}_{\text{init}} \subsetneq \mathcal{V}$ of variables not occurring in \mathcal{R} . The *initialization-free* counterpart \mathcal{R}' of \mathcal{R} is obtained by stepwise replacing any rule $\ell \rightarrow C[f(r_1, \dots, r_i, \dots, r_n)] [\varphi]$ with $f \in \mathcal{D}$ and r_i a value by $\ell \rightarrow C[f(r_1, \dots, v, \dots, r_n)] [\varphi \wedge v = r_i]$ for some fresh $v \in \mathcal{V}_{\text{init}}$, until no such rules remain.

Then, to apply GENERALIZATION to an equation $s \approx t [\varphi_1 \wedge \dots \wedge \varphi_n]$ we choose

$$s \approx t [\bigwedge \{\varphi_i \mid 1 \leq i \leq n \wedge \varphi_i \text{ does not have the form } v = u \text{ with } v \in \mathcal{V}_{\text{init}} \text{ and } u \in \text{Val}\}]$$

as the generalized equation and rename its variables in $\mathcal{V}_{\text{init}}$ to variables in \mathcal{V} .

Finally, we restrict the SIMPLIFICATION and EXPANSION steps to preserve initialization constraints throughout the proof. The strategy we use in Ctrl—which includes an approach to handle the $v \in \mathcal{V}_{\text{init}}$ —is described in § 6.1, but in particular:

- When we rename rules for use in SIMPLIFICATION or EXPANSION, the renaming must respect membership in $\mathcal{V}_{\text{init}}$, i.e., if x is renamed to y , then $y \in \mathcal{V}_{\text{init}}$ iff $x \in \mathcal{V}_{\text{init}}$.
- In \sim -steps, any conjuncts $v = n$ are ignored: to simplify $s \approx t [\varphi \wedge v_1 = n_1 \wedge \dots \wedge v_k = n_k]$, we modify $s \approx t [\varphi]$, obtaining $s' \approx t' [\varphi']$, and continue with $s' \approx t' [\varphi \wedge v_1 = n_1 \wedge \dots \wedge v_k = n_k]$. Thus we avoid, e.g., translating $f(v_i) [v_i = 0]$ back to $f(0) [\text{true}]$.

5.2. Abstracting Equivalent Recursive Calls

Our second generalization technique aims to remove *recursive* symbols where possible.

Definition 5.3. For symbols f, g , let $f \rightsquigarrow g$ if there is a rule $f(\vec{\ell}) \rightarrow r [\varphi]$ with g a symbol in r . A symbol f is *recursive* if it is a defined symbol with $f \rightsquigarrow^+ f$.

The key idea is to identify equivalent occurrences of a recursive call on both sides of an equation and to replace them by a variable. For example, $g(x) + f(y) \approx f(z) + g(x) [y \geq z \wedge y \leq z]$ is replaced by $a + b \approx b + a [\text{true}]$ because for values k, n, m : if $n \geq m \wedge m \leq n$ holds, then both $g(k)$ and $g(m)$, as well as $f(n)$ and $f(m)$, are syntactically equal.

Definition 5.4. A *recursion-abstraction* of $s \approx t [\varphi]$ is any equation of the form $C[x_1, \dots, x_n] \approx D[x_{i_1}, \dots, x_{i_n}]$ such that (a) $s = C[s_1, \dots, s_n]$ and $t = D[t_{i_1}, \dots, t_{i_n}]$ for some \vec{s}, \vec{t} ; (b) $\{i_j \mid 1 \leq j \leq n\} = \{1, \dots, n\}$; (c) neither C nor D contain recursive symbols; (d) each s_j and t_j has a recursive symbol as root symbol; (e) for $1 \leq i \leq n$ and all ground substitutions γ which respect $s \approx t [\varphi]$: $s_i \gamma = t_i \gamma$; (f) x_1, \dots, x_n are fresh w.r.t. s, t .

For a given equation, at most one choice of C, D is possible, and there are only finitely many permutations i_1, \dots, i_n . Requirement (e) can be checked by confirming that an equation $s_j \approx t_j [\varphi]$ is removed by the combination of EQ-DELETION and DELETION.

Example 5.5. In Ex. 4.30, we find an abstraction for (CR.A) by choosing $C = u(x, u(y_1, \square))$, $D = v(x, u(z_1, \square))$, $s_1 = g(y_2)$ and $t_1 = g(z_2)$. Requirement (e) holds: if we write φ for the constraint of (CR.A), EQ-DELETION on $g(y_2) \approx g(z_2) [\varphi]$ produces the unsatisfiable constraint $\varphi \wedge y_2 \neq z_2$. Thus, we generalize the equation to $u(x, u(y_1, a)) \approx v(x, u(z_1, a)) [\varphi]$, which is \sim -equivalent to the equation used in Ex. 4.30.

Example 5.6. Given $g(x) + f(y) \approx f(z) + g(x) [y \geq z \wedge y \leq z]$, let C and D be $\square + \square$, $s_1 = g(x)$, $s_2 = f(y)$, $t_1 = g(x)$, $t_2 = f(z)$, $i_1 = 2$ and $i_2 = 1$. We must see that for all γ which respect $y \geq z \wedge y \leq z$: $g(x)\gamma = g(x)\gamma$ and $f(y)\gamma = f(z)\gamma$. Both are easily confirmed, so we generalize to $x_1 + x_2 \approx x_2 + x_1 [y \geq z \wedge y \leq z] \sim a + b \approx b + a [\text{true}]$ as suggested.

One can see this generalization heuristic as an instance of the inference rule SPECIALIZATION by Aubin [1979] for unconstrained explicit induction; restricted to recursive function calls and combined with SUBSTITUTIVITY OF EQUALITY from the same paper. Here we lift equality from syntactic level to semantic level in SMT-theories.

5.3. Discussion

The first method to generalize equations is strong (§ 5.1), but only for equations of a specific form: we can only use the method if the equation part of the divergence has the same shape every time. This is the case for fact, because the rule that causes the divergence has the form $\text{iter}(x_1, \dots, x_n) \rightarrow \text{iter}(r_1, \dots, r_n) [\varphi]$, preserving its outer shape.

In general, the method is most likely to be successful for the analysis of tail-recursive functions (with accumulators), such as those obtained from procedural programs. We can also handle mutually recursive functions, like $u(x_1, \dots, x_n) \rightarrow w(r_1, \dots, r_m) [\varphi]$ and $w(y_1, \dots, y_m) \rightarrow u(q_1, \dots, q_n) [\psi]$. It is not suitable for analyzing systems with (only) non-tail-recursion, however. Here, the second technique comes in (§ 5.2). Although we do not claim that this technique is very powerful, it is often useful to eliminate apparently simple equations. It is also straightforward to use in practice.

Note that `strlen` and `strcpy` also have the required tail-recursive form to successfully use the first generalization method. However, here we additionally have to collect multiple clauses into a quantification before generalizing, as with equation (LEN.I).

One may wonder if generalizing initializations loses too much; e.g., when removing $v_i = 1$, we also forget that $v_i \geq 0$. However, this is usually not an issue: if a rule is constrained with $v_i \geq 0$, this clause is added to the constraint of the equation via EXPANSION before we generalize, as in the expansion from (LEN.B). There *is* a possible issue with losing information on the relations *between* variables; more on this in § 6.2.

6. IMPLEMENTATION

The method for program verification in this paper can be broken down into two parts:

- (1) transforming a procedural program into an LCTRS;
- (2) proving correctness properties on this LCTRS using rewriting induction.

An initial implementation of part 1, limited to functions on integers and one-dimensional statically allocated integer arrays is available at:

<http://www.trs.css.i.nagoya-u.ac.jp/c2lctrs/>

In future work, we hope to extend this implementation to include the remaining features discussed in § 3 and Appendix A.2 such as floating points and explicit pointers.

Part 2, the core method on LCTRSs, has been implemented in our tool Ctrl [Kop and Nishida 2015], along with basic techniques to verify termination, confluence and quasi-reductivity. To handle constraints, the tool is coupled both with a small internal reasoner and the external SMT solver Z3 [de Moura and Bjørner 2008]. Z3 is equipped to prove *unsatisfiability* as well as satisfiability, which is essential for testing *validity*.

The internal reasoner serves to detect satisfiability or validity of simple statements quickly, without a call to an SMT solver, and to preprocess certain kinds of queries which arise often (e.g., for termination proving by polynomial interpretations, we preprocess queries with $\exists\forall$ -quantifier prefix to \exists -queries). The reasoner is also used to simplify the constraints of equations, by for instance combining statements into quantifications (which is an essential part of the derivations for functions like `strlen` or `strcpy`).

We also translate our array formulas into the SMT-LIB array format as discussed in § 3.6, encoding an array as a function from \mathbb{Z} to \mathbb{Z} with a second variable for its size.

The latest version of Ctrl (tool paper: [Kop and Nishida 2015]) can be downloaded at:

<http://cl-informatik.uibk.ac.at/software/ctrl/>

6.1. Strategy

Let us discuss the various choices made during a derivation with rewriting induction.

6.1.1. What inference rule to apply. Ctrl always selects the first rule (combination) from:

- (1) EQ-DELETION (if applicable) immediately followed by DELETION;
- (2) DISPROVE, but without the limitation to COMPLETE proof states;
- (3) CONSTRUCTOR;
- (4) SIMPLIFICATION;
- (5) a limited form of EXPANSION;
- (6) GENERALIZATION using a recursion-abstraction;
- (7) GENERALIZATION of all initialization variables $v_i \in \mathcal{V}_{\text{init}}$ at once;
- (8) the full form of EXPANSION.

6.1.2. Generalization and backtracking. Core to the rewriting induction process is a backtracking mechanism. Every proof state $(\mathcal{E}, \mathcal{H})$ keeps track of all ancestor states on

which GENERALIZATION was applied; a state is COMPLETE if it has no such ancestors. The completeness restriction on DISPROVE is dropped; however, when DISPROVE succeeds on an incomplete state, the prover does not conclude failure, but instead backtracks to the most recent ancestor and continues without (immediately) generalizing. Typically, if a GENERALIZATION is attempted too soon in the proof and results in an unsound equation, this can be derived very quickly, which allows Ctrl to conclude failure of the GENERALIZATION step and to move on to the remaining expansions.

Example 6.1. Following Ex. 4.26 (but altered with initialization-free rules), our strategy moves from $(\{(\text{LEN.A}')\}, \emptyset)$ to $(\{(\text{LEN.B}')\}, \emptyset)$ as before. But here, “restricted expansion” does not apply (as we will see in Ex. 6.3), nor is there a recursion-abstraction. So we generalize the initializations, obtaining:

$$\left(\left\{ \begin{array}{l} \text{(BGEN)} \\ [0 \leq n < \text{size}(x) \wedge \forall i \in \{0, \dots, n-1\} (\text{select}(x, i) \neq 0) \wedge \text{select}(x, n) = 0] \end{array} \right\}, \emptyset \right)$$

We store $(\{(\text{LEN.B}')\}, \emptyset)$ as an ancestor state of $(\{(\text{BGEN})\}, \emptyset)$. The only option now is EXPANSION. Expanding in the left-hand side gives three equations, including:

$$\begin{array}{l} \text{return}(r_0) \approx \text{return}(n) \\ [0 \leq n < \text{size}(x) \wedge \forall i \in \{0, \dots, n-1\} (\text{select}(x, i) \neq 0) \wedge \text{select}(x, n) = 0] \end{array}$$

CONSTRUCTOR gives $r_0 \approx n [\varphi]$, where φ is satisfied by, e.g., $[r_0 := 0, n := 1, x := [1, 0]]$; by DISPROVE, we obtain \perp . However, the state is incomplete as it has an ancestor stored. Thus, we backtrack to $(\{(\text{LEN.B}')\}, \emptyset)$, and continue with full expansion.

