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To cite this version:
Isabelle Gnaedig, Hélène Kirchner. Computing Constructor Forms with Non Terminating Rewrite Programs - Extended version -. [Research Report] 2006, pp.17. inria-00113146

HAL Id: inria-00113146
https://inria.hal.science/inria-00113146
Submitted on 10 Nov 2006

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Computing Constructor Forms with Non Terminating Rewrite Programs
- Extended version -

Isabelle Gnaedig
LORIA-INRIA
615, rue du Jardin Botanique, BP 101,
F-54602 Villers-lès-Nancy Cedex
Isabelle.Gnaedig@loria.fr

Hélène Kirchner
LORIA-CNRS
Campus Scientifique, BP 239,
F-54506 Vandoeuvre-lès-Nancy Cedex
Helene.Kirchner@loria.fr

Abstract
In the context of the study of rule-based programming, we focus in this paper on the property of \(C\)-reducibility, expressing that every term reduces to a constructor term on at least one of its rewriting derivations. This property implies completeness of function definitions, and enables to stop evaluations of a program on a constructor form, even if the program is not terminating. We propose an inductive procedure proving \(C\)-reducibility of rewriting. The rewriting relation on ground terms is simulated through an abstraction mechanism and narrowing. The induction hypothesis allows assuming that terms smaller than the starting terms rewrite into a constructor term. The existence of the induction ordering is checked during the proof process, by ensuring satisfiability of ordering constraints. The proof is constructive, in the sense that the branch leading to a constructor term can be computed from the proof trees establishing \(C\)-reducibility for every term.

Categories and Subject Descriptors F.3.1 [LOGICS AND MEANINGS OF PROGRAMS]: Specifying and Verifying and Reasoning about Programs—Logics of programs, Mechanical verification, Specification techniques; F.4.2 [MATHEMATICAL LOGIC AND FORMAL LANGUAGES]: Grammars and Other Rewriting Systems; F.4.3 [MATHEMATICAL LOGIC AND FORMAL LANGUAGES]: Formal Languages—Algebraic language theory; I.1.3 [SYMBOLIC AND ALGEBRAIC MANIPULATION]: Languages and Systems—Evaluation strategies, Substitution mechanisms; I.2.3 [ARTIFICIAL INTELLIGENCE]: Deduction and Theorem Proving—Deduction, Inference engines, Mathematical induction; D.3.1 [PROGRAMMING LANGUAGES]: Formal Definitions and Theory; D.2.4 [SOFTWARE ENGINEERING]: Software/Program Verification—Correctness proofs, Formal methods, Validation

General Terms Algorithms, Languages, Verification

Keywords Sufficient Completeness, Constructor, Abstraction, Narrowing, Ordering Constraint, Termination

1. Introducing the problem
Sufficient completeness plays an important role in algebraic specifications, as well as in rewriting-based programs. In both contexts, terms are built on operators, among them we can distinguish constructors. Constructors are basic symbols, allowing us to describe the values of computations. The other symbols, called defined symbols, represent functions defined on these values. One is naturally led to ask for a specification or a program, whether the functions are sufficiently defined to warrant that to every expression or term corresponds a value.

From the point of view of specifications, where properties are described by equations, sufficient completeness ensures that every term is equivalent to a term built on constructors, called constructor term or constructor form. It allows inductive proofs, in particular by consistency methods [7]. Proof assistants like Coq or PVS include decision or semi-decision procedures based on rewrite rules and rely on complete definitions of functions.

From the point of view of programming, sufficient completeness ensures that a program produces a completely computed form for every data. Studying the property in this context is pertinent since rule-based systems have recently gained considerable interest with the development of efficient compilers. Now, systems like ASF+SDF [24], Maude [5, 12], Cafe-OBJ [17], ELAN [1], or TOM [28], are used for various applications like constraint solving, protocol verification, modeling of biological or chemical systems, and more.

Proving sufficient completeness is undecidable in general. It has already been widely studied, for example in [20, 29, 25, 22, 6, 21, 27, 2, 3], but most of the time, the proposed approaches for proving the property need restrictions like termination and confluence.

The property is strongly related to ground reducibility, which expresses that every ground instance of a term is reducible. Indeed, it is equivalent to ground reducibility of all patterns \(f(x_1, \ldots, x_m)\) built on a defined symbol \(f\), provided the rewrite system (RS in short) is terminating, confluent, and the normal form of a constructor term is again a constructor term [21]. Under these conditions, techniques developed for proving ground reducibility hold for sufficient completeness as well [30, 23, 21, 26, 4, 8].

In this paper, we address the problem of sufficient completeness from the programming point of view, and we go beyond the previous usual restrictions.

We observe the case where a program or a RS can be neither confluent, nor terminating, and we study its evaluations. We do not suppose other restrictions used in the domain as the constructor preserving property, or the absence of relation between constructors. The question is to know whether at least one evaluation of a given data
results in a completely evaluated result; in other words, whether for every  
ground term, there exists a rewriting chain that eventually reaches a constructor  
term, even if the chains do not converge to a single term, and are infinite. We call this  
property C-reducibility.

Since several years, we have been studying properties of  
rewriting in the context of programming, where a rewrite  
system interpreter or compiler uses a specific strategy for  
applying rewrite rules, most frequently an innermost strategy  
that consists in rewriting always at the lowest possible  
positions, and corresponds to a call-by-value evaluation of  
functions. We observe rewriting in considering its properties  
from an inductive point of view. We already have proposed  
proof procedures for the termination property, in the specific case of rewriting strategies,  
respectively for the innermost [19], the outermost [14] and local strategies [13]. We also have 
given a constructive proof technique for weak innermost termination, enabling to reach the  
terminating branches [15]. Here, C-reducibility is tackled with a similar inductive  
approach.

It is not based on analyzing the patterns issued from the  
RS, as do most of existing methods for sufficient completeness,  
but on observing the rewriting process starting from any term to see whether it is eventually transformed into a  
constructor term. For that, like for our previous procedures for  
termination, we reason by induction on the property itself,  
assuming that the terms t smaller than the starting terms for an induction ordering are C-reducible.

This approach needs a noetherian ordering on terms, used  
in the induction principle. In contrast to classical induction  
proofs where the ordering is given, we do not need to define it a priori. We only have to check its existence by ensuring  
satisfiability of ordering constraints incrementally set along the proof.

By alternatively applying an abstraction mechanism and  
well-covered narrowing steps, we generate proof trees representing at least one rewriting derivation for every ground  
term. The proof is constructive, in the sense that the rewriting branch leading to a constructor term can be computed from the proof trees establishing C-reducibility for every term, thus avoiding costly breadth-first computations to reach the constructor form. Although our method differs from previous approaches, we naturally encounter basic notions already used for proving sufficient completeness, as unification, used in [27], or covering properties used in many works in the domain, for example in [25, 27]. In [3], pattern trees are also developed, but under the assumption that the RS is terminating and ground confluent.

Before we detail our procedure, we link C-reducibility to close properties like strong C-reducibility, expressing the existence of a constructor form on every rewriting chain, and their variants in the case of innermost rewriting. Sufficient completeness and ground reducibility are also considered in this comparison. In particular, C-reducibility directly implies sufficient completeness. In addition, we justify that ground reducibility just requires weak termination (expressing that every term has at least one terminating rewriting derivation) to imply sufficient completeness, thus weakening the condition of the well-known theorem of [21]. We also note that as the covering property we require is stronger than ground reducibility, this property directly implies sufficient completeness if we suppose weak termination. Thus, with respect to the sufficient completeness property, the proof technique we propose is interesting:

- obviously when the RS is not even weakly terminating,
- when it is terminating, or just weakly terminating, but
  we have no technique to prove it. The weak termination proof is difficult to handle in general, and to our knowledge, our proof procedure [15] is the only one existing today. On the following small example, weakly terminating, and which is a classical definition of the booleans, enriched with a rule expressing the double application of not on and, and deliberately oriented in the divergent direction, it fails:

\[
\begin{align*}
\text{and}(1, x) & \rightarrow x \\
\text{and}(0, x) & \rightarrow 0 \\
o r(1, x) & \rightarrow 1 \\
o r(0, x) & \rightarrow x \\
\text{and}(1, x) & \rightarrow \text{not}(\text{not}(\text{and}(1, x))) \\
\text{not}(1) & \rightarrow 0 \\
\text{not}(0) & \rightarrow 1 \\
\text{not}(\text{and}(x, y)) & \rightarrow \text{or}(\text{not}(x), \text{not}(y)).
\end{align*}
\]

• as an alternative of the termination proof, as on the following more realistic example of computation of quotient, which is innermost terminating:

\[
\begin{align*}
\text{quot}(0, s(y), s(z)) & \rightarrow 0 \\
\text{quot}(s(x), s(y), z) & \rightarrow \text{quot}(x, y, z) \\
\text{quot}(x, 0, s(z)) & \rightarrow s(\text{quot}(x, s(z), s(z))) \\
\text{quot}(x, y, 0) & \rightarrow \text{error} \\
\text{quot}(\text{error}, y, z) & \rightarrow \text{error} \\
\text{quot}(x, \text{error}, z) & \rightarrow \text{error} \\
\text{quot}(x, y, \text{error}) & \rightarrow \text{error}.
\end{align*}
\]

In addition, if the RS has no rule with a constructor left-hand side, our proof technique establishes weak termination at the same time. This is the case for the two examples above.

In Section 2, the background is presented. In Section 3, we link known notions of completeness to C-reducibility. Section 4 introduces the basic concepts of the inductive proof mechanism. In Section 5, our procedure is defined with inference rules and a strategy to apply them. In Section 6, we explain how to extract a rewriting chain leading to a constructor term, from the proof of C-reducibility of terms.

2. The background

We assume that the reader is familiar with the basic definitions and notations of term rewriting given for instance in [11]. \( \mathcal{T}(\mathcal{F}, \mathcal{X}) \) is the set of terms built from a given set \( \mathcal{F} \) of function symbols having an arity \( \text{ar} = n \in \mathbb{N} \), and a set \( \mathcal{X} \) of variables denoted \( x, y, \ldots \). In this paper, we only consider finite sets of symbols. \( \mathcal{T}(\mathcal{F}) \) is the set of ground terms (without variables). Symbols of arity 0 are called constants. Positions in a term are represented as sequences of integers. The empty sequence \( c \) denotes the top position. The notation \( t_{i_1} \ldots t_{i_p} \) stands for the subterm of \( t \) at position \( p \). To emphasize that \( u \) contains subterms \( t_{j_1} \ldots t_{j_p} \) respectively at positions \( \{i_1, \ldots, i_p\} \), we write \( u[t_{j_1} \ldots t_{j_p}] \).

