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Checking termination of bottom-up evaluation of logic programs with function symbols*

MARCO CALAUTTI, SERGIO GRECO, FRANCESCA SPEZZANO and IRINA TRUBITSYNA
DIMES, Università della Calabria, 87036 Rende (CS), Italy
(e-mail: {calautti,greco,fspezzano,trubitsyna}@dimes.unical.it)

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Abstract

Recently, there has been an increasing interest in the bottom-up evaluation of the semantics of logic programs with complex terms. The presence of function symbols in the program may render the ground instantiation infinite, and finiteness of models and termination of the evaluation procedure, in the general case, are not guaranteed anymore. Since the program termination problem is undecidable in the general case, several decidable criteria (called program termination criteria) have been recently proposed. However, current conditions are not able to identify even simple programs, whose bottom-up execution always terminates. The paper introduces new decidable criteria for checking termination of logic programs with function symbols under bottom-up evaluation, by deeply analyzing the program structure. First, we analyze the propagation of complex terms among arguments by means of the extended version of the argument graph called propagation graph. The resulting criterion, called $\Gamma$-acyclicity, generalizes most of the decidable criteria proposed so far. Next, we study how rules may activate each other and define a more powerful criterion, called safety. This criterion uses the so-called safety function able to analyze how rules may activate each other and how the presence of some arguments in a rule limits its activation. We also study the application of the proposed criteria to bound queries and show that the safety criterion is well-suited to identify relevant classes of programs and bound queries. Finally, we propose a hierarchy of classes of terminating programs, called $k$-safety, where the $k$-safe class strictly includes the $(k-1)$-safe class.

KEYWORDS: Logic programming with function symbols, bottom-up execution, program termination, stable models

1 Introduction

Recently, there has been an increasing interest in the bottom-up evaluation of the semantics of logic programs with complex terms. Although logic languages under stable model semantics have enough expressive power to express problems in the second level of the polynomial hierarchy, in some cases function symbols make

* This work refines and extends results from the conference paper (Greco et al. 2012).
languages compact and more understandable. For instance, several problems can be
naturally expressed using list and set constructors, and arithmetic operators. The
presence of function symbols in the program may render the ground instantiation
infinite, and finiteness of models and termination of the evaluation procedure, in the
general case, are not guaranteed anymore. Since the program termination problem
is undecidable in the general case, several decidable sufficient conditions (called
program termination criteria) have been recently proposed.

The program termination problem has received a significant attention since the
beginning of logic programming and deductive databases (Krishnamurthy et al.
1996) and has recently received an increasing interest. A considerable body of work
has been done on termination of logic programs under top-down evaluation (Schreye
and Decorte 1994; Marchiori 1996; Ohlebusch 2001; Bonatti 2004; Codish et al.
2005; Serebrenik and De Schreye 2005; Bruynooghe et al. 2007; Nguyen et al. 2007;
Baselice et al. 2009; Schneider-Kamp et al. 2009a; Schneider-Kamp et al. 2009b;
Nishida and Vidal 2010; Schneider-Kamp et al. 2010; Ströder et al. 2010; Voets
and Schreye 2010; Brockschmidt et al. 2012; Liang and Kifer 2013). In this context,
the class of finitary programs, allowing decidable (ground) query computation using
a top-down evaluation, has been proposed in (Bonatti 2004; Baselice et al. 2009).
Moreover, there are other research areas, such as these of term rewriting (Zantema
1995; Ferreira and Zantema 1996; Arts and Giesl 2000; Sternagel and Middeldorp
2008; Endrullis et al. 2008) and chase termination (Fagin et al. 2005; Marnette 2009;
Meier et al. 2009; Greco and Spezzano 2010; Greco et al. 2011), whose results can
be of interest to the logic program termination context.

In this paper, we consider logic programs with function symbols under the stable
model semantics (Gelfond and Lifschitz 1988; Gelfond and Lifschitz 1991) and thus,
all the excellent works mentioned above cannot be straightforwardly applied to
our setting. Indeed, the goal of top-down termination analysis is to detect, for a
given program and query goal, sufficient conditions guaranteeing that the resolution
algorithm terminates. On the other side, the aim of the bottom-up termination
analysis is to guarantee the existence of an equivalent finite ground instantiation
of the input program. Furthermore, as stated in (Schreye and Decorte 1994),
even restricting our attention to the top-down approach, the termination of logic
programs strictly depends on the selection and search rules used in the resolution
algorithm. Considering the different aspects of term rewriting and termination
of logic programs, we address readers to (Schreye and Decorte 1994) (pp. 204–
207).

In this framework, the class of finitely ground programs (FG) has been proposed
in (Calimeri et al. 2008). The key property of this class is that stable models
(answer sets) are computable as for each program $P$ in this class, there exists
a finite and computable subset of its instantiation (grounding), called intelligent
instantiation, having precisely the same answer sets as $P$. Since the problem of
deciding whether a program is in $FG$ is not decidable, decidable subclasses, such as
finite domain programs (Calimeri et al. 2008), $\omega$-restricted programs (Syrjänen 2001),
$\lambda$-restricted programs (Gebser et al. 2007b), and the most general one, argument-
restricted programs (Lierler and Lifschitz 2009), have been proposed.
Current techniques analyze how values are propagated among predicate arguments to detect whether a given argument is limited, i.e. whether the set of values which can be associated with the argument, also called active domain, is finite. However, these methods have limited capacity in comprehending that arguments are limited in the case where different function symbols appear in the recursive rules. Even the argument-restricted criterion, which is one the most general criteria, fails in such cases.

Thus, we propose a new technique, called $\Gamma$-acyclicity, whose aim is to improve the argument-restricted criterion without changing the (polynomial) time complexity of the argument-restricted criterion. This technique makes use of the so-called propagation graph, that represents the propagation of values among arguments and the construction of complex terms during the program evaluation.

Furthermore, since many practical programs are not recognized by current termination criteria, including the $\Gamma$-acyclicity criterion, we propose an even more general technique, called safety, which also analyzes how rules activate each other. The new technique allows us to recognize as terminating many classical programs, still guaranteeing polynomial time complexity.

Example 1
Consider the following program $P_1$ computing the length of a list stored in a fact of the form $input(L)$:

\[
\begin{align*}
  r_0 & : list(L) \leftarrow input(L). \\
  r_1 & : list(L) \leftarrow list([X|L]). \\
  r_2 & : count([],0). \\
  r_3 & : count([X|L],I+1) \leftarrow list([X|L]), count(L,I).
\end{align*}
\]

where $input$ is a base predicate defined by only one fact of the form $input([a,b,...])$. $\blacksquare$

The safety technique, proposed in this paper, allows us to understand that $P_1$ is finitely ground and, therefore, terminating under the bottom-up evaluation.

Contribution.

- We first refine the method proposed in (Lierler and Lifschitz 2009) by introducing the set of restricted arguments and we show that the complexity of finding such arguments is polynomial in the size of the given program.
- We then introduce the class of $\Gamma$-acyclic programs, that strictly extends the class of argument-restricted programs. Its definition is based on a particular graph, called propagation graph, representing how complex terms in non-restricted arguments are created and used during the bottom-up evaluation. We also show that the complexity of checking whether a program is $\Gamma$-acyclic is polynomial in the size of the given program.
- Next we introduce the safety function whose iterative application, starting from the set of $\Gamma$-acyclic arguments, allows us to derive a larger set of limited arguments, by analyzing how rules may be activated. In particular, we define the activation graph that represents how rules may activate each other and
design conditions detecting rules whose activation cannot cause their head
arguments to be non-limited.
• Since new criteria are defined for normal logic programs without negation, we
extend their application to the case of disjunctive logic programs with negative
literals and show that the computation of stable models can be performed
using current ASP systems, by a simple rewriting of the source program.
• We propose the application of the new criteria to bound queries and show
that the safety criterion is well suited to identify relevant classes of programs
and bound queries.
• As a further improvement, we introduce the notion of active paths of length
k and show its applicability in the termination analysis. In particular, we
generalize the safety criterion and show that the $k$-safety criteria define a
hierarchy of terminating criteria for logic programs with function symbols.
• Complexity results for the proposed techniques are also presented. More
specifically, we show that the complexity of deciding whether a program $\mathcal{P}$
is $\Gamma$-acyclic or safe is polynomial in the size of $\mathcal{P}$, whereas the complexity of
deciding whether a program is $k$-safe, with $k > 1$ is exponential.