The COMPLETENESS rule is implemented via the same mechanism: if $(\mathcal{E}, \mathcal{H})$ has a most recent ancestor $(\mathcal{E}', \mathcal{H}')$ with $\mathcal{E} \subseteq \mathcal{E}'$, then $(\mathcal{E}', \mathcal{H}')$ is dropped from the ancestor list. If a DISPROVE succeeds when the list is empty, we conclude failure, resulting in NO if the system is confluent and MAYBE otherwise.

Example 6.2. In Ex. 4.16, we would add $(\{(\text{FCT.J})\}, \{(\text{FCT.D}^{-1})\})$ to the list of ancestors when generalizing (FCT.J) to (FCT.M). Once (FCT.T) is removed in Ex. 4.19, we are allowed to remove this state from the list (although since the proof is finished at that point, it is not really necessary in this example).

Aside from backtracking due to DISPROVE, there is a second backtracking mechanism: although SIMPLIFICATION and EXPANSION prioritize choices (for positions and rules) most likely to result in success, sometimes the first choice does not work out, but the second one does. Thus, Ctrl uses an evaluation limit: when a path has more than N expansions, it is aborted, and the prover backtracks to a direct parent. Ctrl starts with $N = 2$ and increases this limit if it does not result in a successful proof or disproof.

6.1.3. Simplification. For SIMPLIFICATION, there are three choices to be made: the position, the rule and how to instantiate fresh variables in that rule.

For the position, Ctrl selects the leftmost, innermost position where a rule matches. This prevents a need to reevaluate a term after its subterms change.

For the rule, rules in \mathcal{H} are attempted before rules in \mathcal{R} ; if a rule leads to a (presumed) divergence, the backtracking mechanism ensures that the next one is tried.

In some cases—in particular for induction rules—the right-hand side and perhaps the constraint of a rule contain variables not occurring in the left-hand side, such as (FCT.M⁻¹) in Ex. 4.14 and (LEN.K) in Ex. 4.26. Here, Ctrl tries to instantiate as many variables in the rule by variables in the equation $\ell \rightarrow r [\psi_1 \wedge \dots \wedge \psi_m]$, we first determine a γ such that $s = \ell\gamma$ and $\gamma(v_i) = v_i$ for all $v_i \in \mathcal{V}_{\text{init}}$. If any ψ_i has the form $C[x, y_1, \dots, y_k]$ with $x \in \text{Dom}(\gamma)$ and all $y_i \notin \text{Dom}(\gamma)$, and there is some $\varphi_j = C\gamma[\gamma(x), s_1, \dots, s_k]$, then

we extend γ with $[y_i := s_i]$ for all i . This process is finite and corresponds to the choices for the equations (FCT.S) and (LEN.O). Other variables are chosen fresh.

Note: if some rule *can* be applied, but the backtracking mechanism aborts all attempts, Ctrl backtracks to the parent state rather than continuing with EXPANSION. This is because testing suggests that allowing EXPANSION to be applied on terms not in \mathcal{R} -normal form is generally not effective and causes an explosive number of states.

6.1.4. Expansion. To categorize EXPANSIONS for step (5) and (8) of § 6.1.1, we analyze *recursion*. Let $f \succsim g$ if $f \rightsquigarrow^* g$ (following Def. 5.3), and let $f \succ g$ if $f \succsim g$ and $g \not\prec f$. Symbols are split into five categories: *constructors*, *calculation symbols*, *non-recursive defined symbols*, *tail-recursive symbols*, and *non-tail-recursive symbols*. A recursive symbol is *tail-recursive* if its only defining rules (in \mathcal{R}) have either the form $f(\ell_1, \dots, \ell_k) \rightarrow x[\varphi]$ with x a variable, or the form $f(\ell_1, \dots, \ell_k) \rightarrow g(r_1, \dots, r_m)[\varphi]$ with $f \succ h$ for all h in any r_i . Recursive functions not of this form are *non-tail-recursive*.

An expansion of $s \simeq t[\varphi]$ at position p of s , with $s|_p = f(\vec{u})$, is *restricted*—so eligible for step (5)—if (a) f is non-recursive, or (b) the induction rule $s \rightarrow t[\varphi]$ is admissible and either f is tail-recursive and $(\text{Var}(s) \cup \text{Var}(t)) \cap \mathcal{V}_{\text{init}} = \emptyset$, or f is non-tail-recursive. The induction rule is added only in case (b). Here, a rule $\rho: g(\ell_1, \dots, \ell_k) \rightarrow r[\varphi]$ is *admissible* if $\mathcal{R} \cup \mathcal{H} \cup \{\rho\}$ is terminating and $g \in \mathcal{D}$: we do not add rules with a constructor or calculation symbol as root symbol g , as this makes it harder to prove termination, which may prevent the addition of more promising rules later on.

For *unrestricted* expansion, an induction rule is added when admissible, unless f is tail-recursive. The unrestricted tail-recursive case concerns rules such as those got from (FCT.J), (LEN.B), and (LEN.E), which—testing suggests—are typically not useful. Omitting them lets Ctrl skip many termination checks, a bottleneck in the process. Similarly, we do not add induction rules when expanding at a non-recursive position.

Example 6.3. In Examples 4.2–4.18, the first expansion occurs in $(\{(FCT.D')\}, \emptyset, \text{COMPLETE})$, in the right-hand side. This is not an arbitrary choice: restricted expansion cannot be used with the tail-recursive symbol *iter*, only the non-tail-recursive symbol *factrec*. Then, our strategy closely follows the given derivation. When we reach $(\{(FCT.J)\}, \{(FCT.D^{-1})\}, \text{COMPLETE})$, restricted expansion is impossible, so we generalize instead. After this, an expansion on the *iter* symbol on either side is restricted. We can complete the example without backtracking or using unrestricted EXPANSION.

For the position to expand at, we follow the same approach as for SIMPLIFICATION, trying all suitable positions via the backtracking mechanism. However, rather than a pure leftmost innermost choice, in the restricted case (step (5) of § 6.1.1), we prioritize the more promising equations by first attempting expansions on a non-tail-recursive symbol, then those with a non-recursive defined symbol, and finally those with a tail-recursive one. In the unrestricted setting, we follow the leftmost innermost strategy.

Testing shows that this method is very effective for proving equivalence between a non-tail-recursive and a tail-recursive function (as needed for equivalence of a recursive and an iterative C function). The examples of § 4 show its effect: by eliminating the non-tail-recursive functions early on, we are more likely to arrive at a diverging sequence where all equations have the same outer shape; e.g., $(u(\vec{q}_i) \approx C[u(\vec{v}_i), u(\vec{w}_i)] \mid i \in \mathbb{N})$. As observed in § 5.3, this is ideal for our generalization method.

Following an EXPANSION, we first process those new equations in $\text{Expd}(s \approx t[\varphi], p)$ whose multiset of new symbols is smallest in the recursion order \succ . Thus, for example in Ex. 4.10, after expanding (FCT.D) we consider (FCT.E)—which has new symbols $\{\text{return}, 1\}$ —before (FCT.F)—with new symbols $\{\text{mul}, \text{factrec}, -, 1\}$ —since *factrec* \succ *return*, 1. Intuitively, “smaller” terms are “closer” to the end of a function, which allows DISPROVE to succeed faster and thus aids the backtracking mechanism.

6.1.5. Constraint Modification. Following SIMPLIFICATION and EXPANSION, Ctrl modifies the constraint, as follows. First, when a clause φ_i in the constraint $\varphi_1 \wedge \dots \wedge \varphi_n$ is implied by the others, it is removed unless it is a definition clause $v_i = n$. We also remove clauses for variables which do not play a role. Most importantly, Ctrl introduces *ranged quantifications* $\forall x \in \{k_1, \dots, k_n\}(\varphi(x))$ whenever possible, provided $n \geq 3$ (to lessen the effect of coincidence). Formally, we could describe our approach as follows:

if φ has clauses $C[a], C[b], C[c]$ for some context C and variables a, b, c , as well as $b = f(a)$ and $c = f(b)$, then we may replace the C -clauses by $\forall i \in \{0, \dots, 2\}(C[f^i(a)])$

This is more general than what we use; it lets us for instance replace $a[i] = 0 \wedge a[j] = 0 \wedge a[k] = 0 \wedge j = i + 2 \wedge k = j + 2$ by $\forall l \in \{0, \dots, 2\}(a[i + 2 \cdot l] = 0)$, for $f = \lambda x.x + 2$. But to represent f^i , Ctrl must know the relevant theory. Therefore, we currently only consider clauses where $b = a + 1$ and $c = b + 1$, and replace them by $\forall i \in \{a, \dots, c\}(C[i])$. Since we implement loop counters as integers, this still captures a large group of constraints.

After \forall -introduction, if a boundary of the range (0 and 2 in the example) is some $v_i \in \mathcal{V}_{\text{init}}$, we replace it by the value it is defined as, to avoid generalizing the starting point of a quantification. Thus, e.g., $\forall j \in \{v_0, \dots, k\}(\text{select}(x, j) \neq 0) \wedge v_0 = 0 \wedge i = v_0 + 1 \wedge k = i + 1$ is replaced by $\forall j \in \{0, \dots, k\}(\text{select}(x, j) \neq 0) \wedge v_0 = 0 \wedge i = v_0 + 1 \wedge k = i + 1$.

6.1.6. Non-Confluence. Our strategy is admittedly unfair to non-confluent systems: a successful application of DISPROVE is treated as evidence of an unsound equation, which is not the case without confluence: the non-confluent (LC)TRSs $\mathcal{R} = \{f \rightarrow a, f \rightarrow b, g \rightarrow a, g \rightarrow b\}$ along with the inductive theorem $f \approx g$ highlights that we only have to prove that two functions *can* produce the same result, not that they *always* do.

This is deliberate: when proving that two functions produce the same result, we can see non-confluent LCTRSs as inherently incorrect. Thus, we focus on confluent systems. For LCTRSs whose confluence is unknown, it is preferable to show non-equivalence (which translates to a MAYBE in the output) over equivalence.