A substitution is an assignment from \( \mathcal{X} \) to \( \mathcal{T}(\mathcal{F}, \mathcal{X}) \), written \( \sigma = (x \mapsto t) \ldots (y \mapsto u) \). It uniquely extends to an endomorphism of \( \mathcal{T}(\mathcal{F}, \mathcal{X}) \). We identify a substitution \( \sigma = (x \mapsto t) \ldots (y \mapsto u) \) with the finite conjunction of equations \( x = t \) \& \ldots \& \( y = u \). The result of applying \( \sigma \) to a term \( t \in \mathcal{T}(\mathcal{F}, \mathcal{X}) \) is written \( \sigma(t) \) or \( t\sigma \). The domain of \( \sigma \), denoted \( \text{Dom}(\sigma) \), is the finite subset of \( \mathcal{X} \) such that \( x\sigma \neq x \). The range of \( \sigma \), denoted \( \text{Ran}(\sigma) \), is the set \( \bigcup_{x \in \text{Dom}(\sigma)} \text{Var}(\sigma x) \).

A ground substitution or instantiation is an assignment from \( \mathcal{X} \) to \( \mathcal{T}(\mathcal{F}) \). The composition of substitutions \( \sigma_1 \) fol-
lowed by $\sigma_2$ is denoted $\sigma_2 \sigma_1$. Given two substitutions $\sigma_1$ and $\sigma_2$, we write $\sigma_1 \leq \sigma_2$ iff $\exists \theta$ such that $\sigma_2 = \theta \sigma_1$. We say that $\sigma_1$ is a generalization of $\sigma_2$.

For $F$, we declare a set $C$ of constructors and a set $D$ of defined symbols. The terms of $T(C)$ are called constructor terms, or more briefly $C$-terms (or $C$-forms). Given a set $R$ of rewrite rules or term rewriting system on $T(F,X)$, the rewriting relation induced by $R$ is denoted by $\rightarrow_R$ ($\rightarrow$ if there is no ambiguity on $R$). We denote $s \rightarrow^{p,l,r,\sigma}_s t$ (where either $p$ or $l \rightarrow r$ or $\sigma$ may be omitted) if $s$ rewrites into $t$ at position $p$ with the rule $l \rightarrow r$ and the substitution $\sigma$, i.e., $s = s[\sigma[p]]$ and $t = s[\sigma[r]]$. The reflexive transitive closure of the rewriting relation induced by $R$ is denoted by $\rightarrow^*_R$.

An ordering $\succ$ on $T(F,X)$ is said to be noetherian iff there is no infinite decreasing chain for this ordering. It is monotone iff for every pair of terms $t, t'$ of $T(F,X)$, for every context $f(\ldots,t\ldots)$, $t \succ t'$ implies $f(\ldots,t'\ldots) \succ f(\ldots,t\ldots)$. It has the subterm property iff for every $t$ of $T(F,X)$, $t \succ f(\ldots,t\ldots)$. It has the constructor subterm property iff for every $g(t_1,\ldots,t_n)$ of $T(F,X)$ with $g \in C$, we have $g(t_1,\ldots,t_n) \succ t_1,\ldots,t_n$. Note that if $\succ$ is monotone and has the subterm property, as $F$ is finite, then it is noetherian.

A $RS$ $R$ (innermost) terminates iff every (innermost) rewriting derivation (or derivation chain) for $R$ is finite. It weakly terminates if at least one derivation chain starting from every term is finite. It is confluent iff for every term $t$ such that $t \rightarrow^*_u$ and $t \rightarrow^*_v$, there exists $t'$ such that $u \rightarrow^* t'$ and $v \rightarrow^* t'$. It is constructor preserving iff for each rule $l \rightarrow r$ of $R$, whenever $l \in T(C)$, then $r \in T(C)$.

3. $C$-reducibility and related properties

Before we compare them, let us now formally define the properties presented in the introduction.

**Definition 1.** Let $R = \{ l_i \rightarrow r_i, i \in [1..m]\}$ be a $RS$ on $T(F,X)$. The equational theory associated to $R$ is the equational theory induced by the set of equations $E = \{ l_i = r_i, i \in [1..m]\}$. It is also noted $E$.

**Definition 2.** Let $E$ be an equational theory on $T(F,X)$, where $F = C \cup D$. $E$ is sufficiently complete (with respect to $D$) if for every term $t \in T(F)$, there is a term $u \in T(C)$ such that $t \equiv_E u$.

**Definition 3.** Let $R$ be a $RS$ on $T(F,X)$. A term $t \in T(F,X)$ is ground reducible if every ground instance of $t$ is reducible with $R$.

**Definition 4.** Let $R$ be a $RS$ on $T(F,X)$, where $F = C \cup D$. A ground term is (innermost) (strongly) $C$-reducible iff it reduces into a $C$-term on at least one of its (resp. all) (innermost) rewriting derivations. The $RS$ $R$ is said to be (innermost) (strongly) $C$-reducible if every term $t$ of $T(F)$ is (innermost) (strongly) $C$-reducible.

Let us make some remarks on this property. If $R$ does not have rules whose left-hand sides are $C$-terms, the reached $C$-forms are normal forms.

If there are such rules in $R$, which is a case of relations between constructors, the $C$-forms may be reducible. Then, either $R$ is constructor-preserving and the next terms in the rewriting derivations are still $C$-terms. Or $R$ is not constructor-preserving and terms with defined symbols can be introduced after the $C$-terms in the derivations.

In non terminating cases, $C$-reducibility allows us to introduce a weak pseudo-termination notion, expressing that the evaluation of every data stops on the first constructor encountered form, ensured to exist on at least one rewriting branch.

Let us now observe the following diagram, summarizing the links between the different properties we consider. “Ground reducibility” stands for ground reducibility of all terms of $T(F)$. Arrows represent implications between properties; they are labeled by the necessary conditions on the RS required by the implications. All links are quite obvious, except those between ground reducibility and sufficient completeness, given by Theorem 7 in [21].

From the diagram, we immediately see the following theorem, which weakens the terminating condition of the “if” part of Theorem 7 in [21].

**Theorem 1.** Let $R$ be an innermost terminating $RS$ on $T(F,X)$, where $F = C \cup D$. If every term $f(x_1,\ldots,x_m), f \in$
tree issued from a reference term
ground instance of a term
trelevant sets of ground terms.
the inductive steps, and abstraction constraints combined
formulas: ordering constraints, set to warrant the validity of
that are solutions of the constraint.
of ground terms: all ground instances of the current term,
proof trees are composed of a current term of
achieved through constraints. The nodes of the developed
subterms.
presses the application of the induction hypothesis on these
reasoning by induction allows us to only suppose the ex-
representing respectively any of their constructor forms.

t, the resulting term $u = \ell(X_j)_{i_1, \ldots, i_p}$ (where $i_1, \ldots, i_p$ are
the abstraction positions in $t$) is narrowed in all possible ways into terms $v$, according to the possible instances of
the constraint associated to $u$) with all possible rewrite rules.
This technique is inspired from the one we have pro-
posed for proving innermost termination. But in that case,
as abstracted subterms were representing innermost nor-
mal forms, alternating abstraction and innermost narrowing
steps allowed us to simulate all innermost rewriting deriv-
ations. Here, abstracted subterms do not generally represent
normal forms, and we use standard narrowing, so the process
only simulates some possible rewriting derivations, for the
standard rewriting relation (without strategy) among which
we try to prove the existence of a $C$-term.
However, we have to ensure that every ground term
represented by $u$ is $C$-reducible, and so we have to prove
the existence of at least one branch leading to a $C$-term for
every ground instance of $u$.
To exhibit such a branch for every ground instance of $u$
containing at least one defined symbol, the narrowing steps
have to cover (in a sense we will define later) all possible
ground instances of $u$.
Then $C$-reducibility of the ground instances of $t$ is re-
duced to $C$-reducibility of the ground instances of the terms
$v$. Now, if $\theta_{t, v} \succ \theta v$ for every ground substitution $\theta$ that is
a solution of the constraint associated to $v$, by the induction
hypothesis, $\theta v$ is supposed to be $C$-reducible. Else, the pro-
cess is iterated on $v$, until we get a term $t' \succ \theta t'$ such that either
$\theta_{t, t'} \succ \theta t'$, or $\theta t'$ is a $C$-term.
We now introduce some concepts to formalize and auto-
mate this mechanism.

4. Inductively proving $C$-reducibility
For proving that a term $t$ of $T(F)$ is $C$-reducible, we proceed
by induction on $T(F)$ with a nonempty ordering $\succ$, assum-
ing the property for any $t'$ such that $t \succ t'$. To warrant non
emptiness of $T(F)$ and $T(C)$, we assume that $C$ con-
tains at least a constant. The main intuition is to observe
rewriting derivations starting from a ground term $t \in T(F)$
which is any instance of a term $g(x_1, \ldots, x_m) \in T(F, X)$,
for some defined function symbol $g \in D$, and variables
$x_1, \ldots, x_m$. Proving the $C$-reducibility property on ground
terms amounts to prove that at least one derivation leads to
a constructor term.
Rewriting derivations are simulated, using a lifting mech-
anism, by a proof tree developed from $g(x_1, \ldots, x_m) \in T(F, X)$,
for every $g \in D$, by alternatively using two main concepts,
namely narrowing and abstraction. More pre-
cisely, narrowing schematizes the rewriting possibilities of
terms. Abstraction simulates the reduction of subterms in
the derivations until these subterms become $C$-terms. It ex-
presses the application of the induction hypothesis on these
subterms.
The schematization of ground rewriting derivations is
achieved through constraints. The nodes of the developed
proof trees are composed of a current term of $T(F, X)$, and
a set of ground substitutions represented by a constraint
progressively built along the successive abstraction and nar-
rowing steps. Each node in an abstract tree schematizes a set
of ground terms: all ground instances of the current term,
that are solutions of the constraint.
The constraint is in fact composed of two kinds of for-
mulas: ordering constraints, set to warrant the validity of
the inductive steps, and abstraction constraints combined to
narrowing substitutions, which effectively define the rel-
evant sets of ground terms.
As said previously, we consider any term $t$ of $T(F)$ as a
ground instance of a term $t$ of $T(F, X)$ occurring in a proof
tree issued from a reference term $t_{ref}$.