A preliminary version of this paper has been presented at the 28th International
Conference on Logic Programming (Greco et al. 2012). Although the concepts of
$\Gamma$-acyclic program and safe program have been introduced in the conference paper,
the definitions contained in the current version are different. Moreover, most of the
theoretical results and all complexity results contained in this paper as well as the
definition of $k$-safe program are new.

Organization. The paper is organized as follows. Section 2 introduces basic notions
on logic programming with function symbols. Section 3 presents the argument-
restriction criterion. In Section 4 the propagation of complex terms among arguments
is investigated and the class of $\Gamma$-acyclic programs is defined. Section 5 analyzes
how rules activate each other and introduces the safety criterion. In Section 6
the applicability of the safety criterion to (partially) ground queries is discussed.
Section 7 presents further improvements extending the safety criterion. Finally, in
Section 8 the application of termination criteria to general disjunctive programs
with negated literals is presented.

2 Logic Programs with Function symbols

Syntax. We assume to have infinite sets of constants, variables, predicate symbols,
and function symbols. Each predicate and function symbol $g$ is associated with a
fixed arity, denoted by $\text{ar}(g)$, which is a non-negative integer for predicate symbols
and a natural number for function symbols.

A term is either a constant, a variable, or an expression of the form $f(t_1, \ldots, t_m)$,
where $f$ is a function symbol of arity $m$ and the $t_i$’s are terms. In the first two cases
we say the term is simple while in the last case we say it is complex. The binary
relation subterm over terms is recursively defined as follows: every term is a subterm
of itself; if $t$ is a complex term of the form $f(t_1, \ldots, t_m)$, then every $t_i$ is a subterm of
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$\delta$ for $1 \leq i \leq m$; if $t_1$ is a subterm of $t_2$ and $t_2$ is a subterm of $t_3$, then $t_1$ is a subterm of $t_3$. The depth $d(u,t)$ of a simple term $u$ in a term $t$ that contains $u$ is recursively defined as follows:

$$d(u, u) = 0,$$
$$d(u, f(t_1, ..., t_m)) = 1 + \max_{i : t_i \text{ contains } u} d(u, t_i).$$

The depth of a term $t$, denoted by $d(t)$, is the maximal depth of all simple terms occurring in $t$.

An atom $\alpha$ is of the form $p(t_1, \ldots, t_n)$, where $p$ is a predicate symbol of arity $n$ and the $t_i$'s are terms (we also say that the atom is a $p$-atom). A literal is either an atom $A$ (positive literal) or its negation $\neg A$ (negative literal).

A rule $r$ is of the form:

$$A_1 \lor \ldots \lor A_m \leftarrow B_1, \ldots, B_k, \neg C_1, \ldots, \neg C_n$$

where $m > 0$, $k \geq 0$, $n \geq 0$, and $A_1, \ldots, A_m, B_1, \ldots, B_k, C_1, \ldots, C_n$ are atoms. The disjunction $A_1 \lor \ldots \lor A_m$ is called the head of $r$ and is denoted by $\text{head}(r)$; the conjunction $B_1, \ldots, B_k, \neg C_1, \ldots, \neg C_n$ is called the body of $r$ and is denoted by $\text{body}(r)$.

The positive (resp. negative) body of $r$ is the conjunction $B_1, \ldots, B_k$ (resp. $\neg C_1, \ldots, \neg C_n$) and is denoted by $\text{body}^+(r)$ (resp. $\text{body}^-(r)$). With a slight abuse of notation we use $\text{head}(r)$ (resp. $\text{body}(r)$) to also denote the set of atoms (resp. literals) appearing in the head (resp. positive body, negative body) of $r$. If $m = 1$, then $r$ is normal; if $n = 0$, then $r$ is positive. If a rule $r$ is both normal and positive, then it is standard.

A program $P$ is a finite set of rules. A program is normal (resp. positive, standard) if every rule in it is normal (resp. positive, standard). A term (resp. an atom, a literal, a rule, a program) is said to be ground if no variables occur in it. A ground normal rule with an empty body is also called a fact. For any atom $A$ (resp. set of atoms, rule), $\text{var}(A)$ denotes the set of variables occurring in $A$.

We assume that programs are range restricted, i.e. for each rule, the variables appearing in the head or in negative body literals also appear in some positive body literal.

The definition of a predicate symbol $p$ in a program $P$ consists of all rules in $P$ with $p$ in the head. Predicate symbols are partitioned into two different classes: base predicate symbols, whose definition can contain only facts (called database facts), and derived predicate symbols, whose definition can contain any rule. Database facts are not shown in our examples as they are not relevant for the proposed criteria.

Given a program $P$, a predicate $q$ depends on a predicate $p$ if there is a rule $r$ in $P$ such that $p$ appears in the head and $q$ in the body, or there is a predicate $s$ such that $p$ depends on $s$ and $s$ depends on $q$. A predicate $p$ is said to be recursive if it depends on itself, whereas two predicates $p$ and $q$ are said to be mutually recursive if $p$ depends on $q$ and $q$ depends on $p$. A rule $r$ is said to be recursive if its body contains a predicate symbol mutually recursive with a predicate symbol in the head.

Given a rule $r$, $\text{body}(r)$ denotes the set of body atoms whose predicate symbols are mutually recursive with the predicate symbol of an atom in the head. We say that $r$ is linear if $|\text{body}(r)| \leq 1$. We say that a recursive rule $r$ defining a predicate $p$ is
strongly linear if it is linear, the recursive predicate symbol appearing in the body is p and there are no other recursive rules defining p. A predicate symbol p is said to be linear (resp. strongly linear) if all recursive rules defining p are linear (resp. strongly linear).

A substitution is a finite set of pairs \( \theta = \{X_1/t_1, \ldots, X_n/t_n\}\) where \( t_1, \ldots, t_n \) are terms and \( X_1, \ldots, X_n \) are distinct variables not occurring in \( t_1, \ldots, t_n \). If \( \theta = \{X_1/t_1, \ldots, X_n/t_n\}\) is a substitution and \( T \) is a term or an atom, then \( T\theta \) is the term or atom obtained from \( T \) by simultaneously replacing each occurrence of \( X_i \) in \( T \) by \( t_i (1 \leq i \leq n) \)– \( T\theta \) is called an instance of \( T \).

Given a set \( S \) of terms (or atoms), \( S\theta = \{T\theta \mid T \in S\} \). A substitution \( \theta \) is a unifier for a finite set of terms (or atoms) \( S \) if \( S\theta \) is a singleton. We say that a set of terms (or atoms) \( S \) unifies if there exists a unifier \( \theta \) for \( S \). Given two substitutions, \( \theta = \{X_1/t_1, \ldots, X_n/t_n\}\) and \( \eta = \{Y_1/u_1, \ldots, Y_m/u_m\}\), their composition, denoted by \( \theta \circ \eta \), is the substitution obtained from the set \( \{X_1/t_1\theta, \ldots, X_n/t_n\theta, Y_1/u_1, \ldots, Y_m/u_m\}\) by removing every \( X_i/t_i\theta \) such that \( X_i = t_i\eta \) and every \( Y_j/u_j \) such that \( Y_j \in \{X_1, \ldots, X_n\}\). A substitution \( \theta \) is more general than a substitution \( \eta \) if there exists a substitution \( \eta' \) such that \( \eta = \theta \circ \eta' \).