6.2. Experiments

To assess performance and precision of Ctrl empirically, we tested five assignments from a group of students in the first-year programming course in Nagoya, all automatically translated to LCTRSs by `c2lctrs`: `sum`: given n , implement $\sum_{i=1}^n i$; `fib`: compute the n th Fibonacci number; `sumfrom`: given n, m , implement $\sum_{i=n}^m i$; `strlen` and `strcpy`. We compared the first three to LCTRS-versions of recursive reference implementations,⁶ for `strlen` and `strcpy` we used a specification as in Ex. 4.26 and 4.28.⁷ We also tested our own implementations of `fact` from Ex. 4.2 and `arrsum` from Ex. 4.27, along with 25 function comparisons from the literature and 12 memory-safety benchmarks from the *Competition on Software Verification* [SV-COMP]. The benchmarks (also from the literature) are typically fairly small: the largest, `lit03_GS13_fig6`, has 70 lines of C code and 55 rewrite rules. We used an Intel i7-5600U CPU at 2.6 GHz under Linux.

We quickly found that many of the student programs had failed to account for boundary conditions, such as empty strings or negative input. This causes a NO, or a MAYBE if the system cannot be proved confluent, so if not all variables are initialized. To limit the impact of these errors, we did a second test, where we altered the specification to account for these mistakes. The results of both tests are summarized in Figure 1.

⁶However, honesty compels us to mention that for `fib`, we used a manual translation because the one obtained from `c2lctrs` was impractical: where our manual translation has a rule `fibrec(x) \rightarrow plus(fibrec(x - 1), fibrec(x - 2))` [$x \geq 2$], the automatic one splits the two recursive calls (recall § 3.5). Therefore, a more sophisticated termination argument is needed, and it is harder to eliminate the recursion in the inductive process. Handling such cases in the future will likely necessitate an additional lemma generation technique.

⁷Interestingly, in `strcpy02` the student's `strlen` solution is called as a helper function for `strcpy`.

function	YES	NO	MAYBE	time
sum	9/9	0/0	6/6	2.1/ 2.1
fib	4/10	6/1	3/2	7.6/ 5.6
sumfrom	3/3	1/0	2/3	1.8/ 2.1
strlen	1/2	0/0	5/4	4.1/ 4.0
strcpy	3/5	0/0	3/1	21.8/ 17.1
arrsum	1/1	0/0	0/0	3.9/ 3.9
fact	1/1	0/0	0/0	2.2/ 2.2
literature	4/5	3/2	18/18	4.0/ 3.9
safety	3/3	2/2	7/7	22.3/ 22.3
total	29/39	12/5	44/41	

Legend: YES indicates that a proof was found, NO a disproof (so a conclusion \perp); MAYBE denotes that Ctrl found no proof or disproof, took more than 60 seconds, or failed to prove termination of the LCTRS. The time column lists the average runtime on YES and NO results.

Fig. 1. Results of Ctrl in the initial test (before the slash), and with obvious mistakes fixed (after the slash).

We found five classes of recurring failures. First, cases where the function was wrong, but Ctrl could not answer NO as it could not prove confluence. This accounts for six MAYBEs in the initial test and two in the second, and could be considered an incorrect implementation. Second (six failures in either table) is the termination requirement: we need termination independent from the starting symbol, which is often not satisfied or cannot be proved by our admittedly limited termination module.

The remaining groups of failures each demonstrate a weakness of our method. The third failure occurs when generalization drops a relation between two variables; e.g., when x and y are both initialized to 0 and then increased by 1 in every loop iteration (with loops corresponding to tail-recursive functions); after generalizing, the information that they are equal is lost. Typically, this manifests as an EXPANSION where the non-diverging case can easily be removed before generalization, but afterwards gives an equation that can be disproved. This suggests a natural direction for improvement.

The fourth group are those benchmarks where our primary generalization technique (§ 5.1) does not apply because there are no variables to generalize. This happens when both sides have non-tail-recursive functions or loops counting down rather than up. Recursion-abstraction (§ 5.2) lets us solve several benchmarks, but further lemma generation will be needed for the majority. Nonetheless, this generalization technique does allow us to handle Ex. 4.30, which can be challenging for existing approaches.

The final group concerns nested loops. Ctrl's strategy fails because the counters for the inner and outer loop are generalized at the same time. However, inductive proofs with Ctrl's interactive mode show that such benchmarks *can* be handled by our method. Thus, in future work a more sophisticated generalization strategy would be desirable.

Demonstrative examples of these last three issues are given in Appendix D. A full evaluation page, including exact problem statements, is given at:

<http://cl-informatik.uibk.ac.at/software/ctrl/tocl/>

7. RELATED WORK

The related work can be split into two categories. First, the literature on rewriting induction; and second, the work on program verification and equivalence analysis.

7.1. Rewriting Induction

Our inductive theorem proving method builds on a long literature about rewriting induction (see e.g., [Bouhoula 1997; Falke and Kapur 2012; Reddy 1990; Sakata et al. 2009]). Its core method extends existing techniques to the LCTRS formalism introduced in [Kop and Nishida 2013], thus generalizing the possibilities of earlier work.

The most relevant related works are [Falke and Kapur 2012; Sakata et al. 2009], defining rewriting induction for different styles of constrained rewriting. Both use only

integer functions and predicates; it is not clear how to generalize these approaches to more advanced theories. The more general setting of LCTRSs enables rewriting induction also for systems with, e.g., arrays, bitvectors, or real numbers. Moreover, not restricting the predicates in Σ_{theory} enables (a limited form of) quantifiers in constraints.

These advantages are enabled by subtle changes to the inference rules, in particular SIMPLIFICATION and EXPANSION. Our changes let us modify constraints of an equation and handle *irregular* rules with fresh variables in the constraint. This additionally enables EXPANSION steps to create such (otherwise infeasible) rules. The method requires a very different implementation from previous definitions: we need separate strategies to simplify constraints (e.g., deriving quantified statements), and, for the desired generality, must rely primarily on external solvers to manipulate constraints.

Moreover, we have introduced a completely new generalization technique, as a powerful tool for analyzing loops in particular. Nakabayashi et al. [2010] use a similar idea (abstracting the initialization values), but the execution is very different: for an equation $s \approx t [\varphi]$, first $s \approx t$ is adapted via templates obtained from the rules, then φ is generalized via a set of relations between positions tracked by the proof process. In our method, the constraint carries all the information. We succeed on all examples in [Nakabayashi et al. 2010], and on some where their method fails (cf. Appendix C; e.g., for non-negative n , a for-loop summing up from 1 to n is compared to $n \cdot (n+1) / 2$).

For *unconstrained* systems, the literature contains several generalization methods, e.g., [Kapur and Sakhanenko 2003; Kapur and Subramaniam 1996; Urso and Kounalis 2004]. Mostly, our method in § 5.1 is very different from these approaches. Most similar, perhaps, is [Kapur and Sakhanenko 2003], which also proposes a method to generalize initial values. As observed by Nakabayashi et al. [2010], this method is not sufficient for even our simplest benchmarks `sum` and `fact`, as the argument for the loop variable cannot be generalized; in contrast, our method has no problem with such variables. As discussed in § 5.2, the recursion-abstraction technique presented there essentially lifts a technique from explicit induction [Aubin 1979] to constrained rewriting induction.

As far as we are aware, there is no other work for lemma generation of rewrite systems (or functional programs) obtained from procedural programs.

Like Giesl et al. [2007], we verify procedural programs via a transformation to a functional program, followed by an invocation of an inductive theorem prover. In an unconstrained setting, they propose an equivalence-preserving program transformation to a non-tail-recursive program to eliminate accumulator arguments. A combination of their approach with ours could be beneficial e.g. for programs with nested loops.

7.2. Automatic Program Verification and Equivalence Proving

Our goal is to (automatically) verify correctness properties of procedural programs. Fully automated verifiers for properties like (memory) safety and termination are regularly assessed at the *Competition on Software Verification* [SV-COMP]. However, a comparison with these tools does not seem useful. While we can, to some extent, tackle (memory) safety and termination, our main topic is *equivalence*, which is not studied in SV-COMP. Technically, equivalence problems can be formulated as safety problems (by self-composition [Barthe et al. 2011]: call both programs on equal inputs and assert that their results are also equal). However, none of the tools in the “recursive” category of SV-COMP 2015 could prove equivalence for our simplest (integer) example `sum`.

Apart from constrained rewriting, another intermediate representation for verification of imperative programs is based on (constrained) logic programs or, closely related, Horn clauses [Albert et al. 2007; Gupta et al. 2011]. It should be possible to express our contributions also in this framework, provided that constructor terms are supported.

For the setting of Ex. 1.1, automated grading, Vujosevic-Janjic et al. [2013] apply verification techniques like bounded model checking. While this enables significant

improvements over classic testing, there is still a non-zero risk of missing bugs due to under-approximation. Thus, it could be beneficial to add our approach to the portfolio.

For program equivalence, we discuss (fully) automated techniques for proving *partial equivalence* and its special case *total equivalence*. Two programs P_1 and P_2 are *partially equivalent* if for the same inputs, terminating executions of P_1 and P_2 return the same value. They are *totally equivalent* if they moreover both terminate on all inputs (see [Godlin and Strichman 2008] for a more extensive discussion).

This paper addresses total equivalence: we require termination to analyze partial equivalence. We allow *constrained* equivalence queries so that only certain inputs are considered. This includes properties that cannot be checked programmatically, like the *size* of an array in a C program. As mentioned in § 6.1, for non-confluent programs P_1 and P_2 , we analyze if running P_1 on the input *can* lead to the same result as P_2 .

Godlin and Strichman [2008] propose a Hoare-style proof rule for partial equivalence of recursive programs (among other properties). To analyze two recursive functions f_1 and f_2 , these symbols are first replaced in recursive calls in their bodies by the same *uninterpreted function symbol* f . Under this premise, it is then proved (e.g., by a bounded model checker) that the bodies of f_1 and f_2 also have equivalent results. In this sense, Godlin and Strichman [2008] also use inductive reasoning. However, our approach proves equivalence of Ex. 4.30 with different recursion base cases, whereas their proof rule is not applicable. Moreover, the use of uninterpreted function symbols requires that the programs must be deterministic, in contrast to our approach.

Lopes and Monteiro [2016] prove partial equivalence for programs on integers and undefined function symbols (which may arise also as abstractions of deterministic complex functions). They combine self-composition [Barthe et al. 2011], a safety-preserving transformation of undefined functions to polynomials (yielding a program on integers only), recurrence solving for loops, and a standard software model checker. However, their approach does not support *mutable* arrays, whose content can be changed during the program's execution (as in Ex. 4.28 for `strcpy`), in contrast to our method.

Verdoolaege et al. [2012] use widening to prove program equivalence. For validation of compiler optimizations [Necula 2000], they consider programs with (linear-)affine arithmetic and arrays. A restriction of their approach is that it does not exploit the *semantics* of arithmetic operations beyond associativity and commutativity.

Recently, *regression verification* has become an active topic of research in program equivalence proving [Godlin and Strichman 2013; Lahiri et al. 2012; Felsing et al. 2014]. As in regression testing, two programs are compared that are syntactically *almost* the same, e.g., different revisions of the same code base with a refactored function. Regression verification then analyzes if the two programs are semantically equivalent.