- first, some subterms $t_i |_{j}$ of the current term $t$ of the proof
tree are supposed to reduce into a $C$-term, by the induc-
tion hypothesis, if $\theta_{t, t_i} \succ \theta t_i$ for the induction ordering
$\succ$ and for every $\theta$ solution of the constraint associated
to $t$. They are replaced in $t$ by abstraction variables $X_j$
representing respectively any of their constructor forms.
Reasoning by induction allows us to only suppose the ex-
stistence of the constructor forms without explicitly com-
puting them. If the resulting term only contains construc-
tor symbols and abstraction variables, the initial term is
$C$-reducible. Else,
is that \( \triangleright \) is stable by substitution and \( t \triangleright t' \) for every constraint \( t \triangleright t' \) of \( C \).

The inductive termination proof on ground terms by application of Theorem 2 requires that the noetherian induction ordering has the constructor subterm property. Simplification orderings fulfill such conditions. So in practice, it is sufficient to find a simplification ordering \( \triangleright \) such that \( t \triangleright t' \) for any constraint \( t \triangleright t' \) of \( C \).

Solving ordering constraints in finding simplification orderings is a well-known problem in rewriting. The simplest and automatable way to proceed is to test simple existing orderings like the subterm ordering, the Recursive Path Ordering, or the Lexicographic Path Ordering. This is often sufficient for the constraints considered here. If these simple orderings are not powerful enough, automatic solvers like Cine2 [9] can provide adequate polynomial orderings.

Other constraints are introduced by the abstraction mechanism. As said previously, to abstract a term \( t \) at positions \( i_1, \ldots, i_p \), where the \( t_j, j \in \{i_1, \ldots, i_p \} \) are supposed to have a reduced C-form \( t_j \mid_{i_j} \), we replace the \( t_j \) by abstraction variables \( X_j \) representing respectively one of their possible C-forms. Let us define these special variables more formally.

**Definition 6.** Let \( N \) be a set of new variables disjoint from \( X \). Symbols of \( N \) are called C-variables (or constructor variables). Substitutions and instantiations are extended to \( T(\mathcal{F}, X \cup N) \) in the following way. Let \( X \subseteq N; \) for any substitution \( \sigma \) (resp. instantiation \( \theta \)) such that \( X \in \text{Dom}(\sigma), \sigma X \subseteq T(C,N) \) (resp. \( \theta X \subseteq T(C) \) and then \( \text{Var}(\sigma X) \subseteq N \).

**Definition 7 (abstraction).** The term \( t(t_j \mid_{i_1}, \ldots, i_p) \) is said to be abstracted into the term \( u \) (called abstraction of \( t \)) at positions \( i_1, \ldots, i_p \) iff \( u = t[X_j \mid_{i_1}, \ldots, i_p] \), where the \( X_j, j \in \{i_1, \ldots, i_p \} \) are fresh distinct C-variables.

C-reducibility on \( T(\mathcal{F}) \) is proved by reasoning on terms with abstraction variables, i.e. on terms of \( T(\mathcal{F}, X \cup N) \). Ordering constraints are extended to pairs of terms of \( T(\mathcal{F}, X \cup N) \). When subterms \( t_j \) are abstracted by \( X_j \), we state constraints on abstraction variables, called abstraction constraints, to express that their instances can only be C-terms, rewritten from the corresponding instances of \( t_j \). Initially, they are of the form \( t_j \mid_{i_1} = X \) where \( t \in T(\mathcal{F}, X \cup N) \), and \( X \in N \), but we will see later on how they will be combined with the substitutions used for the narrowing process.

### 4.2 Narrowing

After abstraction of the current term \( t \) into the term \( u \), if \( u \) contains defined symbols, we test whether the ground instances of \( u \) are reducible, according to the possible values of the instances of the \( X_j \). This is achieved by narrowing \( u \) in all possible ways. We recall the definition of narrowing.

**Definition 8 (narrowing).** A term \( t \in T(\mathcal{F}, X \cup N) \) narrows into a term \( t' \in T(\mathcal{F}, X \cup N) \) at the non-variable position \( p \), using the rule \( l \rightarrow r \in \mathcal{R} \) with the substitution \( \sigma \) which is written \( t \overset{\mathcal{R}, \sigma}{\rightarrow} t' \) iff

\[
\begin{align*}
\sigma(l) &= \sigma(t[p]) \quad \text{and} \quad t' &= \sigma(t[r])
\end{align*}
\]

where \( \sigma \) is the most general unifier of \( t \) at position \( p \) and \( l \).

A few remarks can be made on the choice of variables and on the domain of the substitutions generated during the proof process. It is always assumed that there is no variable in common between the rule and the term, i.e. that \( \text{Var}(t) \cap \text{Var}(\sigma) = \emptyset \). This requirement of disjoint variables is easily fulfilled by an appropriate renaming of variables in the rules when narrowing is performed. Observe that for the most general unifier \( \sigma \) used in the above definition, \( \text{Dom}(\sigma) \subseteq \text{Var}(t) \cap \text{Var}(\sigma) \) and we can choose \( \text{Ran}(\sigma) \cap \text{Var}(\sigma) = \emptyset \), thus introducing in the range of \( \sigma \) only fresh variables. Moreover, narrowing is only performed on terms \( t \in T(\mathcal{F}, N) \), since an abstracting step is first applied on the reference terms, of the form \( g(x_1, \ldots, x_n) \), replacing \( x_1, \ldots, x_m \in X \) by \( X_1, \ldots, X_m \in N \). Then from Definition 6 we infer that for the most general unifiers \( \sigma \) produced during the proof process, all variables of \( \text{Ran}(\sigma) \) are C-variables.

Notice also that in our process, we are interested in the narrowing substitution applied to the current term \( t \), but not in its definition on the variables of the left-hand side of the rule. So the narrowing substitutions we consider are restricted to the variables of the narrowed term \( t \).

As said before, we only narrow the term \( u \) if all narrowing possibilities represent at least one rewriting step for all possible ground instances of \( u \).

Else, the proof process stops with failure. But we cannot conclude that \( t \) is not C-reducible, since, although all ground instances of \( u \) are not reducible at that step, they may be reduced to a C-form by other rewriting derivations, not observed by the process, as shows the following example.

Let \( \mathcal{F} = \{f, a, b, c\}, C = \{b, c\}, \mathcal{R} = \{a \rightarrow b, f(f(b)) \rightarrow c, b \rightarrow f(b), f(c) \rightarrow c\} \). Abstracting the pattern \( f(x) \) gives \( f(X) \), where \( X \) only represents C-forms. Then, \( f(X) \) is only narrowed into \( c \) for \( X = c \). The narrowing substitution \( X = f(b) \) is not allowed since \( f(b) \not\in T(C,N) \). However, the other ground instance of \( f(X) \), which is \( f(b) \), is also C-reducible, but the rewriting chain \( f(b) \rightarrow f(f(b)) \rightarrow c \) is not represented by the abstract-narrow-based process.

**Definition 9.** A set of substitutions \( \Sigma \), is said to cover a term \( u \in T(\mathcal{F}, X \cup N) \), or to be u-covering iff for any ground instance \( \theta u \) of \( u \), \( \exists \Sigma \in \Sigma \) such that \( \text{Dom}(\sigma) \supseteq \text{Var}(u) \), and \( \theta \sigma = \mu \theta u \) for some ground substitution \( \mu \).

The following obvious proposition expresses that Definition 9 provides a sufficient condition for the previous requirement on narrowing.

**Proposition 1.** Let \( \mathcal{R} \) be a RS, \( u \) a term of \( T(\mathcal{F}, N) \), and \( \Sigma \) the set of narrowing substitutions of \( u \) for \( \mathcal{R} \). If \( \Sigma \) is u-covering, then every ground instance \( \alpha \mu u \) of \( u \) is such that \( \alpha \mu u \overset{\mathcal{P}, \mathcal{T}, \mathcal{R}, \beta \sigma}{\rightarrow} \alpha \mu v \) for some \( v \in T(\mathcal{F}, N) \).

As a first narrowing step is applied on the patterns \( g(X_1, \ldots, X_m), g \in \mathcal{D} \), issued from abstraction of the reference patterns \( g(x_1, \ldots, x_m) \), the development of each proof tree at least requires for the narrowing substitutions of \( g(X_1, \ldots, X_m) \) to cover \( g(X_1, \ldots, X_m) \).

The following two propositions allow us to show that the covering property of the patterns \( g(X_1, \ldots, X_m), g \in \mathcal{D} \) is stronger than usual ground reducibility (i.e. of the patterns \( g(x_1, \ldots, x_m) \)). The first one is an obvious consequence of Proposition 1.

**Proposition 2.** Let \( \mathcal{R} \) be a RS, \( g \in \mathcal{D}, X_1, \ldots, X_m \in N \). Let \( \Sigma \) be the set of narrowing substitutions of \( g(X_1, \ldots, X_m) \) with \( \mathcal{R} \). If \( \Sigma \) covers \( g(X_1, \ldots, X_m) \), then \( g(X_1, \ldots, X_m) \) is ground reducible.
The converse is not true, as shows the following example. Let $F = \{f, 0, 1\}$, $C = \{0, 1\}$, and $R = \{f(0) \rightarrow 0, 1 \rightarrow 0\}$. The term $f(X)$ is ground reducible since $f(0)$ and $f(1)$ are reducible. However, the set $\Sigma$ of narrowing substitutions of $f(X)$, equal to $\{\sigma = (X = 0)\}$ is not covering for $f(X)$: $f(1)$ is not a ground instance of $f(0)$.

**Proposition 3.** Let $R$ be a RS on $T(F, X)$, $X_1, \ldots, X_m \in N$, $x_1, \ldots, x_m \in X$. Then $f(X_1, \ldots, X_m)$ is ground reducible for every $f \in D$, iff $f(x_1, \ldots, x_m)$ is ground reducible for every $f \in D$.

Given a term $u$, a sufficient condition for a substitution set $\Sigma$ to be $u$-covering can be established as follows, provided the variables of the considered term $u$ are $C$-variables, whose ground instantiations can only be constructor terms. Let $P$ be the set of constructor patterns $\{f(Y_1, \ldots, Y_m) \mid f \in C, Y_1, \ldots, Y_m \in N\}$, with $\text{arity}(f) = m$. For $u \in T(F, N)$, let $\Sigma^u$ be the set of all possible pattern substitutions of $u$ i.e. the set $\{\sigma^u = (X_1 = t_1, \ldots, X_p = t_p) \mid (X_1, \ldots, X_p) = \text{Var}(u), t_1, \ldots, t_p \in P\}$.