An interpretation of \( \mathcal{P} \) is any subset \( I \) of \( B_\mathcal{P} \). The truth value of a ground atom \( A \) w.r.t. \( I \), denoted by \( \text{value}_I(A) \), is true if \( A \in I \), false otherwise. The truth value of \( \neg A \) w.r.t. \( I \), denoted by \( \text{value}_I(\neg A) \), is true if \( A \notin I \), false otherwise. The truth value of a conjunction of ground literals \( C = L_1, \ldots, L_n \) w.r.t. \( I \) is \( \text{value}_I(C) = \min(\{\text{value}_I(L_i) \mid 1 \leq i \leq n\}) \)—here the ordering \( false < true \) holds—whereas the truth value of a disjunction of ground literals \( D = L_1 \lor \ldots \lor L_n \) w.r.t. \( I \) is \( \text{value}_I(D) = \max(\{\text{value}_I(L_i) \mid 1 \leq i \leq n\}) \); if \( n = 0 \), then \( \text{value}_I(C) = true \) and \( \text{value}_I(D) = false \). A ground rule \( r \) is satisfied by \( I \), denoted by \( I \models r \), if \( \text{value}_I(\text{head}(r)) \geq \text{value}_I(\text{body}(r)) \); we write \( I \not\models r \) if \( r \) is not satisfied by \( I \).

Thus, a ground rule \( r \) with empty body is satisfied by \( I \) if \( \text{value}_I(\text{head}(r)) = true \). An interpretation of \( \mathcal{P} \) is a model of \( \mathcal{P} \) if it satisfies every ground rule in \( \text{ground}(\mathcal{P}) \).

A model \( M \) of \( \mathcal{P} \) is minimal if no proper subset of \( M \) is a model of \( \mathcal{P} \). The set of minimal models of \( \mathcal{P} \) is denoted by \( M.M(\mathcal{P}) \).

Given an interpretation \( I \) of \( \mathcal{P} \), let \( \mathcal{P}^I \) denote the ground positive program derived from \( \text{ground}(\mathcal{P}) \) by (i) removing every rule containing a negative literal \( \neg A \) in the body with \( A \in I \), and (ii) removing all negative literals from the remaining rules. An interpretation \( I \) is a stable model of \( \mathcal{P} \) if and only if \( I \in M.M(\mathcal{P}^I) \) (Gelfond and Lifschitz 1988; Gelfond and Lifschitz 1991). The set of stable models of \( \mathcal{P} \) is denoted by \( \mathcal{P}.M(\mathcal{P}) \). It is well known that stable models are minimal models (i.e.

Semantics. Let \( \mathcal{P} \) be a program. The Herbrand universe \( H_\mathcal{P} \) of \( \mathcal{P} \) is the possibly infinite set of ground terms which can be built using constants and function symbols appearing in \( \mathcal{P} \). The Herbrand base \( B_\mathcal{P} \) of \( \mathcal{P} \) is the set of ground atoms which can be built using predicate symbols appearing in \( \mathcal{P} \) and ground terms of \( H_\mathcal{P} \). A rule \( r' \) is a ground instance of a rule \( r \) in \( \mathcal{P} \) if \( r' \) can be obtained from \( r \) by substituting every variable in \( r \) with some ground term in \( H_\mathcal{P} \).

We use \( \text{ground}(r) \) to denote the set of all ground instances of \( r \) and \( \text{ground}(\mathcal{P}) \) to denote the set of all ground instances of the rules in \( \mathcal{P} \), i.e. \( \text{ground}(\mathcal{P}) = \cup r \in \mathcal{P} \text{ground}(r) \).

An interpretation of \( \mathcal{P} \) is any subset \( I \) of \( B_\mathcal{P} \). The truth value of a ground atom \( A \) w.r.t. \( I \), denoted by \( \text{value}_I(A) \), is true if \( A \in I \), false otherwise. The truth value of \( \neg A \) w.r.t. \( I \), denoted by \( \text{value}_I(\neg A) \), is true if \( A \notin I \), false otherwise. The truth value of a conjunction of ground literals \( C = L_1, \ldots, L_n \) w.r.t. \( I \) is \( \text{value}_I(C) = \min(\{\text{value}_I(L_i) \mid 1 \leq i \leq n\}) \)—here the ordering \( false < true \) holds—whereas the truth value of a disjunction of ground literals \( D = L_1 \lor \ldots \lor L_n \) w.r.t. \( I \) is \( \text{value}_I(D) = \max(\{\text{value}_I(L_i) \mid 1 \leq i \leq n\}) \); if \( n = 0 \), then \( \text{value}_I(C) = true \) and \( \text{value}_I(D) = false \). A ground rule \( r \) is satisfied by \( I \), denoted by \( I \models r \), if \( \text{value}_I(\text{head}(r)) \geq \text{value}_I(\text{body}(r)) \); we write \( I \not\models r \) if \( r \) is not satisfied by \( I \).

Thus, a ground rule \( r \) with empty body is satisfied by \( I \) if \( \text{value}_I(\text{head}(r)) = true \). An interpretation of \( \mathcal{P} \) is a model of \( \mathcal{P} \) if it satisfies every ground rule in \( \text{ground}(\mathcal{P}) \).

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The active domain of a program \( \mathcal{P} \) is defined as \( \mathcal{A}(\mathcal{P}) \). Furthermore, minimal and stable model semantics coincide for positive programs (i.e. \( \mathcal{M}(\mathcal{P}) = \mathcal{M}(\mathcal{P}) \)). A standard program has a unique minimal model, called minimum model.

Given a set of ground atoms \( S \) and a predicate \( g \) (resp. an atom \( A \)), \( S[g] \) (resp. \( S[A] \)) denotes the set of \( g \)-atoms (resp. ground atoms unifying with \( A \)) in \( S \).

Analogously, for a given set \( M \) of sets of ground atoms, we shall use the following notations \( M[g] = \{S[g] \mid S \in M\} \) and \( M[A] = \{S[A] \mid S \in M\} \). Given a set of ground atoms \( S \), and a set \( G \) of predicates symbols, then \( S[G] = \cup_{g \in G} S[g] \).

**Argument graph.** Given an \( n \)-ary predicate \( p \), \( p[i] \) denotes the \( i \)-th argument of \( p \), for \( 1 \leq i \leq n \). If \( p \) is a base (resp. derived) predicate symbol, then \( p[i] \) is said to be a base (resp. derived) argument. The set of all arguments of a program \( \mathcal{P} \) is denoted by \( \text{args}(\mathcal{P}) \); analogously, \( \text{args}_b(\mathcal{P}) \) and \( \text{args}_d(\mathcal{P}) \) denote the sets of all base and derived arguments, respectively.

For any program \( \mathcal{P} \) and \( n \)-ary predicate \( p \) occurring in \( \mathcal{P} \), an argument \( p[i] \), with \( 1 \leq i \leq n \), is associated with the set of values it can take during the evaluation; this domain, called **active domain** of \( p[i] \), is denoted by \( AD(p[i]) = \{t_i | p(t_1, \ldots, t_n) \in M \land M \in \mathcal{A}(\mathcal{P})\} \). An argument \( p[i] \) is said to be **limited** iff \( AD(p[i]) \) is finite.