Godlin and Strichman [2013] improve modularity over [Godlin and Strichman 2008] by decomposing the proof obligations into smaller units via the call graph of the program. Hawblitzel et al. [2013] propose *mutual summaries*, relating the postconditions of two program functions. This generalizes uninterpreted functions as summaries and allows analysis of non-deterministic programs. A challenge is to find such mutual summaries automatically. Felsing et al. [2014] address this problem via Horn constraint solving to find *coupling predicates* over linear arithmetic between program points. It would be interesting to adapt their approach for lemma generation. They also analyze *total* equivalence: a separate termination proof is required. The web interface of their tool `llrêve` currently fails on the same example as [Nakabayashi et al. 2010] (cf. § 7.1). They mention an extension to arrays and heap data structures as future work.

8. DIRECTIONS FOR FUTURE WORK

This paper is by no means intended as the end station for inductive theorem proving on LCTRSs, but rather as the beginning. The generalization methods we supply are

powerful together, but they do not suffice for more complicated systems or equations. A mere two methods cannot bypass the need to search for loop invariants altogether.

A natural extension would thus be both to adapt existing lemma generation techniques to the constrained setting and to adapt techniques for finding loop invariants towards the setting of rewriting induction, e.g., to suggest suitable lemmas. It might also be worthwhile to directly look at the constraints and develop advanced methods for constraint modification, which could be followed by a generalization step. Moreover, our generalization technique from § 5.1 could be improved to generalize not only initializations with constants, but also initializations with other values, e.g., copies of function parameters. This is motivated by loops that count down instead of up. Additionally, inspired by [Lopes and Monteiro 2016], one might consider LCTRSs with uninterpreted functions to model functions with unknown implementations.

For a different direction, we may extend the translation from § 3, e.g., by translating structs to term data structures (cf. [Otto et al. 2010]). The ideas from § 3 can also be applied for languages such as Python or Java, enabling equivalence proofs between functions in different languages. This could be particularly interesting for a reference implementation in an inherently memory-safe language like F# or Java, and an efficient implementation in a language like C that has no such memory safety guarantees.

Finally, we hope to extend the implementation in the future, both to increase the strength of the inductive theorem proving—adding new theory and testing for more sophisticated heuristics—and to add more features to the translation from C code.

9. CONCLUSIONS

In this paper, we have done two things. First, we have discussed a transformation from procedural programs to constrained term rewriting. By abstracting from the memory model underlying a particular programming language and instead encoding concepts like integers and arrays in an intuitive way, this transformation can be applied to various different (imperative) programming languages. The resulting LCTRS is close to the original program and has built-in error checking for all mistakes of interest.

Second, we have extended rewriting induction to the setting of LCTRSs. We have shown how this method can be used to prove correctness of procedural programs. The LCTRS formalism is a good analysis backend for this, since the techniques from standard rewriting can typically be extended to it, and native support for logical constraints and data types like integers and arrays is present.

We have also introduced two new techniques to generalize equations. The idea of the core method is to identify constants used as *variable initializations*, keep track of them during the proof process, and abstract from these constants when a proof attempt diverges. The LCTRS setting is instrumental in the simplicity of this method, as it boils down to dropping a (cleverly chosen) part of a constraint. The second method recognizes—and abstracts—recursive calls on semantically equivalent arguments.

In addition to the theory, we provide an implementation of these techniques. Initial results on a small database of programs from students and the literature are very promising. In future work, we aim to increase the strength of our implementations.

ELECTRONIC APPENDIX

The electronic appendix for this article can be accessed in the ACM Digital Library.

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A. TRANSLATING C PROGRAMS TO LCTRSS

This appendix provides further details on the translation from C programs to LCTRSSs.

A.1. Optimizing LCTRSSs

After generating the LCTRS, we simplify the (left-linear) result by the following steps:

- (1) *Combining unconstrained rules.* Like [Falke et al. 2011], we repeat the following:
 - select any unconstrained rule ρ of the form $u(x_1, \dots, x_n) \rightarrow r$ where u is not the initial symbol of a C function (like `fact` in Ex. 3.1), and u neither occurs in r nor in the left-hand side of any other rule; the repetition stops if no such ρ exists;
 - rewrite all right-hand sides with ρ ;
 - remove both the rule ρ and the symbol u .

This process does not substantially alter the multi-step reduction relation \rightarrow^* as the only symbols removed are those which we think of as “intermediate” symbols.
- (2) *Combining constrained rules.* If there are distinct rules $\ell \rightarrow r [\varphi]$ and $\ell \rightarrow r [\psi]$ (modulo renaming), these are combined into $\ell \rightarrow r [\varphi \vee \psi]$. Given rules $\ell \rightarrow u(s_1, \dots, s_m) [\varphi]$ and $u(x_1, \dots, x_m) \rightarrow r_i [\psi_i]$ for $1 \leq i \leq n$ with all x_j variables, we may replace them by $\ell \rightarrow r_i[x_1 := s_1, \dots, x_m := s_m] [\varphi \wedge \psi_i[x_1 := s_1, \dots, x_m := s_m]]$ for $1 \leq i \leq n$, if:
 - u is not the initial symbol of a function and does not occur in any other rule, or ℓ ;
 - the terms s_j do not contain defined symbols (as then we might remove a non-terminating subterm, which *would* impact the multi-step reduction relation).
- (3) *Removing unused arguments.* For all function symbols and all their argument positions, we mark whether the position is “used”:
 - all argument(s) of every return_f and initial symbols (e.g. `fact`) are used;
 - for other symbols u_i of arity n and every $1 \leq j \leq n$: if there is a rule $u_i(\ell_1, \dots, \ell_n) \rightarrow r [\varphi]$ where ℓ_j is not a variable (which can arise for instance with the transformation in § 3.5) or occurs in φ , then argument j is used in u_i ;
 - for all rules $u_i(\ell_1, \dots, \ell_n) \rightarrow r [\varphi]$ and $1 \leq j \leq n$: argument j is used in u_i if ℓ_j is a variable occurring at a *used position* in r ; here, a position p is used in s if either $p = \epsilon$ or $p = i \cdot p'$, $s = f(\vec{t})$, argument i is used in f and position p' is used in t_i .

The last, recursive, step essentially calculates a fixpoint; in summary, an argument position is *used* if it is possible to reduce to a term where we actually need the subterm at that position as part of a constraint or the function’s return value. When a variable is not used in any later statement, we will avoid carrying it along.
- (4) *Simplifying constraints.* Constraints may be brought into an equivalent form, e.g., by removing duplicate clauses or replacing, e.g., $\neg(x > y)$ by $x \leq y$. Here, φ is “equivalent” to ψ in a rule $\ell \rightarrow r [\varphi]$ if $\forall \vec{x}(\exists \vec{y}(\varphi) \leftrightarrow \exists \vec{z}(\psi))$ holds, where $\text{Var}(\ell) \cup \text{Var}(r) = \{\vec{x}\}$, $\text{Var}(\varphi) \setminus \{\vec{x}\} = \{\vec{y}\}$, and $\text{Var}(\psi) \setminus \{\vec{x}\} = \{\vec{z}\}$ (much like the observation on \sim below Def. 2.14). We typically only remove negations and unused variables.

Example A.1. As an example, let us consider the simplification of a toy function.

<pre>int f(int x) { int y,z; if (x < 0) return 0; z = 0; while (x > 0) { x--; z += x; } y = z + x; return y; }</pre>	$f(x) \rightarrow u_1(x, y, z)$ $u_1(x, y, z) \rightarrow u_2(x, y, z) \quad [x < 0]$ $u_1(x, y, z) \rightarrow u_3(x, y, z) \quad [\neg(x < 0)]$ $u_2(x, y, z) \rightarrow \text{return}_f(0)$ $u_3(x, y, z) \rightarrow u_4(x, y, 0)$ $u_4(x, y, z) \rightarrow u_5(x, y, z) \quad [x > 0]$ $u_4(x, y, z) \rightarrow u_7(x, y, z) \quad [\neg(x > 0)]$ $u_5(x, y, z) \rightarrow u_6(x - 1, y, z)$ $u_6(x, y, z) \rightarrow u_4(x, y, z + x)$ $u_7(x, y, z) \rightarrow u_8(x, z + x, z)$ $u_8(x, y, z) \rightarrow \text{return}_f(y)$
--	---

The rule $f(x) \rightarrow u_1(x, y, z)$ has unconstrained variables y and z in the right-hand side which do not occur on the left. A step with this rule instantiates y and z by arbitrary type-correct values. This reflects that in the C program the variables y and z are at first not initialized and may contain an arbitrary value (depending on the compiler). In the simplified version, this does not occur; consider the remainder obtained from combining rules:

$$\begin{array}{ll} f(x) \rightarrow \text{return}_f(0) & [x < 0] \\ f(x) \rightarrow u_4(x, y, 0) & [\neg(x < 0)] \\ u_4(x, y, z) \rightarrow u_4(x - 1, y, z + x - 1) & [x > 0] \\ u_4(x, y, z) \rightarrow \text{return}_f(z + x) & [\neg(x > 0)] \end{array}$$

Now, the first and third arguments of u_4 are used (in the constraint and return value), but the second is not: it is merely passed along in the

recursive call. Removing this variable and simplifying the constraints, we obtain:

$$\begin{array}{ll} f(x) \rightarrow \text{return}_f(0) & [x < 0] \\ f(x) \rightarrow u_4(x, 0) & [x \geq 0] \\ u_4(x, z) \rightarrow u_4(x - 1, z + x - 1) & [x > 0] \\ u_4(x, z) \rightarrow \text{return}_f(z + x) & [x \leq 0] \end{array}$$

This system is *orthogonal* in the sense of [Kop and Nishida 2013] and thus *confluent*, which is beneficial for analysis. The original LCTRS was also confluent, but this was harder to see.

Correctness relies on the fact that the LCTRSs created using the transformation described in § 3 are “well behaved”; most importantly, all rules are left-linear.

A.2. Translating C Programs with Explicit Pointers

As observed at the end of § 3.6, the simple translation explored there has both up- and downsides. On the one hand, by abstracting from the memory model, we can simplify analysis. On the other hand, there are certain programs we cannot handle.

For C programs with dynamically allocated arrays and/or explicit pointer use, we consider the memory model from the C standard. Declaring or allocating an array selects an amount of currently unused space in memory and designates it for use by the given array. The allocated space is *not* guaranteed to be at a given position in memory relative to existing declarations; when an array is indexed out of its declared bounds, the resulting behavior is undefined—so this can safely be considered an error (see paragraph 6.5.6:9 in: <http://www.open-std.org/jtc1/sc22/wg14/www/docs/n1570.pdf>).

We will think of a program’s memory as a set of *blocks*, each block corresponding to a sequence of values. A pointer then becomes a location in such a block. In an LCTRS we will model this using a “global memory” variable, which lists the blocks as a sequence of arrays; a pointer is a pair of integers, selecting a memory block and its offset.