Then $\Sigma$ is $u$-covering if for every $\sigma^u \in \Sigma^u$, there exists $\sigma \in \Sigma$ such that $\sigma^u = \sigma \theta$ for some substitution $\theta$; in other words, if for every $\sigma^u \in \Sigma^u$, there exists in $\Sigma$ a generalization of $\sigma^u$. This sufficient condition is automatable.

### 4.3 Cumulating constraints

Abstraction constraints have to be combined with the narrowing substitutions to characterize the ground terms schematized by the current term $t$ in the proof tree. Indeed, a narrowing branch on the current term $u$ with narrowing substitution $\sigma$ represents a rewriting branch for any ground instance of $\sigma u$.

In addition, $\sigma$ has to satisfy the constraints on variables of $u$, already set in $A$. So $\sigma$, considered as the narrowing constraint attached to the narrowing branch, is added to $A$. This leads to the introduction of abstraction constraint formulas.

**Definition 10.** An abstraction constrained formula (ACF in short) is a formula $\bigwedge \{t_i \leq c = t_i' \} \land \bigwedge \{x_i = u_i\}$, where $x_i \in X \cup N$, $t_i, t_i', u_i \in T(F, X \cup N)$.

**Definition 11.** An abstraction constrained formula $A = \bigwedge \{t_i \leq c = t_i' \} \land \bigwedge \{x_j = u_j\}$ is satisfiable iff there exists at least one instantiation $\theta$ such that $\bigwedge (\theta t_i \downarrow \leq t_i') \land \bigwedge (\theta x_i = \theta u_i)$. The instantiation $\theta$ is then said to satisfy the ACF $A$ and is called solution of $A$.

For a better readability on examples, we can propagate $\sigma$ into $A$ (by applying $\sigma$ to $A$), thus getting instantiated abstraction constraints of the form $t_i \downarrow \leq c = t_i'$ from initial abstraction constraints of the form $t_i \downarrow \leq c = X_i$.

An ACF is attached to each term $u$ in the proof trees; the ground substitutions solutions of $A$ define the instances of the current term $u$ in the proof tree, for which we are observing $C$-reducibility. When $A$ has no solution, the current node of the proof tree represents no ground term. Such nodes are then irrelevant for the proof. Detecting and suppressing them during a narrowing step allows us to control the narrowing mechanism. So we have the choice between generating only the relevant nodes of the proof tree, by testing satisfiability of $A$ at each step, or stopping the proof on a branch on an irrelevant node, by testing unsatisfiability of $A$.

Checking satisfiability of $A$ is in general undecidable, but it is often easy in practice to exhibit an instantiation satisfying it. Automatable sufficient conditions are also under study.

Unsatisfiability of $A$ is also undecidable in general, but automatable simple sufficient conditions can be used, as testing whether $A$ contains equalities $t_i \downarrow = u$, where $u$ contains at least a symbol of $D$.

In the following, we present the procedure dealing with satisfiability of $A$.

### 5. Inference rules for the inductive proof

We are now ready to describe the inference rules defining our mechanism. They transform a set $T$ of 3-tuples $(U, A, C)$ where $U = \{t\}$ or $\emptyset$, the current term whose $C$-reducibility has to be proved, $A$ is an abstraction constraint formula, $C$ is a conjunction of ordering constraints.

- The first rule abstracts the current term $t$ at given positions $i_1, \ldots, i_p$. The constraints $t_{\text{ref}} > t_{\{i_1, \ldots, i_p\}}$ are set, allowing us to suppose, by induction, the existence of $C$-forms for $t_{\{i_1, \ldots, i_p\}}$. Then, $t_{\{i_1, \ldots, i_p\}}$ are abstracted into abstraction variables $X_{i_1}, \ldots, X_{i_p}$. The abstraction constraint $t_{\{i_1 \downarrow \leq c = i_p\}} = X_{i_1}, \ldots, t_{\{i_p\} \downarrow \leq c = X_{i_p}}$ is added to the ACF $A$. We call that rule Abstract.

The abstraction positions are chosen so that the abstraction mechanism captures the greatest possible number of rewriting steps: then we abstract all greatest possible subterms of $t = f(t_1, \ldots, t_m)$. More concretely, we try to abstract $t_1, \ldots, t_m$ and, for each $t_i = g(t_{i_1}, \ldots, t_{i_k})$ that cannot be abstracted, we try to abstract $t_{i_1}, \ldots, t_{i_k}$, and so on. In the worst case, we are driven to abstract leaves of the term, which are either variables, or constants.

Note also that it is not useful to abstract subterms if they are in $T(C, N)$. Indeed, by Definition 6, every ground instance of such subterms is already a $C$-term.

- The second rule narrows the resulting term $u$, if it is not a term of $T(C, N)$, in all possible ways in one step, with all possible rewrite rules of the rewrite system $R$, and all possible substitutions $\sigma_1, \ldots, \sigma_m$, into terms $v_1, \ldots, v_k$, according to Definition 8, if $\{\sigma_1, \ldots, \sigma_m\}$ is $u$-covering. This step is a branching step, creating as many states as narrowing possibilities. The substitution $\sigma$ is integrated to $A$. This is the Narrow rule. As $\{\sigma_1, \ldots, \sigma_m\}$ is $u$-covering, the narrowing step effectively simulates at least one rewriting step for each possible ground instance of $u$.

- We finally have a Stop rule halting the proof process on the current branch of the proof tree, when the ground instances of the current term can be stated as $C$-reducible. This happens when the whole current term $u$ can be abstracted, i.e. when the induction hypothesis applies on it, or when $u \in T(C, N)$.

As said before, ordering constraints have to be satisfied by a noetherian ordering having the constructor subterm property. Any simplification ordering holds. We can make further assumptions on the ordering to enable constraints to be satisfied. In particular, for a Recursive Path Ordering or Lexicographic Path Ordering whose precedence $\succ$ is such that $f \succ c \forall f \in D, c \in C$, we have $t \succ u$ for $t \in T(F, X)$ containing at least a symbol of $D$ and $u \in T(C)$.

Exploiting equalities in $A$ can also help to solve ordering constraints: $t_i \downarrow \leq c = X_i$ means that $\theta X_i$ is a $C$-form obtained by rewriting $\theta t_i$, for every ground substitution $\theta$. As we do not abstract terms of $T(C, N)$, if $t_i \neq x_i \in X$, $\theta t_i$ contains at least a symbol of $D$, and then $\theta t_i \succ \theta X_i$ for a RPO or an
LPO defined as above. As \( t_{\text{ref}} \) also contains a symbol of \( D \), we also have \( \Theta t_{\text{ref}} \supset \Theta X \) for \( X \in \mathcal{N} \).

Before giving the inference rules, let us note that the inductive reasoning can be completed in the following way. When the induction hypothesis cannot be applied to a term \( u \), it may be possible to prove \( C \)-reducibility of every ground instance of \( u \) by another way. Let \( C \rightarrow \text{RED}(u) \) be a predicate that is true if every ground instance of \( u \) is \( C \)-reducible. In the first and third previous steps of the inductive reasoning, we then associate the alternative predicate \( C \rightarrow \text{RED}(u) \) to the condition \( t > u \). It is true in particular if \( u \in T(C, \mathcal{N}) \), as said in the presentation of the \textbf{Stop} rule.

The rules are given in Table 1. They use a reference term \( t_{\text{ref}} = g(x_1, \ldots, x_m) \), where \( x_1, \ldots, x_m \in \mathcal{X} \) and \( g \in D \) (if \( g \) is a constant, then \( t_{\text{ref}} = g \)). To prove \( C \)-reducibility for every term \( t \) in \( T(F) \) we proceed as follows: for each defined symbol \( g \in D \), we apply the rules using the reference term \( t_{\text{ref}} = g(x_1, \ldots, x_m) \) on the initial set of 3-tuples \( \{ (t_{\text{ref}} = g(x_1, \ldots, x_m), T, \top) \} \), with a specific strategy \( S \).

The strategy \( S \) repeats the following steps: first, apply \textbf{Abstract}, and then try \textbf{Stop}. Then try all possible applications of \textbf{Narrow}. Then, try \textbf{Stop} again.

Note that if \( A \) is satisfiable, the transformed forms of \( A \) by \textbf{Abstract} and \textbf{Stop} are also satisfiable. Moreover, the first application of \textbf{Abstract} generates \( A = (\mathcal{A}, x_1 \ldots, c = X_1) \), always satisfied by the constructor constant supposed to exist in \( C \). Thus with the strategy \( S \), it is useless to prove satisfiability of \( A \) in the rules \textbf{Abstract} and \textbf{Stop}.

The process may not terminate if there is an infinite number of applications of \textbf{Abstract} and \textbf{Narrow} on the same branch of a derivation tree. Nothing can be said in that case. It stops if no inference rule applies anymore.

A branch of the derivation tree is said to be successful if it is ended by an application of \textbf{Stop}, i.e. if its final state is of the form \( (\emptyset, A, C) \).

Thus, the inductive \( C \)-reducibility proof is successful if there is at least one successful branch in the proof tree, corresponding to each possible ground term. Let us develop this point.

In fact, branching, produced by \textbf{Narrow}, can generate different states with narrowing substitutions \( \sigma_1, \ldots, \sigma_n \). These substitutions can be compared. For \( \sigma_i \) and \( \sigma_j \), three situations may occur: \( \sigma_i \) is strictly less general than \( \sigma_j \), which is noted \( \sigma_i \succ \sigma_j \), (or \( \sigma_i \) is strictly less general than \( \sigma_j \), \( \sigma_i \) and \( \sigma_j \) are equal up to a renaming, or else \( \sigma_i \) and \( \sigma_j \) are incomparable.

States corresponding to substitutions that are more general than other ones then represent a set of ground instances that contains the other ones. So, for proving \( C \)-reducibility for all ground instances at a branching point, it is sufficient to prove \( C \)-reducibility only for the “most general states”.