The **argument graph** of a program \( \mathcal{P} \), denoted by \( G(\mathcal{P}) \), is a directed graph whose nodes are \( \text{args}(\mathcal{P}) \) (i.e. the arguments of \( \mathcal{P} \)), and there is an edge from \( q[j] \) to \( p[i] \), denoted by \( (q[j], p[i]) \), iff there is a rule \( r \in \mathcal{P} \) such that:

1. an atom \( p(t_1, \ldots, t_n) \) appears in \( \text{head}(r) \),
2. an atom \( q(u_1, \ldots, u_m) \) appears in \( \text{body}^+(r) \), and
3. terms \( t_i \) and \( u_j \) have a common variable.

Consider, for instance, program \( P_1 \) of Example 1. \( G(P_1) = (\text{args}(P_1), E) \), where \( \text{args}(P_1) = \{\text{input}[1], \text{list}[1], \text{count}[1], \text{count}[2]\} \), whereas, considering the occurrences of variables in the rules of \( P_1 \) we have that \( E = \{(\text{input}[1], \text{list}[1]), \)

\((\text{list}[1], \text{list}[1]), (\text{list}[1], \text{count}[1]), (\text{count}[1], \text{count}[1]), (\text{count}[2], \text{count}[2])\} \).

**Labeled directed graphs.** In the following we will also consider labeled directed graphs, i.e. directed graphs with labeled edges. In this case we represent an edge from \( a \) to \( b \) as a triple \((a, b, l)\), where \( l \) denotes the label.

A **path** \( \pi \) from \( a_1 \) to \( b_m \) in a possibly labeled directed graph is a non-empty sequence \((a_1, b_1, l_1), \ldots, (a_m, b_m, l_m)\) of its edges s.t. \( b_i = a_{i+1} \) for all \( 1 \leq i < n \); if the first and last nodes coincide (i.e. \( a_1 = b_m \)), then \( \pi \) is called a cyclic path. In the case where the indication of the starting edge is not relevant, we will call a cyclic path a cycle.

We say that a node \( a \) **depends on** a node \( b \) in a graph iff there is a path from \( b \) to \( a \) in that graph. Moreover, we say that \( a \) **depends on** a cycle \( \pi \) iff it depends on a node \( b \) appearing in \( \pi \). Clearly, nodes belonging to cycle \( \pi \) depend on \( \pi \).

## 3 Argument ranking

The argument ranking of a program has been proposed in (Lierler and Lifschitz 2009) to define the class \( \mathcal{AR} \) of argument-restricted programs.
An argument ranking for a program $\mathcal{P}$ is a partial function $\phi$ from $\text{args}(\mathcal{P})$ to non-negative integers, called ranks, such that, for every rule $r$ of $\mathcal{P}$, every atom $p(t_1, \ldots, t_n)$ occurring in the head of $r$, and every variable $X$ occurring in a term $t_i$, if $\phi(p[i])$ is defined, then body$^+(r)$ contains an atom $q(u_1, \ldots, u_m)$ such that $X$ occurs in a term $u_j$, $\phi(q[j])$ is defined, and the following condition is satisfied:

$$\phi(p[i]) - \phi(q[j]) \geq d(X, t_i) - d(X, u_j).$$

(1)

A program $\mathcal{P}$ is said to be argument-restricted if it has an argument ranking assigning ranks to all arguments of $\mathcal{P}$.

Example 2

Consider the following program $P_2$, where $b$ is a base predicate:

$$\begin{align*}
  r_1 &: p(f(x)) \leftarrow p(x), b(x). \\
  r_2 &: t(f(x)) \leftarrow p(x). \\
  r_3 &: a(x) \leftarrow t(f(x)).
\end{align*}$$

This program has an argument ranking $\phi$, where $\phi(b[1]) = 0$, $\phi(p[1]) = 1$, $\phi(t[1]) = 2$, and $\phi(a[1]) = 1$. Consequently, $P_2$ is argument-restricted.

Intuitively, the rank of an argument is an estimation of the depth of terms that may occur in it. In particular, let $d_1$ be the rank assigned to a given argument $p[i]$ and let $d_2$ be the maximal depth of terms occurring in the database facts. Then $d_1 + d_2$ gives an upper bound of the depth of terms that may occur in $p[i]$ during the program evaluation. Different argument rankings may satisfy condition (1). A function assigning minimum ranks to arguments is denoted by $\phi_{\text{min}}$.

Minimum ranking. We define a monotone operator $\Omega$ that takes as input a function $\phi$ over arguments and gives as output a function over arguments that gives an upper bound of the depth of terms.

More specifically, we define $\Omega(\phi)(p[i])$ as

$$\max(\max\{D(p(t_1, \ldots, t_n), r, i, X) \mid r \in \mathcal{P} \land p(t_1, \ldots, t_n) \in \text{head}(r) \land X \text{ occurs in } t_i\}, 0)$$

where $D(p(t_1, \ldots, t_n), r, i, X)$ is defined as

$$\min\{d(X, t_i) - d(X, u_j) + \phi(q[j]) \mid q(u_1, \ldots, u_m) \in \text{body}^+(r) \land X \text{ occurs in } u_j\}.$$

In order to compute $\phi_{\text{min}}$, we compute the fixpoint of $\Omega$ starting from the function $\phi_0$ that assigns 0 to all arguments. In particular, we have:

$$\begin{align*}
  \phi_0(p[i]) &= 0; \\
  \phi_k(p[i]) &= \Omega(\phi_{k-1})(p[i]) = \Omega^k(\phi_0)(p[i]).
\end{align*}$$

The function $\phi_{\text{min}}$ is defined as follows:

$$\phi_{\text{min}}(p[i]) = \begin{cases} 
  \Omega^k(\phi_0)(p[i]) & \text{if } \exists k \text{ (finite) s.t. } \Omega^k(\phi_0)(p[i]) = \Omega^{k+1}(\phi_0)(p[i]) \\
  \text{undefined} & \text{otherwise}
\end{cases}$$

We denote the set of restricted arguments of $\mathcal{P}$ as $\text{AR}(\mathcal{P}) = \{p[i] \mid p[i] \in \text{args}(\mathcal{P}) \land \phi_{\text{min}}(p[i]) \text{ is defined}\}$. Clearly, from definition of $\phi_{\text{min}}$, it follows that all restricted arguments are limited. Observe that $\mathcal{P}$ is argument-restricted if $\text{AR}(\mathcal{P}) = \text{args}(\mathcal{P})$. 
Example 3
Consider again program $P_2$ from Example 2. The following table shows the first four iterations of $\Omega$ starting from the base ranking function $\phi_0$:

<table>
<thead>
<tr>
<th>$\phi_0$</th>
<th>$\phi_1 = \Omega(\phi_0)$</th>
<th>$\phi_2 = \Omega(\phi_1)$</th>
<th>$\phi_3 = \Omega(\phi_2)$</th>
<th>$\phi_4 = \Omega(\phi_3)$</th>
</tr>
</thead>
<tbody>
<tr>
<td>b[1]</td>
<td>0</td>
<td>0</td>
<td>0</td>
<td>0</td>
</tr>
<tr>
<td>p[1]</td>
<td>0</td>
<td>1</td>
<td>1</td>
<td>1</td>
</tr>
<tr>
<td>t[1]</td>
<td>0</td>
<td>1</td>
<td>2</td>
<td>2</td>
</tr>
<tr>
<td>s[1]</td>
<td>0</td>
<td>0</td>
<td>0</td>
<td>1</td>
</tr>
</tbody>
</table>

Since $\Omega(\phi_3) = \Omega(\phi_2)$, further applications of $\Omega$ provide the same result. Consequently, $\phi_{\text{min}}$ coincides with $\phi_3$ and defines ranks for all arguments of $P_2$.  