Limiting interest to programs on (dynamically allocated) integer or char arrays, we will use a memory variable of sort $\text{array}(\text{array}(\text{int}))$, which represents a sequence of integer arrays (i.e. $(\mathbb{Z}^*)^*$); the default value $0_{\text{array}(\text{int})}$ is the empty sequence $\langle \rangle \in \mathbb{Z}^*$. We use a theory signature with the array symbols introduced in § 3.6, along with:

- $\text{allocate} : [\text{array}(\text{array}(\text{int})) \times \text{array}(\text{int})] \Rightarrow \text{array}(\text{array}(\text{int}))$, where $\mathcal{I}_{\text{allocate}}(\langle a_0, \dots, a_k \rangle, b) = \langle a_0, \dots, a_k, b \rangle$; that is, $\text{allocate}(\text{mem}, \text{arr})$ adds the new sequence arr to the memory;
- $\text{free} : [\text{array}(\text{array}(\text{int})) \times \text{int}] \Rightarrow \text{array}(\text{array}(\text{int}))$, where $\mathcal{I}_{\text{free}}(\langle a_0, \dots, a_k \rangle, n) = \langle a_0, \dots, a_{n-1}, \langle \rangle, a_{n+1}, \dots, a_k \rangle$ if $0 \leq n \leq k$ and $\langle a_0, \dots, a_k \rangle$ otherwise; that is, the memory block indexed by n is considered empty, and any further attempt to address a location in that memory block should be considered an error.

A pointer is represented by a pair (b, o) of a block index and an offset within that block. The NULL-pointer is represented by $(-1, 0)$.

Example A.2. Consider the following example C++ function:

```
int *create(int k) {
  int *a = new int[k];
  int *b = a + 1;
  for (int i = 0; i < k; i += 2) b[i] = 42;
  return a;
}
```

Now, a and b share memory, and new memory is allocated. We might encode this as:

$$\begin{aligned}
\text{create}(mem, k) &\rightarrow u(\text{allocate}(mem, x), k, \text{size}(mem), 0) && [\text{size}(x) = k] \\
u(mem, k, ai, ao) &\rightarrow v(mem, k, ai, ao, ai, ao + 1, 0) \\
v(mem, k, ai, ao, bi, bo, i) &\rightarrow w(mem, k, ai, ao, bi, bo, i) && [i < k] \\
v(mem, k, ai, ao, bi, bo, i) &\rightarrow \text{return}(mem, ai, ao) && [i \geq k] \\
w(mem, k, ai, ao, bi, bo, i) &\rightarrow \text{error} && [bo + i < 0 \vee bo + i \geq \text{size}(\text{select}(mem, bi))] \\
w(mem, k, ai, ao, bi, bo, i) &\rightarrow v(\text{store}(mem, bi, \text{store}(\text{select}(mem, bi), bo + i, 42)), k, \\
&\quad ai, ao, bi, bo, i + 2) && [0 \leq bo + i < \text{size}(\text{select}(mem, bi))]
\end{aligned}$$

(For clarity, we omit the optimization step that combines the first two rules, and the one that combines the third with the last two.)

Consider how this example is executed, starting from empty memory. We will use $\langle \cdot \rangle$ to refer to specific arrays of type $\text{array}(\text{array}(\text{int}))$ and $[\cdot]$ for arrays of type $\text{array}(\text{int})$.

- (1) We call $\text{create}(\langle \rangle, 2)$, representing a function call when no arrays have been allocated.
- (2) By the first rule, we get $u(\text{allocate}(\langle \rangle, x), 2, \text{size}(\langle \rangle), 0)$, where x is a *random* array. All we know is that it has size 2—this rule uses irregularity to represent the randomness involved in an allocation. Thus, assume the sequence $[-4, 9]$ is chosen. Using calculation steps to evaluate allocate and size , we get $u(\langle [-4, 9] \rangle, 2, 0, 0)$. Here, the pair $(0, 0)$ represents the array a : the first block in memory, read from the start (offset 0).
- (3) Then by the second rule (and a calculation), we reduce to $v(\langle [-4, 9] \rangle, 2, 0, 0, 0, 1, 0)$. The new pair $(0, 1)$ represents b : the same memory block as a , but with offset 1. This location points to the sequence $[9]$. The final 0 is the index for the loop counter i .
- (4) Entering the loop (as indeed $0 < 2$), we reduce to $w(\langle [-4, 9] \rangle, 2, 0, 0, 0, 1, 0)$.
- (5) Here, we do an array store: $b[i] = 42$; . The LCTRS first tests whether $b[i]$ corresponds to a position in allocated memory and reduces to an error state if not. This is done by selecting the corresponding block from mem , then testing whether the offset for b and i together exceed the block's bounds. We succeed, as $0 \leq 0 + 1 < \text{size}(\text{select}(\langle [-4, 9] \rangle, 0)) \Leftrightarrow 0 \leq 1 < \text{size}([-4, 9]) \Leftrightarrow 0 \leq 1 < 2$.
- (6) Thus, the update is done: we reduce to:

$$\begin{aligned}
&v(\text{store}(\langle [-4, 9] \rangle, 0, \text{store}(\text{select}(\langle [-4, 9] \rangle, 0), 1 + 0, 42)), 2, 0, 0, 0, 1, 0 + 2) \\
&\rightarrow_{\text{calc}}^* v(\text{store}(\langle [-4, 9] \rangle, 0, \text{store}([-4, 9], 1, 42)), 2, 0, 0, 0, 1, 2) \\
&\rightarrow_{\text{calc}}^* v(\text{store}(\langle [-4, 9] \rangle, 0, [-4, 42]), 2, 0, 0, 0, 1, 2) \\
&\rightarrow_{\text{calc}} v(\langle [-4, 42] \rangle, 2, 0, 0, 0, 1, 2)
\end{aligned}$$
 So we retrieve the space for b from memory (getting the full block $[-4, 9]$), update the position corresponding to $b[0]$ (which is the same as $a[1]$), get $[-4, 42]$, and store the result into the corresponding position in memory. Then we carry on with $i + 2$.
- (7) Since $2 \geq 2$, we reduce to $\text{return}(\langle [-4, 42] \rangle, 0, 0)$, returning the dynamic array $[-4, 42]$.

Note that in step 5, we do not test whether b corresponds to currently allocated memory. This is safe because, if b is the NULL-pointer or corresponds to previously freed memory, then $\text{select}(mem, bi)$ is $\langle \rangle$, and any indexing in this array will cause an error regardless. Note also that this function gives a non-error result only for *even* k .

While Ex. A.2 considers only integer arrays, we could also handle programs with dynamically allocated arrays of varying types. In this case, we would simply use *multiple* memory variables with different type declarations.

B. CORRECTNESS PROOF

In this appendix, we give the full correctness proof, which was only sketched in § 4.4.

First, we prove Lemma 4.31, reformulated as follows:

LEMMA 4.31. *The following statements are equivalent:*

- all equations in \mathcal{E} are inductive theorems;
- $\leftrightarrow_{\mathcal{E}} \subseteq \leftrightarrow_{\mathcal{R}}^*$ on ground terms (so if s, t are ground and $s \leftrightarrow_{\mathcal{E}} t$, then also $s \leftrightarrow_{\mathcal{R}}^* t$).

PROOF. Suppose $\leftrightarrow_{\mathcal{E}} \subseteq \leftrightarrow_{\mathcal{R}}^*$ on ground terms. If $s \approx t [\varphi] \in \mathcal{E}$ and the ground constructor substitution γ respects this equation, then $s\gamma$ and $t\gamma$ are ground (since, by definition of “respects” (Def. 4.3), $\text{Var}(s) \cup \text{Var}(t) \subseteq \text{Dom}(\gamma)$). Since obviously $s\gamma \leftrightarrow_{\mathcal{E}} t\gamma$ (with empty C), by assumption $s\gamma \leftrightarrow_{\mathcal{R}}^* t\gamma$. Thus, $s \approx t [\varphi]$ is an inductive theorem.

Suppose that all equations in \mathcal{E} are inductive theorems, and $u \leftrightarrow_{\mathcal{E}} w$ for ground u, w ; we must see that $u \leftrightarrow_{\mathcal{R}}^* w$. We have $u = C[s\gamma]$ and $w = C[t\gamma]$ for some $s \approx t [\varphi] \in \mathcal{E}$ and substitution γ that respects φ and maps all variables in s, t to ground terms. Let δ be a substitution such that each $\delta(x)$ is a normal form of $\gamma(x)$; by termination of \mathcal{R} , such a δ exists, and by quasi-reductivity, it is a ground constructor substitution. As values cannot be reduced, also δ respects φ . Therefore $s\delta \leftrightarrow_{\mathcal{E}} t\delta$, which implies $s\delta \leftrightarrow_{\mathcal{R}}^* t\delta$. We conclude: $C[s\gamma] \leftrightarrow_{\mathcal{R}}^* C[s\delta] \leftrightarrow_{\mathcal{R}}^* C[t\delta] \leftrightarrow_{\mathcal{R}}^* C[t\gamma]$, giving the desired result. \square

Recall also the following key lemma (whose proof has been given in the main text):

LEMMA 4.32 ([SAKATA ET AL. 2009]). *Let \rightarrow_1 and \rightarrow_2 be binary relations over some set A . Then, $\leftrightarrow_1^* = \leftrightarrow_2^*$ if all of the following hold:*

- $\rightarrow_1 \subseteq \rightarrow_2$,
- \rightarrow_2 is well founded, and
- $\rightarrow_2 \subseteq (\rightarrow_1 \cdot \rightarrow_2^* \cdot \leftrightarrow_1^* \cdot \leftarrow_2^*)$.

Lemma 4.33 in the main text is the combination of the following Lemmas B.1–B.4.

LEMMA B.1. *Let s, t be terms, φ a constraint and p a position of s such that $s|_p$ has the form $f(s_1, \dots, s_n)$ with f a defined symbol and all s_i constructor terms. Suppose that the variables in s, t, φ are distinct from those in \mathcal{R} . Then:*

- (1) *For any ground constructor substitution γ which respects $s \approx t [\varphi]$ and any choice of $\text{Expd}(s \approx t [\varphi], p)$, we have:*

$$s\gamma \left(\rightarrow_{\mathcal{R}, p} \cdot \leftrightarrow_{\text{Expd}(s \approx t [\varphi], p)} \right) t\gamma$$

Here, $\rightarrow_{\mathcal{R}, p}$ indicates a reduction at position p with a rule in $\mathcal{R} \cup \mathcal{R}_{\text{calc}}$.

- (2) *For any $s' \rightarrow t' [\varphi']$ in any choice of $\text{Expd}(s \approx t [\varphi], p)$ and any ground constructor substitution δ which respects $s' \approx t' [\varphi']$, we have:*

$$s'\delta \left(\leftarrow_{\mathcal{R}, p} \cdot \leftrightarrow_{\{s \approx t [\varphi]\}} \right) t'\delta$$

PROOF. $s\gamma|_p = s|_p\gamma = f(s_1\gamma, \dots, s_n\gamma)$, where all $s_i\gamma$ are ground constructor terms. Since f is defined, $f(\overrightarrow{s}\gamma)$ reduces by quasi-reductivity, which can only be a root reduction. Thus, $s\gamma = (s\gamma)[\ell\delta]_p$ for some rule $\ell \rightarrow r [\psi]$ and substitution δ which respects ψ . Since the rule variables are distinct from the ones in the equation, we can assume that δ is an extension of γ , so $s\gamma = s[\ell]_p\delta$. Clearly, both $\varphi\delta$ and $\psi\delta$ evaluate to \top .