A branching node in a proof tree can only be a state, on which the Narrow rule applies. Let \( \Sigma \) be the set of narrowing substitutions (possibly with different rewrite rules) at a given branching node. Let \( \Sigma_0 \) be the reduced set from \( \Sigma \) such that \( \sigma \in \Sigma_0 \) iff \( \sigma \in \Sigma \) and \( \beta \sigma \in \Sigma \) such that \( \sigma \succ \sigma' \) on \( (\mathcal{Dom}(\sigma) \setminus \mathcal{Var}(l)) \cup (\mathcal{Dom}(\sigma') \setminus \mathcal{Var}(l')) \), where \( l \) and \( l' \) are the left-hand sides of rules respectively used to produce the narrowing substitutions \( \sigma \) and \( \sigma' \). The set \( \Sigma_0 \) may yet contain equivalent (equal up to a renaming) substitutions which are marked as such. So for any two substitutions in \( \Sigma_0 \), either they are equivalent, or they are incomparable.

A proof tree is \textit{weakenly successful} if at each branching node:

- for each class of equivalent substitutions, there exists at least one weakly successful subtree corresponding to a substitution in this class,
- all subtrees corresponding to incomparable substitutions are weakly successful,
- a tree reduced to the state \((\emptyset, A, C)\) is weakly successful.

So the strategy \( S \) can be optimized as follows: at each branching point of a proof tree, with set of substitutions \( \Sigma \), we only develop the subtrees corresponding to \( \Sigma_0 \). Moreover, given two subtrees corresponding to equivalent substitutions, as soon as one of them is weakly successful, the other one is cut.

We write \textit{SUCCESS}(\( g, \succ \)) if the proof tree obtained by application on \( \{ (g(x_1, \ldots, x_m), T, \top) \} \), with the strategy \( S \), of the inference rules whose conditions are satisfied by an ordering \( \succ \), is weakly successful.

**Theorem 2.** Let \( R \) be a RS on a set \( F \) of symbols. If there exists a noetherian ordering \( \succ \) having the constructor subterm property, such that for each defined symbol \( g \), we have \textit{SUCCESS}(\( g, \succ \)), then \( R \) is \( C \)-reducing.

Recall the important point that the ordering \( \succ \) has to be the same for all \( g(x_1, \ldots, x_m) \in D \).

The subtree cut process is similar to the one defined for weak termination, for which a formal description with a complete set of inference rules is given in [16].

Let us come back to the first example of the introduction.

**Example:** We prove that the RS

- \( and(1, x) \rightarrow x \)
- \( and(0, x) \rightarrow 0 \)
- \( or(1, x) \rightarrow 1 \)
- \( or(0, x) \rightarrow x \)
- \( and(1, x) \rightarrow not(not(and(1, x))) \)
- \( not(1) \rightarrow 0 \)
- \( not(0) \rightarrow 1 \)
- \( not(and(x, y)) \rightarrow or(not(x), not(y)) \)

on \( T(F) \), with \( F = \{ and : 2, or : 2, not : 1, 1 : 0, 0 : 0 \} \), and \( C = \{ 0, 1 \} \), is \( C \)-reducing.

Applying the rules on \( not(x) \), we get:

- \( not(x_1) \)
  - \( A = T, C = T \)

**Abstract**

- \( not(X_1) \)
  - \( A = (x_1, c = X_1) \)
  - \( C = (not(x_1) > x_1) \)

**Narrow**

\( \begin{align*}
1 & \sigma_1 = (X_1 = 0) \\
A & = (x_1, c = 0) \\
C & = (not(x_1) > x_1)
\end{align*} \)

\( \begin{align*}
0 & \sigma_2 = (X_1 = 1) \\
A & = (x_1, c = 1) \\
C & = (not(x_1) > x_1)
\end{align*} \)

**Stop(twice)**

\( \emptyset \)

- \( A = (x_1, c = 0) \)
  - \( C = (not(x_1) > x_1) \)

\( \emptyset \)

- \( A = (x_1, c = 1) \)
  - \( C = (not(x_1) > x_1) \)

**Narrow** would also apply with \( \sigma_3 = (X_1 = and(X_2, X_3)) \), but \( \sigma_3 \) is not licit since \( and(X_2, X_3) \notin T(C, N) \).
Moreover $\Sigma = \{\sigma_1, \sigma_2, \sigma_3\}$ is not($X_1$)-covering since for every possible ground instance $\theta X_1$ of $X_1$ (0 or 1), there exists $\sigma$ in $\Sigma$ such that $\theta$ not($X_1$) $= \mu \sigma$ not($X_1$) for some $\mu$, which is here equal to $Id$.

Stop applies (twice) since 0 and 1 are terms of $T(C, N)$.

The ordering constraints are satisfied by any term ordering having the subterm property, for example by a RPO with any precedence.

Now, considering or($x_1, x_2$), we get:

or($x_1, x_2$) $A = \top$ $C = \top$

Abstract

or($X_1, X_2$) $A = (x_1 \downarrow C = X_1 \land x_2 \downarrow C = X_2)$
$C = (or(x_1, x_2) > x_1, x_2)$

Narrow

$X_2$

$\sigma_1 = (X_1 = 0)$
$A = (x_1 \downarrow C = 0 \land x_2 \downarrow C = X_2)$
$C = (or(X_1, X_2) > x_1, x_2)$

Stop(twice)

$\emptyset$

$A = (x_1 \downarrow C = 0 \land x_2 \downarrow C = X_2)$
$C = (or(X_1, X_2) > x_1, x_2)$

$\emptyset$

$A = (x_1 \downarrow C = 1 \land x_2 \downarrow C = X_2)$
$C = (or(X_1, X_2) > x_1, x_2)$

$\Sigma = \{\sigma_1, \sigma_2\}$ is or($X_1, X_2$)-covering.

Stop applies since $X_2$ and 1 are terms of $T(C, N)$.

Finally, considering and($x_1, x_2$), we get:

and($x_1, x_2$) $A = \top$ $C = \top$

Abstract

and($X_1, X_2$) $A = (x_1 \downarrow C = X_1 \land x_2 \downarrow C = X_2)$
$C = (and(x_1, x_2) > x_1, x_2)$

Narrow

0

$\sigma_1 = (X_1 = 0)$
$A = (x_1 \downarrow C = 0 \land x_2 \downarrow C = X_2)$
$C = (and(X_1, X_2) > x_1, x_2)$

$X_2$

$\sigma_2 = (X_1 = 1)$
$A = (x_1 \downarrow C = 1 \land x_2 \downarrow C = X_2)$
$C = (and(X_1, X_2) > x_1, x_2)$

not(not(and($X_1, X_2$)))

$\sigma_3 = (X_1 = 1)$
$A = (x_1 \downarrow C = 1 \land x_2 \downarrow C = X_2)$
$C = (and(X_1, X_2) > x_1, x_2)$

Stop(three times)

0

$A = (x_1 \downarrow C = 0 \land x_2 \downarrow C = X_2)$
$C = (and(X_1, X_2) > x_1, x_2)$

0

$A = (x_1 \downarrow C = 1 \land x_2 \downarrow C = X_2)$
$C = (and(X_1, X_2) > x_1, x_2)$

0

$A = (x_1 \downarrow C = 1 \land x_2 \downarrow C = X_2)$
$C = (and(X_1, X_2) > x_1, x_2)$

$\Sigma = \{\sigma_1, \sigma_2, \sigma_3\}$ is and($X_1, X_2$)-covering.

Stop applies on the two first branches because 0 and $X_2$ are terms of $T(C, N)$, and on the third one because it is issued from the same narrowing substitution as the second one, on which Stop applies.

The ordering constraints of the last two proof trees are satisfied by the same ordering as previously.

Note that, as $R$ does not contain rules whose left-hand side is a $C$-term, $C$-reducibility ensures weak termination of the RS.
Now, if we add to \( R \) the relations between constructors:

\[
\begin{align*}
1 & \rightarrow \text{true} \\
\text{true} & \rightarrow 1 \\
0 & \rightarrow \text{false} \\
\text{false} & \rightarrow 0
\end{align*}
\]

where \text{true} and \text{false} are also constructor symbols, we lose the weak termination property. But \( C \)-reducibility is proved in the same way than previously, since we have no additional proof tree, and the new proof trees are the same as the previous ones.

Let us now consider the second example of the introduction.

Example:

\[
\begin{align*}
\text{quot}(0, s(y), s(z)) & \rightarrow 0 \\
\text{quot}(s(x), s(y), z) & \rightarrow \text{quot}(x, y, z) \\
\text{quot}(x, 0, s(z)) & \rightarrow \text{quot}(x, s(z), s(z)) \\
\text{quot}(x, y, 0) & \rightarrow \text{error} \\
\text{quot}(\text{error}, y, z) & \rightarrow \text{error} \\
\text{quot}(x, \text{error}, z) & \rightarrow \text{error} \\
\text{quot}(x, y, \text{error}) & \rightarrow \text{error}
\end{align*}
\]

with \( F = \{ \text{quot} : 3, s : 1, 0 : 0, \text{error} : 0 \} \) and \( C = \{ s, 0, \text{error} \} \).

So, either we prove innermost termination of the RS, (the dependency pair method [18] works), and we infer sufficient completeness, by Proposition 2 and Theorem 1, in proving coveredness of all patterns \( f(X_1, \ldots, X_n), f \in D \), or we directly apply the \( C \)-reducibility proof method. Let us develop the second solution.

Applying the rules on \( \text{quot}(x_1, x_2, x_3) \), we get:

\[
\begin{align*}
\text{quot}(x_1, x_2, x_3) & \quad A = \top, \quad C = \top \\
\text{Abstract} & \quad \text{quot}(X_1, X_2, X_3) = (x_1 \parallel_c = X_1 \land x_2 \parallel_c = X_2 \\
& \quad \land x_3 \parallel_c = X_3) \\
& \quad C = (\text{quot}(x_1, x_2, x_3) > x_1, x_2, x_3) \\
\text{Narrow} & \quad \sigma_1 = (X_1 = 0 \land X_2 = s(X'_2) \\
& \quad \land X_3 = s(X'_3)) \\
& \quad A = (x_1 \parallel_c = 0 \land x_2 \parallel_c = s(X'_2) \\
& \quad \land x_3 \parallel_c = s(X'_3)) \\
& \quad C = (\text{quot}(x_1, x_2, x_3) > x_1, x_2, x_3) \\
\text{quot}(X'_1, X'_2, X_3) & \quad \sigma_2 = (X_1 = s(X'_1) \land X_2 = s(X'_2)) \\
& \quad A = (x_1 \parallel_c = s(X'_1) \\
& \quad \land x_2 \parallel_c = s(X'_2) \land x_3 \parallel_c = X_3) \\
& \quad C = (\text{quot}(x_1, x_2, x_3) > x_1, x_2, x_3)
\end{align*}
\]

Applying the rule \text{Narrow} here gives seven branches, following the seven rules of \( R \). Let \( u = \text{quot}(X_1, X_2, X_3) \).