Let $M = |\text{args}(\mathcal{P})| \times d_{\text{max}}$, where $d_{\text{max}}$ is the largest depth of terms occurring in the heads of rules of $\mathcal{P}$. One can determine whether $\mathcal{P}$ is argument-restricted by iterating $\Omega$ starting from $\phi_0$ until:

1. one of the values of $\Omega^k(\phi_0)$ exceeds $M$, in such a case $\mathcal{P}$ is not argument-restricted;
2. $\Omega^{k+1}(\phi_0) = \Omega^k(\phi_0)$, in such a case $\phi_{\text{min}}$ coincides with $\phi_k$, $\phi_{\text{min}}$ is total, and $\mathcal{P}$ is argument-restricted.

Observe that if the program is not argument-restricted the first condition is verified with $k \leq M \times |\text{args}(\mathcal{P})| \leq M^2$, as at each iteration the value assigned to at least one argument is changed. Thus, the problem of deciding whether a given program $\mathcal{P}$ is argument-restricted is in $P\text{Time}$. In the following section we will show that the computation of restricted arguments can be done in polynomial time also when $\mathcal{P}$ is not argument-restricted (see Proposition 1).

4 $\Gamma$-acyclic programs
In this section we exploit the role of function symbols for checking program termination under bottom-up evaluation. Starting from this section, we will consider standard logic programs. Only in Section 8 we will refer to general programs, as it discusses how termination criteria defined for standard programs can be applied to general disjunctive logic programs with negative literals. We also assume that if the same variable $X$ appears in two terms occurring in the head and body of a rule respectively, then at most one of the two terms is a complex term and that the nesting level of complex terms is at most one. As we will see in Section 8, there is no real restriction in such an assumption as every program could be rewritten into an equivalent program satisfying such a condition.

The following example shows a program admitting a finite minimum model, but the argument-restricted criterion is not able to detect it. Intuitively, the definition of argument-restricted programs does not take into account the possible presence
of different function symbols in the program that may prohibit the propagation of
values in some rules and, consequently, guarantee the termination of the bottom-up
computation.

Example 4
Consider the following program \( P_4 \):

\[
\begin{align*}
r_0 : & \ a(X) \leftarrow b(X), \\
r_1 : & \ r(f(X)) \leftarrow a(X), \\
r_2 : & \ q(f(X)) \leftarrow r(X), \\
r_3 : & \ a(X) \leftarrow q(g(X)).
\end{align*}
\]

where \( b \) is a base predicate symbol. The program is not argument-restricted since
the argument ranking function \( q_{\min} \) cannot assign any value to \( r[1], q[1], \) and \( a[1] \).
However the bottom-up computation always terminates, independently from the
database instance.

In order to represent the propagation of values among arguments, we introduce
the concept of labeled argument graphs. Intuitively, it is an extension of the argument
graph where each edge has a label describing how the term propagated from one
argument to another changes. Arguments that are not dependent on a cycle can
propagate a finite number of values and, therefore, are limited.

Since the active domain of limited arguments is finite, we can delete edges ending in
the corresponding nodes from the labeled argument graph. Then, the resulting graph,
called propagation graph, is deeply analyzed to identify further limited arguments.

Definition 1 (Labeled argument graph)
Let \( \mathcal{P} \) be a program. The labeled argument graph \( G_L(\mathcal{P}) \) is a labeled directed graph
\((\text{args}(\mathcal{P}), E)\) where \( E \) is a set of labeled edges defined as follows. For each pair of
nodes \( p[i], q[j] \in \text{args}(\mathcal{P}) \) such that there is a rule \( r \) with \( \text{head}(r) = p(v_1, \ldots, v_n) \),
\( q(u_1, \ldots, u_m) \in \text{body}(r) \), and terms \( u_j \) and \( v_i \) have a common variable \( X \), there is an
edge \((q[j], p[i], \alpha) \in E \) such that

- \( \alpha = e \) if \( u_j = v_i = X \),
- \( \alpha = f \) if \( u_j = X \) and \( v_i = f(...,X,...) \),
- \( \alpha = \overline{f} \) if \( u_j = f(...,X,...) \) and \( v_i = X \).

In the definition above, the symbol \( e \) denotes the empty label which concatenated
to a string does not modify the string itself, that is, for any string \( s, se = es = s \).

The labeled argument graph of program \( P_4 \) is shown in Figure 1 (left). The edges
of this graph represent how the propagation of values occurs. For instance, edge
\((b[1], a[1], e)\) states that a term \( t \) is propagated without changes from \( b[1] \) to \( a[1] \) if
rule \( r_0 \) is applied; analogously, edge \((a[1], r[1], f)\) states that starting from a term \( t \)
in \( a[1] \) we obtain \( f(t) \) in \( r[1] \) if rule \( r_1 \) is applied, whereas edge \((q[1], a[1], \overline{g})\) states
that starting from a term \( g(t) \) in \( q[1] \) we obtain \( t \) in \( a[1] \) if rule \( r_3 \) is applied.

Given a path \( \pi \) in \( G_L(\mathcal{P}) \) of the form \( (a_1, b_1, z_1), \ldots, (a_m, b_m, z_m) \), we denote with
\( \lambda(\pi) \) the string \( z_1 \ldots z_m \). We say that \( \pi \) spells a string \( w \) if \( \lambda(\pi) = w \). Intuitively, the
string \( \lambda(\pi) \) describes a sequence of function symbols used to compose and decompose
complex terms during the propagation of values among the arguments in \( \pi \).
Example 5

Consider program $P_5$ derived from program $P_4$ of Example 4 by replacing rule $r_2$ with the rule $q(g(X)) \leftarrow r(X)$. The labeled argument graph $G_L(P_5)$ is reported in Figure 1 (right). Considering the cyclic path $\pi = (s[1], r[1], f)$, $(r[1], q[1], g)$, $(q[1], s[1], g)$, $\lambda(\pi) = fgg$ represents the fact that starting from a term $t$ in $s[1]$ we may obtain the term $g(f(t))$ in $r[1]$, then we may obtain term $g(f(t))$ in $q[1]$, and term $f(t)$ in $s[1]$, and so on. Since we may obtain a larger term in $s[1]$, the arguments depending on this cyclic path may not be limited.

Consider now program $P_4$, whose labeled argument graph is shown in Figure 1 (left), and the cyclic path $\pi' = (s[1], r[1], f)$, $(r[1], q[1], f)$, $(q[1], s[1], g)$. Observe that starting from a term $t$ in $s[1]$ we may obtain term $f(t)$ in $r[1]$ (rule $r_1$), then we may obtain term $f(f(t))$ in $q[1]$ (rule $r_2$). At this point the propagation in this cyclic path terminates since the head atom of rule $r_2$ containing term $f(X)$ cannot match with the body atom of rule $r_3$ containing term $g(X)$. The string $\lambda(\pi') = fgg$ represents the propagation described above. Observe that for this program all arguments are limited.

Let $\pi$ be a path from $p[i]$ to $q[j]$ in the labeled argument graph. Let $\hat{\lambda}(\pi)$ be the string obtained from $\lambda(\pi)$ by iteratively eliminating pairs of the form $\alpha \beta$ until the resulting string cannot be further reduced. If $\hat{\lambda}(\pi) = $, then starting from a term $t$ in $p[i]$ we obtain the same term $t$ in $q[j]$. Consequently, if $\hat{\lambda}(\pi)$ is a non-empty sequence of function symbols $f_i, f_{i+1}, \ldots, f_n$, then starting from a term $t$ in $p[i]$ we may obtain a larger term in $q[j]$. For instance, if $k = 2$ and $f_i$ and $f_j$ are of arity one, we may obtain $f_j(f_i(t))$ in $q[j]$. Based on this intuition we introduce now a grammar $\Gamma_{\wp}$ in order to distinguish the sequences of function symbols used to compose and decompose complex terms in a program $\wp$, such that starting from a given term we obtain a larger term.