As δ unifies $s|_p$ and ℓ , there is a most general unifier η , so $s|_p\eta = \ell\eta$ and we can write $\delta = \delta' \circ \eta$ for some δ' . Now, by definition of constrained term reduction, any choice of

$Expd(s \approx t [\varphi], p)$ has an element $s' \approx t' [\varphi']$ where we can write (for suitable u, η' etc.):

$$\begin{array}{l} s\eta[\ell\eta]_p \approx t\eta [\varphi\eta \wedge \psi\eta] \\ \sim u[\ell\eta']_p \approx t'' [\varphi''] \\ \xrightarrow{\ell \rightarrow r[\psi], 1 \cdot p} u[r\eta']_p \approx t'' [\varphi''] \\ \sim s' \approx t' [\varphi'] \end{array}$$

Consider the “term” $s\delta \approx t\delta$. This is an instance of the first constrained term in this reduction, so by Thm. 2.19, this “term” reduces at position $1 \cdot p$ to $s'\delta'' \approx t'\delta''$ for some substitution δ'' which respects φ' . As the reduction happens inside $s\delta$, we see that $t\delta = t'\delta''$. Thus, $s\gamma = s\delta \rightarrow_{\mathcal{R}} s'\delta'' \leftrightarrow_{Expd(s \approx t [\varphi], p)} t'\delta'' = t\delta = t\gamma$.

As for the second part, note that by definition of $Expd$ there are a substitution γ and constraint ψ such that the constrained term $s\gamma \approx t\gamma [\varphi\gamma \wedge \psi\gamma]$ reduces to $s' \approx t' [\varphi']$ at position $1 \cdot p$. By Thm. 2.20, we find a substitution η which respects $\varphi\gamma \wedge \psi\gamma$, such that $s\gamma\eta \approx t\gamma\eta \rightarrow_{\mathcal{R}} s'\delta \approx t'\delta$ at position $1 \cdot p$. Since the reduction takes place in the left part of \approx , we have $t\gamma\eta = t'\delta$ and $s\gamma\eta \rightarrow_{\mathcal{R}} s'\delta$. We are done if also $s\gamma\eta \leftrightarrow_{s \approx t [\varphi]} t\gamma\eta$, which indeed holds because $\eta \circ \gamma$ respects φ (as $(\varphi\gamma \wedge \psi\gamma)\eta$ implies $\varphi\gamma\eta$). \square

LEMMA B.2. *Suppose that $(\mathcal{E}, \mathcal{H}, \text{flag}) \vdash_{\text{ri}} (\mathcal{E}', \mathcal{H}', \text{flag}')$ by any inference rule other than COMPLETENESS. Then,*

$$\leftarrow \parallel_{\mathcal{E}} \subseteq (\rightarrow_{\mathcal{R} \cup \mathcal{H}}^* \cdot \leftarrow \parallel_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}'}^*)$$

on ground terms.

Here, $\leftarrow \parallel_{\mathcal{E}'}$ denotes a *parallel* application of zero or more $\leftrightarrow_{\mathcal{E}'}$ steps.

PROOF. It suffices to show that $\leftrightarrow_{\mathcal{E} \setminus \mathcal{E}'} \subseteq (\rightarrow_{\mathcal{R} \cup \mathcal{H}}^* \cdot \leftarrow \parallel_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}'}^*)$ on ground terms: if $C[u_1, \dots, u_n] \leftarrow \parallel_{\mathcal{E}} C[v_1, \dots, v_n]$ because each $u_i \leftrightarrow_{\rho_i} v_i$ for some $\rho_i \in \mathcal{E}$, then this gives $u_i \rightarrow_{\mathcal{R} \cup \mathcal{H}'}^* u'_i \leftarrow \parallel_{\mathcal{E}'} v'_i \leftarrow_{\mathcal{R} \cup \mathcal{H}'}^* v_i$ if $\rho_i \notin \mathcal{E}'$ and $u_i = u'_i \leftarrow \parallel_{\mathcal{E}'} v'_i = v_i$ if $\rho_i \in \mathcal{E}'$, so (sequentializing parallel steps) $C[u_1, \dots, u_n] \rightarrow_{\mathcal{R} \cup \mathcal{H}'}^* C[u'_1, \dots, u'_n] \leftarrow \parallel_{\mathcal{E}'} C[v'_1, \dots, v'_n] \leftarrow_{\mathcal{R} \cup \mathcal{H}'}^* C[v_1, \dots, v_n]$ as desired. For all inference rules (except COMPLETENESS) either $\mathcal{E} \setminus \mathcal{E}' = \emptyset$ or we can write $\mathcal{E} \setminus \mathcal{E}' = \{s \simeq t [\varphi]\}$. Consider which inference rule is applied for \vdash_{ri} .

- (SIMPLIFICATION). Suppose that $s \simeq t [\varphi]$ is replaced by $u \approx t [\psi]$ where $s \approx t [\varphi] \rightarrow_{\mathcal{R} \cup \mathcal{H}} u \approx t [\psi]$. Let $C[s\gamma] \leftrightarrow_{\{s \simeq t [\varphi]\}} C[t\gamma]$, where γ is a substitution which respects φ . It follows from Thm. 2.19 that $s\gamma \approx t\gamma \rightarrow_{\mathcal{R} \cup \mathcal{H}} u\delta \approx t\delta$ where δ is a substitution which respects ψ , and thus, as \approx is a constructor, $C[s\gamma] \rightarrow_{\mathcal{R} \cup \mathcal{H}} C[u\delta]$ and $t\gamma = t\delta$. Then, $C[u\delta] \leftrightarrow_{\{u \approx t [\psi]\}} C[t\delta] = C[t\gamma]$, and we have $C[s\gamma] \rightarrow_{\mathcal{R} \cup \mathcal{H}} \cdot \leftrightarrow_{\mathcal{E}'} C[t\gamma]$. Symmetrically, if $C[t\gamma] \leftrightarrow_{\{s \simeq t [\varphi]\}} C[s\gamma]$, then $C[t\gamma] \leftrightarrow_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}} C[s\gamma]$. Thus, $\leftrightarrow_{s \simeq t [\varphi]} \subseteq (\rightarrow_{\mathcal{R} \cup \mathcal{H}}^* \cdot \leftrightarrow_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}}^*)$. This suffices because in this case $\mathcal{H} = \mathcal{H}'$.
- (DELETION). In the case that $s = t$, the relation $\leftrightarrow_{\mathcal{E} \setminus \mathcal{E}'}$ is the identity. Otherwise, φ is unsatisfiable, so $s \simeq t [\varphi]$ is never used, i.e., $\leftrightarrow_{\mathcal{E} \setminus \mathcal{E}'} = \emptyset$.
- (EXPANSION). Suppose $C[s\gamma] \leftrightarrow_{s \simeq t [\varphi]} C[t\gamma]$, where γ respects $s \simeq t [\varphi]$; as we only consider ground terms, $\gamma(x)$ is ground for all variables in its domain. Noting that by quasi-reductivity and termination every ground term reduces to a ground constructor term, let δ be a substitution where for each $x \in \text{Dom}(\gamma)$, $\delta(x)$ is a constructor term such that $\gamma(x) \rightarrow_{\mathcal{R}}^* \delta(x)$. Then it follows from Lemma B.1 that $C[s\gamma] \rightarrow_{\mathcal{R}}^* C[s\delta] (\rightarrow_{\mathcal{R}} \cdot \leftrightarrow_{\mathcal{E}'} C[t\delta] \leftarrow_{\mathcal{R}}^* C[t\gamma])$. The situation where $C[t\gamma] \leftrightarrow_{s \simeq t [\varphi]} C[s\gamma]$ is symmetric.
- (EQ-DELETION). Let $s = C[s_1, \dots, s_n]$ and $t = C[t_1, \dots, t_n]$ where $s_1, t_1, \dots, s_n, t_n \in \text{Terms}(\Sigma_{\text{theory}}, \text{Var}(\varphi))$. Any ground substitution γ which respects φ , and whose domain contains all variables in the terms s_i and t_i , must map these variables to values. Therefore, $s_i\gamma \rightarrow_{\text{calc}}^* v_i$ and $t_i\gamma \rightarrow_{\text{calc}}^* w_i$, where v_i is the value of $s_i\gamma$ and w_i is the value of $t_i\gamma$. Now, suppose $q \leftrightarrow_{s \simeq t [\varphi]} u$ for ground q, u . Then (a) $q = D[C[s_1, \dots, s_n]]\gamma$ and $u = D[C[t_1, \dots, t_n]]\gamma$ for some ground γ which respects φ , or (b) $u = D[C[\vec{t}]]\gamma$ and

- $q = D[C[\vec{s}]]\gamma$. In case (a), $q \rightarrow_{\mathcal{R}}^* D\gamma[C\gamma[v_1, \dots, v_n]]$ and $u \rightarrow_{\mathcal{R}}^* D\gamma[C\gamma[w_1, \dots, w_m]]$. If each $v_i = w_i$, then clearly $q \rightarrow_{\mathcal{R}}^* \leftarrow_{\mathcal{R}}^* u$. Otherwise, $(\varphi \wedge \neg(s_1 = t_1 \wedge \dots \wedge s_n = t_n))\gamma$ is valid, so we easily get the desired $q \rightarrow_{\mathcal{R}}^* \leftrightarrow_{\mathcal{E}'} \leftarrow_{\mathcal{R}}^* u$. Case (b) is symmetric.
- (DISPROVE) In this case we do not have $(\mathcal{E}, \mathcal{H}, b) \vdash_{\text{ri}} (\mathcal{E}', \mathcal{H}', b')$.
 - (CONSTRUCTOR) Let $s = f(s_1, \dots, s_n)$, $t = f(t_1, \dots, t_n)$, and suppose $C[s\gamma] \leftrightarrow_{\{s \approx t [\varphi]\}} C[f(t\gamma)]$, where γ is a substitution which respects φ . Since \mathcal{E}' contains all equations $s_i \approx t_i [\varphi]$, we have $C[s\gamma] = C[f(s_1\gamma, \dots, s_n\gamma)] \leftrightarrow_{\mathcal{E}'} C[f(t_1\gamma, \dots, t_n\gamma)] = C[t\gamma]$.
 - (POSTULATE) $\mathcal{E} \setminus \mathcal{E}' = \emptyset$, so there is nothing to prove!
 - (GENERALIZATION) Suppose that $s \approx t [\varphi]$ is replaced by $s' \approx t' [\psi]$. Suppose that $C[s\gamma] \leftrightarrow_{\{s \approx t [\varphi]\}} C[t\gamma]$ for some substitution γ which respects φ . Then there exists a substitution δ which respects ψ such that $C[s\gamma] = C[s'\delta] \leftrightarrow_{\{s' \approx t' [\psi]\}} C[t'\delta] = C[t\gamma]$. \square

LEMMA B.3. *Suppose that $(\mathcal{E}, \mathcal{H}, \text{flag}) \vdash_{\text{ri}} (\mathcal{E}', \mathcal{H}', \text{flag}')$ by any inference rule other than COMPLETENESS. If*

$$\rightarrow_{\mathcal{R} \cup \mathcal{H}} \subseteq (\rightarrow_{\mathcal{R}} \cdot \rightarrow_{\mathcal{R} \cup \mathcal{H}}^* \cdot \leftrightarrow_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}}^*)$$

on ground terms, then

$$\rightarrow_{\mathcal{R} \cup \mathcal{H}'} \subseteq (\rightarrow_{\mathcal{R}} \cdot \rightarrow_{\mathcal{R} \cup \mathcal{H}'}^* \cdot \leftrightarrow_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}'}^*)$$

on ground terms.