The set \( \Sigma = \{ \sigma_i, i \in [1..7] \} \) is \( \alpha \)-covering, since every \( \sigma_i \) of \( \Sigma \), where \( P = \{ 0, \text{error}, s(X) | X \in N \} \), has a generalization in \( \Sigma \).

Then \text{Stop} applies on all branches, except the third one, for the following reasons. On the second branch and the last four ones, we get \( C \)-terms as current terms.

On the second branch, we have \( \text{quot}(x_1, x_2, x_3) \Rightarrow \text{quot}(X'_1, X'_2, X_3) \) for a Lexicographic Path Ordering with any precedence and a left-to-right status for \( \text{quot} \). Indeed, \( R \) does not contain any rule whose left-hand side is a \( C \)-term. Then, if \( x_1 \) represents a \( C \)-term, it is in normal form, and then \( x_1 = x_1 \parallel_c \). In \( A \), we have \( x_1 \parallel_c = s(X'_1) \). Then \( x_1 > X'_1 \).

If \( x_1 \) represents a term containing a defined symbol, then \( x_1 > X'_1 \) (see Section 5). In a similar way, we obtain \( x_2 > X'_2 \).

Now, if \( x_3 \) represents a term containing a defined symbol, as previously for \( x_1 \), we have \( x_3 > X'_3 \). If \( x_3 \) represents a \( C \)-term, we get \( x_3 = x_3 \parallel_c = X_3 \) so \( x_3 > X_3 \).

By definition of the LPO, as we have \( x_1 > X'_1 \), we have to verify that \( \text{quot}(x_1, x_2, x_3) > X'_2, X_3 \). This is true since \( x_2 > X'_2 \) and \( x_3 > X_3 \).

Then we get:

\[
\begin{align*}
\text{Stop (six times)} & \quad A = (x_1 \parallel_c = 0 \land x_2 \parallel_c = s(X'_2) \\
& \quad \land x_1 \parallel_c = s(X'_1)) \\
& \quad C = (\text{quot}(x_1, x_2, x_3) > x_1, x_2, x_3)
\end{align*}
\]
6. Finding a good derivation chain

Our proof process, as it simulates the rewriting mechanism, gives complete information on a rewriting branch leading to a C-form. It allows us extracting the exact application of rewrite rules that yields a C-form. To rewrite a term, it is enough to follow the rewriting scheme simulated by abstraction and narrowing in the proof trees.

We now formalize the use of the proof trees to compute a C-form for any term.

DEFINITION 12. Let \( R \) be a RS proved C-reducing with Theorem 2. The strategy tree \( S T_f \) associated to \( f \in D \) is the proof tree obtained from the initial state \( \{\{f(x_1,\ldots,x_m)\}\}, \top, \top \).

DEFINITION 13. Let \( R \) be a RS proved C-reducing with Theorem 2. Let \( S T = \{S T_f | f \in D \} \) be the set of strategy trees of \( R \) and \( s = f(s_1,\ldots,s_m) \in T(F) \). Rewriting \( s \) with respect to \( S T \) into a C-form, \( C form_{S T}(s) \), is defined in the following way, where \( \leftarrow \) denotes a transformation step:

- if \( f \in C \), then \( C form_{S T}(f(s_1,\ldots,s_m)) = f(C form_{S T}(s_1),\ldots,C form_{S T}(s_m)) \).
- if \( f \in D \), then reducing \( s \) with respect to \( S T \) into \( C form_{S T}(s) \) is performed by following the steps in the strategy tree \( S T_f \) of \( f \), where \( t \) is any term of the transformation chain of \( s \) with respect to \( S T \) and \( u \) is the corresponding term in \( S T_f \):
  - if the step is Abstract, and abstracts \( u \) at positions \( i_1,\ldots,i_p \), then \( t \leftarrow t[i'_j]_{j \in \{i_1,\ldots,i_p\}} \),
    where \( i'_j = \{ i'_j \}_{j \in \{i_1,\ldots,i_p\}} \).
  - if the step is Narrow with \( u \leftarrow_{\beta,\sigma} u' \), then \( t \leftarrow t' \) where \( t' \) is defined by \( t \leftarrow_{\beta,\sigma} t' \) with \( \alpha = \beta\sigma \) on \( Var(u) \cup Dom(u) \) and \( t' = u' \),
  - if the step is Stop, then \( t \leftarrow t' \),
    where \( t' = \{ t[i'_j] \}_{j \in \{i_1,\ldots,i_p\}} \).

Given a RS \( R \), the previous definition assumes that if the predicate \( C\rightarrow RED \) has been used to prove C-reducibility of a particular term \( u \) during the proof that \( R \) is C-reducing, one is able to find a C-form for ground instances of \( u \). It is obviously the case if \( u \in T(C,N) \). Under this assumption, the following theorem holds.

THEOREM 3. Let \( R \) be a RS proved C-reducing with Theorem 2 and \( ST \) its set of strategy trees. Then for any term \( t \in T(F) \), \( C form_{S T}(t) \) is a C-form of \( t \) for \( R \).

7. Some additional examples

Example:
The following RS is neither terminating, nor even innermost terminating:

\[
\begin{align*}
& f(x) \rightarrow if(x, c, f(true)) \\
& if(true, x, y) \rightarrow x \\
& if(false, x, y) \rightarrow y \\
& if(c, x, y) \rightarrow c
\end{align*}
\]
with $\mathcal{F} = \{ \text{if} : 3, s : 1, f : 1, c : 0, \text{true} : 0, \text{false} : 0 \}$ and $\mathcal{C} = \{c, \text{true}, \text{false}\}$. Applying the rules on $f(x_1)$, we get:

$$f(x_1) \quad A = T, \quad C = T$$

Abstract

$$f(X_1) \quad A = (x_1|_c = X_1) \quad C = (f(x_1) > x_1)$$

Narrow

$$\text{if}(X_1, c, f(\text{true})) \quad \sigma = \text{Id} \quad A = (x_1|_c = X_1) \quad C = (f(x_1) > x_1)$$

Then, the last branch, whose narrowing substitution is equal to those of the fourth one, can be cut.

Applying the rules on $\text{if}(x_1, x_2, x_3)$, we get:

$$\text{if}(x_1, x_2, x_3) \quad A = T, \quad C = T$$

Abstract

$$\text{if}(X_1, X_2, X_3) \quad A = (x_1|_c = X_1, x_2|_c = X_2, x_3|_c = X_3) \quad C = (f(x_1, x_2, x_3) > x_1, x_2, x_3)$$

Narrow

$$X_2 \quad \sigma_1 = (X_1 = \text{true}) \quad A = (x_1|_c = \text{true}, x_2|_c = X_2, x_3|_c = X_3) \quad C = (f(x_1, x_2, x_3) > x_1, x_2, x_3)$$

$$X_3 \quad \sigma_1 = (X_1 = \text{false}) \quad A = (x_1|_c = \text{false}, x_2|_c = X_2, x_3|_c = X_3) \quad C = (f(x_1, x_2, x_3) > x_1, x_2, x_3)$$

Then, $\text{Stop}$ applies on the three branches, which ends the proof.

Example: The following RS is perhaps terminating, but it cannot be proved with a path ordering:

$$\emptyset \quad A = (x_1|_c = \text{true}, x_2|_c = X_2, x_3|_c = X_3) \quad C = (f(x_1, x_2, x_3) > x_1, x_2, x_3)$$

The first Narrow is covering for $f(X_1)$ and the second, for $\text{if}(X_1, c, f(\text{true}))$.

As the three first branches above are issued from narrowing substitutions, which are less general that the substitution of the fourth branch, they can be cut. We then only keep the last state, on which Narrow (covering for $\text{if}(X_1, c, f(\text{true}), c, f(\text{true}))$) is applied again, which gives the following states.

$$c \quad \sigma_1 = (X_1 = \text{true}) \quad A = (x_1|_c = \text{true}) \quad C = (f(x_1) > x_1)$$

$$\text{if}(\text{true}, c, f(\text{true})) \quad \sigma_2 = (X_1 = \text{false}) \quad A = (x_1|_c = \text{false}) \quad C = (f(x_1) > x_1)$$

$$c \quad \sigma_3 = (X_1 = c) \quad A = (x_1|_c = c) \quad C = (f(x_1) > x_1)$$

$$\text{if}(X_1, c, c) \quad \sigma_4 = \text{Id} \quad A = (x_1|_c = X_1) \quad C = (f(x_1) > x_1)$$

$$\text{if}(X_1, c, \text{if}(\text{true}, c, f(\text{true}))) \quad \sigma_5 = \text{Id} \quad A = (x_1|_c = X_1) \quad C = (f(x_1) > x_1)$$

Again, we can cut the three first branches. Stop applies on the fourth state since we have $f(x_1) \triangleright \text{if}(X_1, c, c)$ for a RPO with the precedence $f \triangleright \text{if}, c$. Indeed, we have $f(x_1) \triangleright c$, and $f(x_1) \triangleright X_1$ following the remarks of Section 5.
Narrow is covering for \( \text{length}(X_1) \).

Now, Stop applies on all states: on the first two ones since terms are \( C \)-terms, on the last two ones since \( \text{length}(x_1) \vdash s(\text{length}(X_2)) \) and \( \text{length}(x_1) \vdash s(\text{length}(X_1)) \) for a LPO with the precedence \( \text{length} \prec \text{s} \) and \( \text{length} \sim \text{length} \). More precisely, \( \text{length}(x_1) \vdash s(\text{length}(X_1)) \) since, in \( A \), we have \( x_1 \downarrow_c = s(X_2) \), and then following the same reasoning than in the example with \( \text{quot} \), \( x_1 \succ X_2 \).

Similarly, \( \text{length}(x_1) \vdash s(\text{length}(X_4)) \) since \( x_1 \succ X_4 \).