Given a program $\wp$, we denote with $F_{\wp} = \{f_1, \ldots, f_m\}$ the set of function symbols occurring in $\wp$, whereas $F_{\wp} = \{f \mid f \in F_{\wp}\}$ and $T_{\wp} = F_{\wp} \cup T_{\wp}$.

Definition 2

Let $\wp$ be a program, the grammar $\Gamma_{\wp}$ is a 4-tuple $(N, T_{\wp}, R, S)$, where $N = \{S, S_1, S_2\}$ is the set of non-terminal symbols, $S$ is the start symbol, and $R$ is the set of production rules defined below:

1. $S \rightarrow S_1 f_1 S_2$, \quad $\forall f_i \in F_{\wp}$;
2. $S_1 \rightarrow f_i S_1 \overline{f_i} S_1 \mid \epsilon$, \quad $\forall f_i \in F_{\wp}$;
3. $S_2 \rightarrow S_1 S_2 \mid f_i S_2 \mid \epsilon$, \quad $\forall f_i \in F_{\wp}$.

The language $L(\Gamma_{\wp})$ is the set of strings generated by $\Gamma_{\wp}$.
Example 6
Let \( F_\mathcal{P} = \{ f, g, h \} \) be the set of function symbols occurring in a program \( \mathcal{P} \). Then strings \( f, g, g, f, f, g, g, h, f, h, g \) belong to \( \mathcal{L}(\Gamma_\mathcal{P}) \) and represent, assuming that \( f \) is a unary function symbol, different ways to obtain term \( f(t) \) starting from term \( t \).

Note that only if a path \( \pi \) spells a string \( w \in \mathcal{L}(\Gamma_\mathcal{P}) \), then starting from a given term in the first node of \( \pi \) we may obtain a larger term in the last node of \( \pi \). Moreover, if this path is cyclic, then the arguments depending on it may not be limited. On the other hand, all arguments not depending on a cyclic path \( \pi \) spelling a string \( w \in \mathcal{L}(\Gamma_\mathcal{P}) \) are limited.

Given a program \( \mathcal{P} \) and a set of arguments \( \mathcal{S} \) recognized as limited by a specific criterion, the propagation graph of \( \mathcal{P} \) w.r.t. \( \mathcal{S} \), denoted by \( \Delta(\mathcal{P}, \mathcal{S}) \), consists of the subgraph derived from \( \mathcal{G}(\mathcal{P}) \) by deleting edges ending in a node of \( \mathcal{S} \). Although we can consider any set \( \mathcal{S} \) of limited arguments, in the following we assume \( \mathcal{S} = \mathcal{A}(\mathcal{P}) \) and, for the simplicity of notation, we denote \( \Delta(\mathcal{P}, \mathcal{A}(\mathcal{P})) \) as \( \Delta(\mathcal{P}) \). Even if more general termination criteria have been defined in the literature, here we consider the \( \mathcal{A}(\mathcal{P}) \) criterion since it is the most general among those so far proposed having polynomial time complexity.

Definition 3 (\( \Gamma \)-acyclic arguments and \( \Gamma \)-acyclic programs)

Given a program \( \mathcal{P} \), the set of its \( \Gamma \)-acyclic arguments, denoted by \( \Gamma\mathcal{A}(\mathcal{P}) \), consists of all arguments of \( \mathcal{P} \) not depending on a cyclic path in \( \Delta(\mathcal{P}) \) spelling a string of \( \mathcal{L}(\Gamma_\mathcal{P}) \). A program \( \mathcal{P} \) is called \( \Gamma \)-acyclic if \( \Gamma\mathcal{A}(\mathcal{P}) = \mathcal{A}(\mathcal{P}) \), i.e. if there is no cyclic path in \( \Delta(\mathcal{P}) \) spelling a string of \( \mathcal{L}(\Gamma_\mathcal{P}) \). We denote the class of \( \Gamma \)-acyclic programs \( \Gamma(\mathcal{P}) \).

Clearly, \( \mathcal{A}(\mathcal{P}) \subseteq \Gamma\mathcal{A}(\mathcal{P}) \), i.e. the set of restricted arguments is contained in the set of \( \Gamma \)-acyclic arguments. As a consequence, the set of argument-restricted programs is a subset of the set of \( \Gamma \)-acyclic programs. Moreover, the containment is strict, as there exist programs that are \( \Gamma \)-acyclic, but not argument-restricted. For instance, program \( P_4 \) from Example 4 is \( \Gamma \)-acyclic, but not argument-restricted. Indeed, all cyclic paths in \( \Delta(P_4) \) do not spell strings belonging to the language \( \mathcal{L}(\Gamma_\mathcal{P}) \).

The importance of considering the propagation graph instead of the labeled argument graph in Definition 3 is shown in the following example.

Example 7
Consider program \( P_7 \) below obtained from \( P_4 \) by adding rules \( r_4 \) and \( r_5 \).

\[
\begin{align*}
    r_0 & : a(X) \leftarrow b(X), \\
    r_1 & : r(f(X)) \leftarrow s(X), \\
    r_2 & : q(f(X)) \leftarrow r(X), \\
    r_3 & : a(X) \leftarrow q(g(X)), \\
    r_4 & : n(f(X)) \leftarrow s(X), b(X), \\
    r_5 & : a(X) \leftarrow n(X).
\end{align*}
\]

The corresponding labeled argument graph \( \mathcal{G}_L(P_7) \) and propagation graph \( \Delta(P_7) \) are reported in Figure 2. Observe that arguments \( n[1] \) and \( s[1] \) are involved in
the red cycle in the labeled argument graph $\mathcal{G}_L(P_7)$ spelling a string of $\mathcal{L}(\Gamma_{P_7})$. At the same time this cycle is not present in the propagation graph $\Delta(P_7)$ since $AR(P_7) = \{b[1], n[1]\}$ and the program is $\Gamma$-acyclic. 

**Theorem 1**

Given a program $\mathcal{P}$,

1. all arguments in $\Gamma A(\mathcal{P})$ are limited;
2. if $\mathcal{P}$ is $\Gamma$-acyclic, then $\mathcal{P}$ is finitely ground.

**Proof**

(1) As previously recalled, arguments in $AR(\mathcal{P})$ are limited. Let us now show that all arguments in $\Gamma A(\mathcal{P}) \setminus AR(\mathcal{P})$ are limited too. Suppose by contradiction that $q[k] \in \Gamma A(\mathcal{P}) \setminus AR(\mathcal{P})$ is not limited. Observe that depth of terms that may occur in $q[k]$ depends on the paths in the propagation graph $\Delta(\mathcal{P})$ that ends in $q[k]$. In particular, this depth may be infinite only if there is a path $\pi$ from an argument $p[i]$ to $q[k]$ (not necessarily distinct from $p[i]$), such that $\lambda(\pi)$ is a string of an infinite length composed by symbols in $F_{P_7}$. But this is possible only if this path contains a cycle spelling a string in $\mathcal{L}(\Gamma_{P_7})$. Thus we obtain contradiction with Definition 3.

(2) From the previous proof, it follows that every argument in the $\Gamma$-acyclic program can take values only from a finite domain. Consequently, the set of all possible ground terms derived during the grounding process is finite and every $\Gamma$-acyclic program is finitely ground.

From the previous theorem we can also conclude that all $\Gamma$-acyclic programs admit a finite minimum model, as this is a property of finitely ground programs.