PROOF. It suffices to consider the case that EXPANSION is applied (for the other cases, we use Lemma B.2). Suppose that $s \rightarrow_{\mathcal{H}' \setminus \mathcal{H}} t$. Using that, by quasi-reductivity and termination, every ground term reduces to a ground constructor term, it follows from Lemma B.1 that there exist ground constructor terms s', t' such that $s \rightarrow_{\mathcal{R}}^* s' (\rightarrow_{\mathcal{R}} \cdot \leftrightarrow_{\mathcal{E}'}) t' \leftarrow_{\mathcal{R}}^* t$, and hence:

$$s (\rightarrow_{\mathcal{R}} \cdot \rightarrow_{\mathcal{R} \cup \mathcal{H}'}^* \cdot \leftrightarrow_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}'}^*) t$$

\square

LEMMA B.4. *Suppose that $(\mathcal{E}, \mathcal{H}, \text{flag}) \vdash_{\text{ri}} \dots \vdash_{\text{ri}} (\mathcal{E}', \mathcal{H}', \text{flag}')$. Then:*

- (1) $\leftrightarrow_{\mathcal{E}'} \subseteq (\rightarrow_{\mathcal{R} \cup \mathcal{H}'}^* \cdot \leftrightarrow_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}'}^*)$ on ground terms,
- (2) if $\rightarrow_{\mathcal{R} \cup \mathcal{H}} \subseteq (\rightarrow_{\mathcal{R}} \cdot \rightarrow_{\mathcal{R} \cup \mathcal{H}}^* \cdot \leftrightarrow_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}}^*)$ on ground terms, then

$$\rightarrow_{\mathcal{R} \cup \mathcal{H}'} \subseteq (\rightarrow_{\mathcal{R}} \cdot \rightarrow_{\mathcal{R} \cup \mathcal{H}'}^* \cdot \leftrightarrow_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}'}^*)$$

on ground terms, and

- (3) if $\mathcal{R} \cup \mathcal{H}$ is terminating, then so is $\mathcal{R} \cup \mathcal{H}'$.

PROOF. In the following, we will consider relations limited to ground terms only. We prove the statements by induction on the number of \vdash_{ri} -steps, where steps in the premise of a COMPLETENESS step are also counted. The base case is evident, so suppose $(\mathcal{E}, \mathcal{H}, \text{flag}) \vdash_{\text{ri}} (\mathcal{E}_1, \mathcal{H}_1, \text{flag}_1) \vdash_{\text{ri}}^* (\mathcal{E}', \mathcal{H}', \text{flag}')$.

- (1) If the first step uses inference rule COMPLETENESS, then $(\mathcal{E}, \mathcal{H}, \text{flag}) \vdash_{\text{ri}}^* (\mathcal{E}_1, \mathcal{H}_1, \text{INCOMPLETE})$ in fewer steps, so by the induction hypothesis:

$$\leftrightarrow_{\mathcal{E}'} \subseteq (\rightarrow_{\mathcal{R} \cup \mathcal{H}_1}^* \cdot \leftrightarrow_{\mathcal{E}_1} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}_1}^*)$$

If the first step uses another inference rule, this same property follows from Lemma B.2. By the induction hypothesis we have

$$\leftrightarrow_{\mathcal{E}_1} \subseteq (\rightarrow_{\mathcal{R} \cup \mathcal{H}'}^* \cdot \leftrightarrow_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}'}^*)$$

It follows from $\mathcal{H}_1 \subseteq \mathcal{H}'$ that

$$\leftrightarrow_{\mathcal{E}'} \subseteq (\rightarrow_{\mathcal{R} \cup \mathcal{H}'}^* \cdot \leftrightarrow_{\mathcal{E}_1} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}'}^*)$$

By replacing $\leftrightarrow_{\mathcal{E}_1}$ with $(\rightarrow_{\mathcal{R} \cup \mathcal{H}'}^* \cdot \leftrightarrow_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}'}^*)$, we thus obtain

$$\leftrightarrow_{\mathcal{E}} \subseteq (\rightarrow_{\mathcal{R} \cup \mathcal{H}'}^* \cdot \rightarrow_{\mathcal{R} \cup \mathcal{H}'}^* \cdot \leftrightarrow_{\mathcal{E}'} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}'}^* \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}'}^*)$$

- (2) Assume $\rightarrow_{\mathcal{R} \cup \mathcal{H}} \subseteq (\rightarrow_{\mathcal{R}} \cdot \rightarrow_{\mathcal{R} \cup \mathcal{H}}^* \cdot \leftrightarrow_{\mathcal{E}} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}}^*)$. By the induction hypothesis (in case of COMPLETENESS) or Lemma B.3 (otherwise),

$$\rightarrow_{\mathcal{R} \cup \mathcal{H}_1} \subseteq (\rightarrow_{\mathcal{R}} \cdot \rightarrow_{\mathcal{R} \cup \mathcal{H}_1}^* \cdot \leftrightarrow_{\mathcal{E}_1} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}_1}^*)$$

We complete by the induction hypothesis on $(\mathcal{E}_1, \mathcal{H}_1, \text{flag}_1) \vdash_{\text{ri}}^* (\mathcal{E}', \mathcal{H}', \text{flag}')$.

- (3) Trivial with the induction hypothesis, with the first step using either the induction hypothesis again (in case of COMPLETENESS), the definition of EXPANSION, or the observation that other inference rules do not alter \mathcal{H} . \square

Thus we obtain Lemma 4.33 or, equivalently, Lemma B.5, as the first part of Thm. 4.4.

LEMMA B.5. *If $(\mathcal{E}, \emptyset, \text{flag}) \vdash_{\text{ri}} \dots \vdash_{\text{ri}} (\emptyset, \mathcal{H}, \text{flag}')$, then every equation in \mathcal{E} is an inductive theorem of \mathcal{R} .*

PROOF. It is clear that $\rightarrow_{\mathcal{R}} \subseteq \rightarrow_{\mathcal{R} \cup \mathcal{H}}$. It follows from Lemma B.4 that:

- $\leftrightarrow_{\mathcal{E}} \subseteq \leftrightarrow_{\mathcal{E}} \subseteq \rightarrow_{\mathcal{R} \cup \mathcal{H}}^* \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}}^*$ on ground terms,
- $\rightarrow_{\mathcal{R} \cup \mathcal{H}} \subseteq \rightarrow_{\mathcal{R}} \cdot \rightarrow_{\mathcal{R} \cup \mathcal{H}}^* \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}}^*$ on ground terms, and
- $\mathcal{R} \cup \mathcal{H}$ is terminating.

By Lemma 4.32 (as equality is included in $\leftrightarrow_{\mathcal{R}}^*$) we find that $\leftrightarrow_{\mathcal{R}}^* = \leftrightarrow_{\mathcal{R} \cup \mathcal{H}}^*$, and hence $\leftrightarrow_{\mathcal{E}} \subseteq \leftrightarrow_{\mathcal{R}}^*$, on ground terms. We complete with Lemma 4.31. \square

Moving on to *disproving*, we need two auxiliary lemmas:

LEMMA B.6. *If \mathcal{R} is confluent and $(\mathcal{E}, \mathcal{H}, \text{COMPLETE}) \vdash_{\text{ri}} \perp$, then \mathcal{E} contains an equation $s \approx t [\varphi]$ which is not an inductive theorem.*

PROOF. By confluence and termination together, we can speak of *the* normal form $u \downarrow_{\mathcal{R}}$ of any term u ; if u is ground, then by quasi-reductivity its normal form is a ground constructor term. A property of confluence is that if $w \leftrightarrow_{\mathcal{R}}^* q$, then $w \downarrow_{\mathcal{R}} = q \downarrow_{\mathcal{R}}$. So, it suffices to prove that for some $s \approx t [\varphi] \in \mathcal{E}$ there is a ground constructor substitution γ which respects this equation, such that $s\gamma$ and $t\gamma$ have distinct normal forms.

The only inference rule that could be used to obtain $(\mathcal{E}, \mathcal{H}, \text{COMPLETE}) \vdash_{\text{ri}} \perp$ is DISPROVE, so $\mathcal{E} = \mathcal{E}' \cup \{s \approx t [\varphi]\}$ and one of the following holds:

- (1) $s, t \in \text{Terms}(\Sigma_{\text{theory}}, \mathcal{V})$ with $\varphi \wedge s \neq t$ satisfiable. That is, there is a substitution γ mapping all variables in the equation to values, such that $\varphi\gamma$ is valid and $s\gamma$ and $t\gamma$ reduce to different values by $\rightarrow_{\text{calc}}$. We are done since all values are normal forms.
- (2) $s = f(s_1, \dots, s_n)$ and $t = g(t_1, \dots, t_m)$ with f and g different constructors, and φ is satisfiable, so there is a substitution δ mapping all variables in φ to values such that $\varphi\delta$ is valid. Let γ be an extension of δ which additionally maps all other variables in s, t to ground terms (by assumption, ground instances of all variables exist). Then $\varphi\gamma$ is still valid, and $s\gamma$ and $t\gamma$ are ground terms with $s\gamma \rightarrow_{\mathcal{R}}^* (s\gamma) \downarrow_{\mathcal{R}} = f((\vec{s}\gamma) \downarrow_{\mathcal{R}}) \neq g((\vec{t}\gamma) \downarrow_{\mathcal{R}}) = (t\gamma) \downarrow_{\mathcal{R}} \leftarrow_{\mathcal{R}}^* t\gamma$.
- (3) $s : \iota$ is a variable not occurring in φ , φ is satisfiable, there are at least two different constructors f, g with output sort ι and either t is a variable distinct from s or t has a constructor symbol at the root. By satisfiability of φ , a substitution δ exists whose domain does not contain s , with $\varphi\delta$ valid. If t is a variable, let γ be an extension of δ mapping s to some ground term rooted by f and t to a ground term rooted by g (by assumption ground instances always exist). If $t = f(\vec{t})$, then let γ be an extension of δ mapping s to some ground term rooted by g and mapping all other variables in t to ground terms as well. Either way, $\varphi\gamma$ is valid and $(s\gamma) \downarrow_{\mathcal{R}} \neq (t\gamma) \downarrow_{\mathcal{R}}$. \square

LEMMA B.7. *Suppose that $\rightarrow_{\mathcal{R} \cup \mathcal{H}}$ is terminating and that $\rightarrow_{\mathcal{R} \cup \mathcal{H}} \subseteq \rightarrow_{\mathcal{R}} \cdot \rightarrow_{\mathcal{R} \cup \mathcal{H}}^*$. $\cdot \leftrightarrow_{\mathcal{E}} \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}}^*$. If, moreover, \mathcal{R} is confluent, $(\mathcal{E}, \mathcal{H}, \text{COMPLETE}) \vdash_{\text{ri}} (\mathcal{E}', \mathcal{H}', \text{COMPLETE})$, and $\leftrightarrow_{\mathcal{E}} \cup \leftrightarrow_{\mathcal{H}} \subseteq \leftrightarrow_{\mathcal{R}}^*$ on ground terms, then $\leftrightarrow_{\mathcal{E}'} \cup \leftrightarrow_{\mathcal{H}'} \subseteq \leftrightarrow_{\mathcal{R}}^*$ on ground terms.*

PROOF. Assume that all conditions are satisfied; we consider the inference rule used to derive $(\mathcal{E}, \mathcal{H}, \text{COMPLETE}) \vdash_{\text{ri}} (\mathcal{E}', \mathcal{H}', \text{COMPLETE})$.