Applying the rules on \( \text{length}(x_1) \), we get:

\[
\begin{align*}
\text{length}(x_1) & \quad A = \top, \quad C = \top \\
\text{Abstract} & \quad A = (x_1 \downarrow C = X_1) \quad C = (\text{length}(x_1) \succ x_1) \\
\text{Narrow} & \quad \sigma = \text{ld} \quad A = (x_1 \downarrow C = X_1) \quad C = (\text{length}(x_1) \succ x_1) \\
\text{Narrow} & \quad \sigma = X_1 = \text{nil} \quad A = (x_1 \downarrow C = \text{nil}) \quad C = (\text{length}(x_1) \succ x_1) \\
\text{s(0)} & \quad \sigma = (X_1 = 0) \quad A = (x_1 \downarrow C = 0) \quad C = (\text{length}(x_1) \succ x_1)
\end{align*}
\]

\[
\begin{align*}
\text{s(length}(X_2)) & \quad \sigma = (X_1 = s(X_2)) \quad A = (x_1 \downarrow C = s(X_2)) \quad C = (\text{length}(x_1) \succ x_1) \\
\text{s(length}(X_4)) & \quad \sigma = (X_1 = \text{cons}(X_3, X_4)) \quad A = (x_1 \downarrow C = \text{cons}(X_3, X_4)) \quad C = (\text{length}(x_1) \succ x_1)
\end{align*}
\]

The first narrowing step is covering for \( \text{length}(X_1) \), and the second one for \( \text{length}(X_3) \).

Now, Stop applies four times, like in the previous proof tree, but with \( \text{length}(x_1) \) as reference term instead of \( \text{length}(x_1) \). With the additional precedence \( \text{length} \succ \text{s} \), we get \( \text{length}(x_1) \vdash s(\text{length}(X_2)) \) and \( \text{length}(x_1) \vdash s(\text{length}(X_4)) \).

Example:

\[
\begin{align*}
\text{incr}(\text{nil}) & \rightarrow \text{nil} \\
\text{incr}(\text{cons}(x, \text{l})) & \rightarrow \text{cons}(s(x), \text{incr}(\text{l})) \\
\text{adx}(\text{nil}) & \rightarrow \text{nil} \\
\text{adx}(\text{cons}(x, \text{l})) & \rightarrow \text{incr}(\text{cons}(x, \text{adx}(\text{l}))) \\
\text{nats} & \rightarrow \text{adx}(\text{zeros}) \\
\text{zeros} & \rightarrow \text{cons}(0, \text{zeros}) \\
\text{head}(\text{cons}(x, \text{l})) & \rightarrow x \\
\text{tail}(\text{cons}(x, \text{l})) & \rightarrow \text{l} \\
\text{adx}(0) & \rightarrow \text{err} \\
\text{adx}(s(x)) & \rightarrow \text{err} \\
\text{incr}(0) & \rightarrow \text{err} \\
\text{incr}(s(x)) & \rightarrow \text{err} \\
\text{tail}(\text{nil}) & \rightarrow \text{nil} \\
\text{tail}(0) & \rightarrow \text{err} \\
\text{tail}(s(x)) & \rightarrow \text{err} \\
\text{head}(\text{nil}) & \rightarrow \text{nil} \\
\text{head}(0) & \rightarrow \text{err} \\
\text{head}(s(x)) & \rightarrow \text{err}
\end{align*}
\]

with \( \mathcal{F} = \{ \text{cons} : 2, \text{incr} : 1, \text{adx} : 1, \text{head} : 1, \text{tail} : 1, s : 1, \text{nats} : 0, \text{zeros} : 0 : 0, \text{nil} : 0, \text{err} : 0 \} \) and \( \mathcal{C} = \{ 0, s, \text{cons}, \text{nil}, \text{err} \} \). Applying the rules on \( \text{incr}(x_1) \), we get:
8. Conclusion and perspectives

In this paper, we have studied the problem of C-reducibility of ground terms, i.e., reducibility of terms into a constructor form on at least one of their rewriting derivations. This property is interesting from several points of view. It allows us to warrant a completely computed form for every data of rewrite programs, and directly implies sufficient completeness of specifications. In the light of C-reducibility, termination of a program becomes less important, provided one is able to reach a C-form, i.e. to compute a result, which gives a criterion to stop the computations.

Our proof method establishes C-reducibility, and then sufficient completeness, by explicit induction on the C-reducibility property itself. It simulates rewriting by alternatively using abstraction of subterms, and narrowing, provided all possible ground instances of the narrowed terms are covered. It works on the ground term algebra using as induction relation a noetherian ordering having the constructor subterm property.

It does not require confluence, nor restrictions like absence of relation between constructors or the constructor preserving property. It does not need any termination property either. When there is no constructor left-hand side of rule, it even provides a proof of weak termination. It is constructive in the sense that a computation branch leading to a constructor form can be deduced from the proof trees.

Thanks to the power of induction, the ordering constraints generated by the proof process are often simple and satisfied by the subterm ordering, or by an usual ordering like the Recursive Path Ordering. If not, they can be delegated to automatic ordering constraint solvers. Verifying unsatisfiability of A is also automatable, with the given sufficient conditions cited in Section 4.3. So we can have an automatic implementation of the proof procedure establishing C-reducibility. As the proof technique is close to the rewriting process itself, it seems to be promising to extend it to conditional, typed or constrained rewriting.

References


[8] H. Comon and F. Jacquemard. Ground reducibility is...


This appendix contains the proof of Proposition 3 and of the theorems.

**Proposition 3.** Let $\mathcal{R}$ be a RS on $T(\mathcal{F}, \mathcal{X})$, $X_1, \ldots, X_m \in \mathcal{N}$, $x_1, \ldots, x_m \in \mathcal{X}$. Then $\text{quot}(x_1, \ldots, X_m)$ is ground reducible for every $f \in D$, if $\text{quot}(x_1, \ldots, x_m)$ is ground reducible for every $f \in D$.

**Proof.** Let $\text{quot}(x_1, \ldots, X_m)$, $f \in D$, and $t = f(t_1, \ldots, t_m)$ a ground instance of $\text{quot}(x_1, \ldots, X_m)$. The proof is made by structural induction on $t$.

- Either $t$ is a constant $a$, and is ground reducible by hypothesis,
- or $t$ is not a constant, and:
  - either there exists a subterm $t_i$ of $t$ of the form $g(u_1, \ldots, u_p)$ with $g \in D$. By induction hypothesis, $t_i$ is reducible, and then $t$ is;
  - or every subterm $t_i$ of $t$ of the form $g(u_1, \ldots, u_p)$, with $g \in C$, and we conclude thanks to the hypothesis.

The converse property is obvious.

**Proof of Theorem 2**

**Theorem 2.** Let $\mathcal{R}$ be a RS on a set $\mathcal{F}$ of symbols. If there exists a noetherian ordering $\succ$ having the constructor subterm property, such that for each defined symbol $g$, we have $\text{SUCCESS}(g, \succ)$, then $\mathcal{R}$ is $C$-reducible.

**Proof.**

We prove by induction on $T(\mathcal{F})$ that every ground instance $\theta \text{quot}(x_1, \ldots, x_m)$ of any term $\text{quot}(x_1, \ldots, x_m) \in T(\mathcal{F}, \mathcal{X})$ is $C$-reducible. The induction ordering is constrained along the proof. At the beginning, it has at least to be noetherian and to have the constructor subterm property. Such an ordering always exists on $T(\mathcal{F})$ (for instance the embedding relation). Let us denote it $\succ$.

If $f$ is a constructor, by constructor subterm property of $\succ$, we have $\theta \text{quot}(x_1, \ldots, x_m) = f(\theta x_1, \ldots, \theta x_m) \succ \theta x_1, \ldots, \theta x_m$. Then, by induction hypothesis, let us suppose that $\theta x_1, \ldots, \theta x_m$ are $C$-reducible. Let $\theta x_1 | C$, $\ldots, \theta x_m | C$ be respectively any of their $C$-forms. It remains to prove that $f(\theta x_1 | C, \ldots, \theta x_m | C)$ is $C$-reducible. This is the case since it is a $C$-term.

If $f$ is not a constructor, let us denote it $g$ and prove that $g(\theta x_1, \ldots, \theta x_m)$ is $C$-reducible for any $\theta$ satisfying $(A_0 = \top, C_0 = \top)$, if application of the inference rules with the strategy $S$ on $((g(x_1, \ldots, x_m)), \top, \top)$, terminates on a weakly successful proof tree. Let us denote $g(x_1, \ldots, x_m)$ by $t_{ref}$ in the sequel of the proof.

To each step of the procedure characterized by a state $s = (\{t\}, A, C)$, we associate the set of ground terms $G = \{ \alpha t \mid \alpha \text{ satisfies } A \}$, that is the set of ground instances represented by $s$. Inference rules of Abstract and Narrow transform $\{t\}, A, C$ into $\{t', A', C'\}$ to which is associated $G' = \{ \beta' t \mid \beta \text{ satisfies } A' \}$. We then prove the following result: for each $\alpha$ such that $\alpha t$ is in $G$, there exists a $\beta$ such that $\beta'$ is in $G'$ and such that $C$-reducibility of $\beta'$ implies $C$-reducibility of $\alpha t$.

- Either Abstract is applied, so the current term $t$ becomes $t' = t[X_j]_{j \in \{1, \ldots, p\}}$, where $i_1, \ldots, i_p$ are the abstraction positions.

According to the condition of application of Abstract at positions $j \in \{i_1, \ldots, i_p\}$, each term $t_j$ is such that:

- either $C - \text{RED}(t_j)$ is true, and then by definition of the predicate $C - \text{RED}$, $\alpha t_j$ is $C$-reducible;
- or $C \land t_{ref} > t_j$ is satisfiable and then, by induction hypothesis, $\alpha t_j$ is $C$-reducible.

Then let us define $\beta = \alpha \cup \bigcup_{j \in \{1, \ldots, p\}} X_j = \alpha t_j | C$.

Clearly $\beta$ satisfies $A'$ provided the $X_j$ are not in $\text{Dom}(\alpha)$, which is the case since they are fresh variables, and $\beta' = \alpha(\alpha t_j | C)_{j \in \{1, \ldots, p\}}$, so $\alpha t \rightarrow^* \beta'$. Therefore $C$-reducibility of $\beta'$, for any possible $C$-form $\alpha t|_{1, C}$ of $\alpha t|_{1, C}$, implies $C$-reducibility of $\alpha t$.

- Or Narrow is applied on $\{x = f(u_1, \ldots, u_m)\}, A, C$.

Clearly $\alpha t$ is in $\mathcal{C}$-reducible and to have the constructor subterm property.

Accordingly, $\text{SUCCESS}(g, \succ)$, then $\mathcal{R}$ is $C$-reducible.

Therefore, $S$-termination is ensured for all terms in all sets $\mathcal{G}$ of the proof tree.