We conclude by observing that since the language $\mathcal{L}(\Gamma_{\mathcal{P}})$ is context-free, the analysis of paths spelling strings in $\mathcal{L}(\Gamma_{\mathcal{P}})$ can be carried out using pushdown automata.

As $\Gamma_{\mathcal{P}}$ is context free, the language $\mathcal{L}(\Gamma_{\mathcal{P}})$ can be recognized by means of a pushdown automaton $M = (\{q_0, q_F\}, T_{\mathcal{P}}, \Lambda, \delta, q_0, Z_0, \{q_F\})$, where $q_0$ is the initial state, $q_F$ is the final state, $\Lambda = \{Z_0\} \cup \{f_i| f_i \in F_{\mathcal{P}}\}$ is the stack alphabet, $Z_0$ is the initial stack symbol, and $\delta$ is the transition function defined as follows:

1. $\delta(q_0, f_i, Z_0) = (q_F, F_iZ_0), \quad \forall f_i \in F_{\mathcal{P}}$,
2. $\delta(q_F, f_i, F_j) = (q_F, F_iF_j), \quad \forall f_i \in F_{\mathcal{P}}$,
3. $\delta(q_F, \bar{F}_j, F_j) = (q_F, \epsilon)$, $\forall f_i \in F_P$.

The input string is recognized if after having scanned the entire string the
automaton is in state $q_F$ and the stack contains at least one symbol $F_i$.

A path $\pi$ is called:

- increasing, if $\hat{\lambda}(\pi) \in L(\Gamma_P)$,
- flat, if $\hat{\lambda}(\pi) = \epsilon$,
- failing, otherwise.

It is worth noting that $\lambda(\pi) \in L(\Gamma_P)$ iff $\hat{\lambda}(\pi) \in L(\Gamma_P)$ as function $\hat{\lambda}$ emulates the
pushdown automaton used to recognize $L(\Gamma_P)$. More specifically, for any path $\pi$
and relative string $\lambda(\pi)$ we have that:

- if $\pi$ is increasing, then the pushdown automaton recognizes the string $\lambda(\pi)$ in
  state $q_F$ and the stack contains a sequence of symbols corresponding to the
  symbols in $\hat{\lambda}(\pi)$ plus the initial stack symbol $Z_0$;
- if $\pi$ is flat, then the pushdown automaton does not recognize the string $\lambda(\pi)$;
  moreover, the entire input string is scanned, but the stack contains only the
  symbol $Z_0$;
- if $\hat{\lambda}(\pi)$ is failing, then the pushdown automaton does not recognize the string
  $\lambda(\pi)$ as it goes in an error state.

**Complexity.** Concerning the complexity of checking whether a program is $\Gamma$-acyclic,
we first introduce definitions and results that will be used hereafter. We start by
introducing the notion of size of a logic program.

We assume that simple terms have constant size and, therefore, the size of a
complex term $f(t_1, \ldots, t_k)$, where $t_1, \ldots, t_k$ are simple terms, is bounded by $O(k)$.
Analogously, the size of an atom $p(t_1, \ldots, t_n)$ is given by the sum of the sizes of the
t_i’s, whereas the size of a conjunction of atoms (resp. rule, program) is given by the
sum of the sizes of its atoms. That is, we identify for a program $P$ the following
parameters: $n_r$ is the number of rules of $P$, $n_b$ is the maximum number of atoms
in the body of rules of $P$, $a_p$ is the maximum arity of predicate symbols occurring
in $P$, and $a_f$ is the maximum arity of function symbols occurring in $P$. We assume
that the size of $P$, denoted by $\text{size}(P)$, is bounded by $O(n_r \times n_b \times a_p \times a_f)$. Finally,
since checking whether a program is terminating requires to read the program, we
assume that the program has been already scanned and stored using suitable data
structures. Thus, all the complexity results presented in the rest of the paper do
not take into account the cost of scanning and storing the input program. We first
introduce a tighter bound for the complexity of computing $AR(\mathcal{P})$.

**Proposition 1**

For any program $\mathcal{P}$, the time complexity of computing $AR(\mathcal{P})$ is bounded by
$O(|\text{args}(\mathcal{P})|^3)$.

**Proof**

Assume that $n = |\text{args}(\mathcal{P})|$ is the total number of arguments of $\mathcal{P}$. First, it is
important to observe the connection between the behavior of operator $\Omega$ and the
structure of the labeled argument graph $\mathcal{G}_L(\mathcal{P})$. In particular, if the applications of the operator $\Omega$ change the rank of an argument $q[i]$ from 0 to $k$, then there is a path from an argument to $q[i]$ in $\mathcal{G}_L(\mathcal{P})$, where the number of edges labeled with some positive function symbol minus the number of edges labeled with some negative function symbol is at least $k$. Given a cycle in a labeled argument graph, let us call it affected if the number of edges labeled with some positive function symbol is greater than the number of edges labeled with some negative function symbol.

If an argument is not restricted, it is involved in or depends on an affected cycle. On the other hand, if after an application of $\Omega$ the rank assigned to an argument exceeds $n$, this argument is not restricted (Lierler and Lifschitz 2009). Recall that we are assuming that $d_{\max} = 1$ and, therefore, $M = n \times d_{\max} = n$.

Now let us show that after $2n^2 + n$ iterations of $\Omega$ all not restricted arguments exceed rank $n$. Consider an affected cycle and suppose that it contains $k$ arguments, whereas the number of arguments depending on this cycle, but not belonging to it is $m$. Obviously, $k + m \leq n$. All arguments involved in this cycle change their rank by at least one after $k$ iterations of $\Omega$. Thus their ranks will be greater than $n + m$ after $(n + m + 1) \times k$ iterations. The arguments depending on this cycle, but not belonging to it, need at most another $m$ iterations to reach the rank greater than $n$. Thus all unrestricted arguments exceed the rank $n$ in $(n + m + 1) \times k + m$ iterations of $\Omega$. Since $(n + m + 1) \times k + m = n(k + m) + mk + (k + m) \leq 2n^2 + n$, the restricted arguments are those that at step $2n^2 + n$ do not exceed rank $n$. It follows that the complexity of computing $AR(\mathcal{P})$ is bounded by $O(n^3)$ because we have to do $O(n^2)$ iterations and, for each iteration we have to check the rank of $n$ arguments. \hfill $\square$

In order to study the complexity of computing $\Gamma$-acyclic arguments of a program, we introduce a directed (not labeled) graph obtained from the propagation graph.

**Definition 4 (Reduction of $\Delta(\mathcal{P})$)**

Given a program $\mathcal{P}$, the reduction of $\Delta(\mathcal{P})$ is a directed graph $\Delta_R(\mathcal{P})$ whose nodes are the arguments of $\mathcal{P}$ and there is an edge $(p[i], q[j])$ in $\Delta_R(\mathcal{P})$ iff there is a path $\pi$ from $p[i]$ to $q[j]$ in $\Delta(\mathcal{P})$ such that $\lambda(\pi) \in F_p$. \hfill $\square$

The reduction $\Delta_R(\mathcal{P})$ of the propagation graph $\Delta(\mathcal{P})$ from Figure 3 is shown in Figure 4. It is simple to note that for each path in $\Delta(\mathcal{P})$ from node $p[i]$ to node $q[j]$ spelling a string of $\mathcal{L}(\Gamma_\mathcal{P})$ there exists a path from $p[i]$ to $q[j]$ in $\Delta_R(\mathcal{P})$ and vice versa. As shown in the lemma below, this property always holds.