First, suppose the rule used was **COMPLETENESS**, so $(\mathcal{E}, \mathcal{H}, \text{COMPLETE}) \vdash_{\text{ri}}^* (\mathcal{E}', \mathcal{H}', \text{INCOMPLETE})$ and $\mathcal{E}' \subseteq \mathcal{E}$. As we have assumed that $\leftrightarrow_{\mathcal{E}} \cup \leftrightarrow_{\mathcal{H}} \subseteq \leftrightarrow_{\mathcal{R}}^*$, certainly $\leftrightarrow_{\mathcal{E}'} \subseteq \leftrightarrow_{\mathcal{E}} \subseteq \leftrightarrow_{\mathcal{R}}^*$. As for $\leftrightarrow_{\mathcal{H}'}$, Lemma B.4 gives us that $\rightarrow_{\mathcal{R} \cup \mathcal{H}'} \subseteq \rightarrow_{\mathcal{R}} \cdot \rightarrow_{\mathcal{R} \cup \mathcal{H}'}$. $\cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}'}^* \cdot \leftarrow_{\mathcal{R} \cup \mathcal{H}'}$, so (using again that $\leftrightarrow_{\mathcal{E}'} \subseteq \leftrightarrow_{\mathcal{R}}$ and that $\leftarrow_{\mathcal{R} \cup \mathcal{H}'} \subseteq \leftarrow_{\mathcal{E}'}$) we can apply Lemma 4.32 and termination of $\rightarrow_{\mathcal{R} \cup \mathcal{H}'}$ to obtain $\leftrightarrow_{\mathcal{H}'} \subseteq \leftrightarrow_{\mathcal{R} \cup \mathcal{H}'} \subseteq \leftrightarrow_{\mathcal{R}}^*$.

If a different rule was applied, then each element in \mathcal{H}' either also belongs to \mathcal{H} or (in the case of **EXPANSION**) corresponds to an equation in \mathcal{E} . Thus, $\leftrightarrow_{\mathcal{H}'} \subseteq \leftrightarrow_{\mathcal{E} \cup \mathcal{H}} \subseteq \leftrightarrow_{\mathcal{R}}^*$. So let $s \approx t [\varphi] \in \mathcal{E}' \setminus \mathcal{E}$; we must see that $\leftrightarrow_{\{s \approx t [\varphi]\}} \subseteq \leftrightarrow_{\mathcal{R}}^*$ on ground terms. By Lemma 4.31, it suffices if for all ground constructor substitutions γ which respect this equation, $s\gamma \leftrightarrow_{\mathcal{R}}^* t\gamma$. We fix γ and use a case analysis on the applied inference rule.

- (**SIMPLIFICATION**). There is $s' \simeq t' [\varphi'] \in \mathcal{E}$ such that $s' \approx t' [\varphi'] \rightarrow_{\mathcal{R} \cup \mathcal{H}} s \approx t [\varphi]$ at position $1 \cdot p$. By Thm. 2.20, we can find δ which respects φ' such that $s'\delta \rightarrow_{\mathcal{R} \cup \mathcal{H}} s\gamma$ at position p and $t'\delta = t\gamma$. As $\rightarrow_{\mathcal{R}} \cup \rightarrow_{\mathcal{H}} \subseteq \leftrightarrow_{\mathcal{R}}^*$ by the assumption, $s\gamma \leftrightarrow_{\mathcal{R}}^* s'\delta \leftrightarrow_{\mathcal{E}} t'\delta = t\gamma$, which suffices because $\leftrightarrow_{\mathcal{E}} \subseteq \leftrightarrow_{\mathcal{R}}^*$.
- (**DELETION**). No equations are added in this case.
- (**EXPANSION**). There is $s' \simeq t' [\varphi'] \in \mathcal{E}$ such that $s \approx t [\varphi] \in \text{Expd}(s' \approx t' [\varphi'], p)$ for some p . By Lemma B.1(2), we have $s\gamma (\leftarrow_{\mathcal{R}} \cdot \leftrightarrow_{\mathcal{E}}) t\gamma$, which suffices because $\leftrightarrow_{\mathcal{E}} \subseteq \leftrightarrow_{\mathcal{R}}^*$.
- (**EQ-DELETION**) $s \simeq t [\varphi'] \in \mathcal{E}$, where $\varphi = \varphi' \wedge \neg(s_1 = t_1 \wedge \dots \wedge s_n = t_n)$, and $s = C[s_1, \dots, s_n]$, $t = C[t_1, \dots, t_n]$ for some C, \vec{s}, \vec{t} . Since any substitution which respects φ also respects φ' , we must have $s\gamma \leftrightarrow_{\mathcal{E}} t\gamma$, so $s\gamma \leftrightarrow_{\mathcal{R}}^* t\gamma$.
- (**DISPROVE**) A reduction with this rule does not have the required form.
- (**CONSTRUCTOR**) There is $f(\dots, s, \dots) \approx f(\dots, t, \dots) [\varphi] \in \mathcal{E}$, and by assumption $f(\dots, s, \dots)\gamma \leftrightarrow_{\mathcal{R}}^* f(\dots, t, \dots)\gamma$. By confluence, this means that $f(\dots, s\gamma, \dots) \downarrow_{\mathcal{R}} = f(\dots, t\gamma, \dots) \downarrow_{\mathcal{R}}$, which implies that $(s\gamma) \downarrow_{\mathcal{R}} = (t\gamma) \downarrow_{\mathcal{R}}$.
- (**POSTULATE, GENERALIZATION**) A reduction with these rules does not have the form required by the lemma (as the **COMPLETE** flag is removed). \square

This leads to the second part of Thm. 4.4, which largely corresponds to Lemma 4.34:

LEMMA B.8. *If \mathcal{R} is confluent and $(\mathcal{E}, \emptyset, \text{COMPLETE}) \vdash_{\text{ri}} \dots \vdash_{\text{ri}} \perp$, then there is some equation in \mathcal{E} which is not an inductive theorem of \mathcal{R} .*

PROOF. If $(\mathcal{E}, \emptyset, \text{COMPLETE}) = (\mathcal{E}_1, \mathcal{H}_1, \text{flag}_1) \vdash_{\text{ri}} \dots \vdash_{\text{ri}} (\mathcal{E}_n, \mathcal{H}_n, \text{flag}_n) \vdash_{\text{ri}} \perp$, then we easily see that $\text{flag}_i = \text{COMPLETE}$ for all i . By Lemma B.6, \mathcal{E}_n contains an equation $s \approx t [\varphi]$ which is not an inductive theorem. Then $\leftrightarrow_{\mathcal{E}_n} \not\subseteq \leftrightarrow_{\mathcal{R}}^*$ on ground terms. By Lemma B.7, Lemma B.4, and induction on $n - i$, this means that $\leftrightarrow_{\mathcal{E}} \cup \leftrightarrow_{\emptyset} \not\subseteq \leftrightarrow_{\mathcal{R}}^*$ on ground terms, so by Lemma 4.31, not all $e \in \mathcal{E}$ are inductive theorems. \square

Proof of Thm. 4.4. Immediately by Lemmas B.5 and B.8. \square

C. SIMPLE SUM

To demonstrate the difference in power between our technique and earlier work, even when not considering advanced data structures which were not supported in [Nakabayashi et al. 2010] or [Falke and Kapur 2012], we have included an example that can be handled with the technique in this paper (and is automatically proved by Ctrl), but not with [Nakabayashi et al. 2010] or [Falke and Kapur 2012] (the latter of which is not surprising, as it does not use any lemma generation at all).

Example C.1. In the programming course in Nagoya, students in the first lecture were asked to implement a function `sum` which computes the summation from 0 to a given non-negative integer x . The teacher’s reference implementation was:

```
int sum(int x) {
    int z = 0;
    for (int i = 1; i <= x; i++) {
        z += i;
    }
    return z;
}
```

Some of the students solved (or tried to solve) this in the clever way instead:

```
int sum1(int x) {
    return x * (x + 1) / 2;
}
int sum2(int x) {
    return x * (x - 1) / 2;
}
```

To stay close to the transformation from [Nakabayashi et al. 2010] (which does not use the return and error symbols), we consider the following translation:

$$\begin{aligned} \text{sum}(x) &\rightarrow u(x, 1, 0) \\ u(x, i, z) &\rightarrow u(x, i + 1, z + i) \quad [i \leq x] \\ u(x, i, z) &\rightarrow z \quad [i > x] \\ \text{sum1}(x) &\rightarrow x * (x + 1) \text{ div } 2 \\ \text{sum2}(x) &\rightarrow x * (x - 1) \text{ div } 2 \end{aligned}$$

Our implementation succeeds in proving that $\text{sum}(n) \approx \text{sum1}(n)$ [$n \geq 0$] is an inductive theorem and that $\text{sum}(n) \approx \text{sum2}(n)$ [$n \geq 0$] is not. We also succeed on the translation using the methods in the current paper. On the other hand, the method in [Nakabayashi et al. 2010] fails to prove or disprove these claims.

D. SOME EXAMPLES WE CANNOT HANDLE.

To demonstrate the kind of problems Ctrl cannot yet handle, we compare a recursive definition `sum` of the function $n \mapsto \sum_{i=1}^n i$ with three iterative implementations.

<pre>int sum(n) { if (n < 0) return 0; return n + sum(n-1); }</pre>	<pre>int sum1(n) { int i = 0, j = 0, sum = 0; for (; i <= n; i++, j++) sum += j; return len; }</pre>
<pre>int sum2(int n){ int i, sum=0; for (i=n; i>=0; i--) sum=sum+i; return sum; }</pre>	<pre>int sum3(n) { int ret = 0; for (int i = 0; i <= n; i++) for (int j = 0; j < i; j++) ret++; return ret; }</pre>

Equivalence between `sum` and each of `sum1`, `sum2` and `sum3` fails for the three main reasons discussed in § 6.2. For `sum1`, generalizing the initialization variables loses the information that always $i = j$. For `sum2`, our main generalization method (§ 5.1) does not apply because we do not recognize $i = n$ as an initialization. For `sum3`, our strategy fails because the two loop counters are generalized together.

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