As the process is initialized with $\{t_{ref}\}$ and a constraint problem satisfiable by any ground substitution, we get that $g(\theta x_1, \ldots, \theta x_m)$ is $C$-reducible, for any $t_{ref} = g(x_1, \ldots, x_m)$, and any ground instance $\theta$.

**Proof of Theorem 3**

We proceed by crossed proof with the following lemma.

**Lemma 1.** Let $s = f(s_1, \ldots, s_m) \in T(\mathcal{F})$ and $t \rightarrow^* t'$ be any step of a transformation chain of $s$ with respect to $ST$ until $C \text{-form}_ST(s)$. If this step is determined relatively to a step $u \rightarrow^* u'$ of the strategy tree $ST_f$, then $t$ is a ground instance of $u$.
Proof. Let \( s = f(s_1, \ldots, s_m) \in T(F) \) be a term and \( t \mapsto t' \) any step of a transformation chain of \( s \) with respect to \( ST \) until \( \text{Cform}_{ST}(s) \). The proof is made by induction on the length \( l \) of the transformation chain of \( s \) with respect to \( ST \).

Let \( l = 1 \).

- if \( f \in C \), by Definition 13, there is nothing to prove since \( ST_f \) is not used for normalizing the whole term \( s \).
- if \( f = g \in D \), then by definition of the strategy \( S \), the first step of each proof tree of \( ST \) is Abstract if \( f \) is not a constant, or Narrow if \( f \) is a constant. Indeed, as the narrowing step is \( a \)-covering, \( a \) narrows. As \( l = 1 \), this should also be the last step of the branch in the proof tree, which is impossible, since by definition of \( S \), the last step is Stop.

Let \( l = n \geq 1 \).

Let \( s' = g(s'_1, \ldots, s'_n) \) be the term obtained from \( s \) with a transformation chain of length \( n \) with respect to \( ST \). Let the property be true for \( l = n \), i.e. for every step \( t \mapsto t' \) of the transformation chain of \( s \) with respect to \( ST \) until \( s' \), determined relatively to a step \( u \mapsto u' \) of the strategy tree \( ST_f \), then \( t \) is a ground instance of \( u \).

Let \( s' \mapsto s'' \) be the \( n+1 \)-th step of the given transformation chain of \( s \) with respect to \( ST \), determined relatively to a step \( u' \mapsto u'' \) of the strategy tree \( ST_f \). We prove that \( u'' \) is a ground instance of \( u' \).

Let \( s'' \mapsto s''' \) be the \( n \)-th step of the given transformation chain of \( s \) with respect to \( ST \), determined relatively to a step \( u'' \mapsto u''' \) of the strategy tree \( ST_f \). By induction hypothesis, \( s''' \) is a ground instance of \( u''' \), so there exists a ground substitution \( \theta \) such that \( \theta u''' = s''' \).

- if \( u'' \mapsto u''' \) is a step Abstract, then \( u''' \) is of the form \( u'[(x_1)_{i_1}, \ldots, (x_p)_{i_p}] \) where \( i_1, \ldots, i_p \) are the abstraction positions of \( u'' \) and the \( x_j \) are new abstraction variables.

By Definition 13, \( s'' \mapsto s''[(x'_j)_{i_1}, \ldots, (x'_i)_{i_p}] \) where the \( x'_j \) are \( s''[x_j]_{i_c} \) if we have \( C-\text{RED}(u'[(x_j)_{i_c}]) \), and \( \text{Cform}_{ST}(s''[(x_j)_{i_c}]) \) otherwise:

- In the first case, \( s''[x_j]_{i_c} \) is a ground instance of \( X_j \); let us note that \( s''[x_j]_{i_c} \) exists since by induction hypothesis \( s''[x_j]_{i_c} \) is a ground instance of \( u''[x_j] \) and we have \( C-\text{RED}(u''[x_j]) \);

- In the second case, by Theorem 3 on each \( s''[x_j] \) strict subterm of \( s'' \), we have \( \text{Cform}_{ST}(s''[x_j]) \) in \( C \)-form, hence a ground instance of \( X_j \).

Therefore \( s'' \) is a ground instance of \( u'' \).

- if \( u'' \mapsto u''' \) is a step Narrow, by definition 13, we have \( s'' = \theta u'' \overset{p_1 \mapsto r}{\rightarrow} s' = \mu u', \) for some ground substitution \( \mu \), where the used branch of the step Narrow (always existing since the set of narrowing substitutions of Narrow is \( u'' \)-covering) is \( u'' \overset{p_1 \mapsto r, \sigma}{\rightarrow} u' \), with \( \theta = \mu \sigma \). So \( s'' \) is a ground instance of \( u''' \).

- if \( u'' \mapsto u''' \) is a step Stop, then \( u'' \) does not exist, since the state in \( ST_f \) after this step is a state of the form \( (\theta, A, C) \). Then there is nothing to prove.

Theorem 3. Let \( R \) be a RS proved C-reducing with Theorem 2 and ST its set of strategy trees. Then for any term \( t \in T(F) \), \( \text{Cform}_{ST}(t) \) is a C-form of \( t \) for \( R \).

Proof. Let \( s = f(s_1, \ldots, s_m) \in T(F) \) be a term to be reduced with respect to \( ST \). We prove if a transformation chain of \( s \) with respect to \( ST \) gives a term \( t \), \( t \) is also a reduced form of \( s \) for \( R \). We also show that the last term of any transformation chain of \( s \) with respect to \( ST \) is a C-term. This term eventually exists, since the process always stops, for \( ST \) is a set of successful trees, whose branches only have a finite number of states.

Let us show that \( t \) is a reduced form of \( s \) for \( R \), by induction on the length \( l \) of the transformation chain of \( s \) with respect to \( ST \). For that, we use an induction on \( T(F) \) and the property : \( \text{Cform}_{ST}(s) \) is a C-reduced form of \( s \) for \( R \), with an ordering \( \succ \) satisfying all constraints generated by \( S \) in proving that \( R \) is C-reducing. Such an ordering exists since, by hypothesis, \( R \) has been proved C-reducing using Theorem 2.

Let \( l = 1 \).

- if \( f \in C \), by Definition 13, \( \text{Cform}_{ST}(f(s_1, \ldots, s_n)) = f(\text{Cform}_{ST}(s_1), \ldots, \text{Cform}_{ST}(s_n)) \). With the second induction, we have \( s = f(s_1, \ldots, s_m) \succ s_1, \ldots, s_m \), and then, by induction hypothesis, \( \text{Cform}_{ST}(s_1), \ldots, \text{Cform}_{ST}(s_n) \) are respective C-reduced forms of \( s_1, \ldots, s_n \) for \( R \). Then \( f(\text{Cform}_{ST}(s_1), \ldots, \text{Cform}_{ST}(s_n)) \) is a reduced form of \( s \) for \( R \), and as \( f \) is a constructor symbol, is a C-term.

- if \( f = g \in D \), then by definition of the strategy \( S \), the first step of each proof tree of \( ST \) is Abstract, or Narrow if \( f \) is a constant. As \( l = 1 \), this should also be the last step of the branch in the proof tree, which is impossible since, by definition of \( S \), the last step of a branch is Stop.

Let \( l = n \geq 1 \).

Let \( t \) be the term obtained with a transformation chain of length \( n \) from \( s \) with respect to \( ST \). Let the property be true for \( l = n \), i.e. \( t \) is a reduced form of \( s \) with \( R \). We prove it for \( l = n + 1 \).

- if the current step in the proof tree of \( f \) is Abstract, by Definition 13, \( t \mapsto t' \overset{g}[\theta]_{i_c} \subseteq (t[\theta]_{i_c})_{i_1}, \ldots, (t[\theta]_{i_p})_{i_p} \), for some \( u \) such that \( t \) is a ground instance of \( u \), where \( i_1, \ldots, i_p \) are the abstraction positions and the \( t[\theta]_{i_c} \) are \( t_{i_c} \) if we have \( C-\text{RED}(u_{i_j}) \), and \( \text{Cform}_{ST}(t_{i_j}) \) otherwise.

For the \( j \) such that \( C-\text{RED}(u_{i_j}) \), the \( t_{i_j} \) are C-reduced forms of the \( t_{i_j} \), by definition, and then reduced forms of the \( t_{i_j} \).

For the \( j \) such that we do not have \( C-\text{RED}(u_{i_j}) \), we have \( u_{r, e f} \succ u_{i_j} \), where \( u_{r, e f} \) is the initial term of the proof tree, in which \( u \) appears. So \( t = \theta u_{r, e f} \succ t_{i_j} \), for some ground substitution \( \theta \), since by Lemma 1, \( t_{i_j} \) is a ground instance of \( u_{i_j} \). Then the induction hypothesis of the second induction applies to \( t_{i_j} \), so \( \text{Cform}_{ST}(t_{i_j}) \) is a C-reduced form of \( t_{i_j} \), and then a reduced form of \( t_{i_j} \).

Then the term \( t' \overset{t_{i_j}}{\rightarrow} = \theta[u_{r, e f} \succ t_{i_j}] \) is a reduced form of \( t \) for \( R \).

Since, by induction hypothesis (with the first induction), \( t \) is a reduced form of \( s \) for \( R \), we conclude that \( t' \) is a reduced form of \( s \) for \( R \).

- if the current step in the proof tree of \( f \) is Narrow, by Definition 13, \( t \mapsto t' \overset{\mu}{\rightarrow} t' = \mu u' \), for some ground substitution \( \mu \), where the used branch of the step Narrow (always existing since Narrow is such
that the set of narrowing substitutions is \(u\)-covering), is
\[ u \leadsto^{p,l \rightarrow r,\sigma} u', \]
with \(\theta = \mu\sigma\).

Then \(\mu t'\) is a reduced form of \(t\). Thus, reasoning as in the previous case, as \(t\) is a reduced form of \(s\) with \(\mathcal{R}\), we conclude that \(t'\) is a reduced form of \(s\) with \(\mathcal{R}\).

- If the current step in the proof tree of \(f\) is \textbf{Stop}, then by Definition 13, \(t \rightarrow t'\) where \(t' = t_{l_c}\) if we have \(C\text{-}RED(u)\), and \(t' = \text{Form}_{ST}(t)\) otherwise.

We prove in a similar manner than in the \textbf{Abstract} case for the \(t|_j\) that \(t'\) is a \(C\)-reduced form of \(t\). Therefore \(t'\) is a \(C\)-term.

This last step is the only step where the process with respect to \(ST\) stops, in the case where \(l = n > 1\). As shown above, this step stops with a \(C\)-form, which ends the proof.