**Lemma 1**

Given a program $\mathcal{P}$ and arguments $p[i], q[j] \in args(\mathcal{P})$, there exists a path in $\Delta(\mathcal{P})$ from $p[i]$ to $q[j]$ spelling a string of $\mathcal{L}(\Gamma_\mathcal{P})$ iff there is a path from $p[i]$ to $q[j]$ in $\Delta_R(\mathcal{P})$.\hfill $\square$
Fig. 4. (Colour online) Reduction $\Delta_R(\mathcal{P})$ of propagation graph $\Delta(\mathcal{P})$

\begin{proof}

($\Rightarrow$) Suppose there is a path $\pi$ from $p[i]$ to $q[j]$ in $\Delta(\mathcal{P})$ such that $\hat{\lambda}(\pi) \in \mathcal{L}(\Gamma_\mathcal{P})$.

Then $\hat{\lambda}(\pi)$ is a non-empty string, say $f_1, \ldots, f_k$, where $f_i \in F_\mathcal{P}$ for $i \in [1, \ldots, k]$.

Consequently, $\pi$ can be seen as a sequence of subpaths $\pi_1, \ldots, \pi_k$, such that $\hat{\lambda}(\pi_i) = f_i$ for $i \in [1, \ldots, k]$. Thus, from the definition of the reduction of $\Delta(\mathcal{P})$, there is a path from $p[i]$ to $q[j]$ in $\Delta_R(\mathcal{P})$ whose length is equal to $|\hat{\lambda}(\pi)|$.

($\Leftarrow$) Suppose there is a path $(n_1, n_2, \ldots, n_k, n_{k+1})$ from $n_1$ to $n_{k+1}$ in $\Delta_R(\mathcal{P})$. From the definition of the reduction of $\Delta(\mathcal{P})$, for each edge $(n_i, n_{i+1})$ there is a path, say $\pi_i$, from $n_i$ to $n_{i+1}$ in $\Delta(\mathcal{P})$ such that $\hat{\lambda}(\pi_i) \in F_\mathcal{P}$. Consequently, there is a path from $n_1$ to $n_{k+1}$ in $\Delta(\mathcal{P})$, obtained as a sequence of paths $\pi_1, \ldots, \pi_k$ whose string is simply $\hat{\lambda}(\pi_1), \ldots, \hat{\lambda}(\pi_k)$. Since $\hat{\lambda}(\pi_i) \in F_\mathcal{P}$ implies that $\hat{\lambda}(\pi_i) \in \mathcal{L}(\Gamma_\mathcal{P})$, for every $1 \leq i \leq k$, we have that $\hat{\lambda}(\pi_1), \ldots, \hat{\lambda}(\pi_k)$ belongs also to $\mathcal{L}(\Gamma_\mathcal{P})$.

\end{proof}

\begin{proposition}

Given a program $\mathcal{P}$, the time complexity of computing the reduction $\Delta_R(\mathcal{P})$ is bounded by $O(|\text{args}(\mathcal{P})|^3 \times |F_\mathcal{P}|)$.

\end{proposition}

\begin{proof}

The construction of $\Delta_R(\mathcal{P})$ can be performed as follows. First, we compute all the paths $\pi$ in $\Delta(\mathcal{P})$ such that $|\hat{\lambda}(\pi)| \leq 1$. To do so, we use a slight variation of the Floyd–Warshall’s transitive closure of $\Delta(\mathcal{P})$ which is defined by the following recursive formula. Assume that each node of $\Delta(\mathcal{P})$ is numbered from 1 to $n = |\text{args}(\mathcal{P})|$, then we denote with $\text{path}(i, j, x, k)$ the existence of a path $\pi$ from node $i$ to node $j$ in $\Delta(\mathcal{P})$ such that $\hat{\lambda}(\pi) = x$, $|x| \leq 1$ and $\pi$ may go only through nodes in $[1, \ldots, k]$ (except for $i$ and $j$).

The set of atoms $\text{path}(i, j, x, k)$, for all values $1 \leq i, j, x \leq n$, can be derived iteratively as follows:

- (base case: $k = 0$) $\text{path}(i, j, x, 0)$ holds if there is an edge $(i, j, x) \in \Delta(\mathcal{P})$,
- (inductive case: $0 < k \leq n$) $\text{path}(i, j, x, k)$ holds if
  - $\text{path}(i, j, x, k-1)$ holds, or
  - $\text{path}(i, k, x_1, k-1)$ and $\text{path}(k, j, x_2, k-1)$ hold, $x = x_1x_2$ and $|x| \leq 1$.

Note that in order to compute all the possible atoms $\text{path}(i, j, x, k)$, we need to first initialize every base atom $\text{path}(i, j, x, 0)$ with cost bounded by $O(n^2 \times |F_\mathcal{P}|)$, as this is the upper bound for the number of edges in $\Delta(\mathcal{P})$. Then, for every $1 \leq k \leq n$, we need to compute all paths, $\text{path}(i, j, x, k)$, thus requiring a cost bounded by $O(n^3 \times |F_\mathcal{P}|)$ operations. The whole procedure will require $O(n^3 \times |F_\mathcal{P}|)$ operations.

Since we have computed all possible paths $\pi$ in $\Delta(\mathcal{P})$ such that $|\hat{\lambda}(\pi)| \leq 1$, we can
obtain all the edges \((i, j)\) of \(\Delta_R(P)\) (according to Definition 4) by simply selecting the atoms \(\text{path}(i, j, z, k)\) with \(z \in F_P\), whose cost is bounded by \(O(n^2 \times |F_P|)\). Then, the time complexity of constructing \(\Delta_R(P)\) is \(O(n^2 \times |F_P|)\).

**Theorem 2**

The complexity of deciding whether a program \(P\) is \(\Gamma\)-acyclic is bounded by 
\[ O(|\text{args}(P)|^3 \times |F_P|). \]

**Proof**

Assume that \(n = |\text{args}(P)|\) is the total number of arguments of \(P\). To check whether \(P\) is \(\Gamma\)-acyclic it is sufficient to first compute the set of restricted arguments \(AR(P)\) which requires time \(O(n^3)\) from Proposition 1. Then, we need to construct the propagation graph \(\Delta(P)\), for which the maximum number of edges is \(n^2 \times (|F_P| + |F_P| + 1)\), then it can be constructed in time \(O(n^2 \times |F_P|)\) (recall that we are not taking into account the cost of scanning and storing the program). Moreover, starting from \(\Delta(P)\), we need to construct \(\Delta_R(P)\), which requires time \(O(n^2 \times F_P)\) (cf. Proposition 2) and then, following Lemma 1, we need to check whether \(\Delta_R(P)\) is acyclic. Verifying whether \(\Delta_R(P)\) is acyclic can be done by means of a simple traversal of \(\Delta_R(P)\) and checking if a node is visited more than once. The complexity of a depth-first traversal of a graph is well known to be \(O(|E|)\) where \(E\) is the set of edges of the graph. Since the maximum number of edges of \(\Delta_R(P)\) is by definition \(n^2 \times |F_P|\), the traversal of \(\Delta_R(P)\) can be done in time \(O(n^2 \times |F_P|)\). Thus, the whole time complexity is still bounded by \(O(n^2 \times |F_P|)\).

**Corollary 1**

For any program \(P\), the time complexity of computing \(\Gamma_A(P)\) is bounded by 
\[ O(|\text{args}(P)|^3 \times |F_P|). \]

**Proof**

Straightforward from the proof of Theorem 2.

As shown in the previous theorem, the time complexity of checking whether a program \(P\) is \(\Gamma\)-acyclic is bounded by \(O(|\text{args}(P)|^3 \times |F_P|)\), which is strictly related to the complexity of checking whether a program is argument-restricted, which is \(O(|\text{args}(P)|^3)\). In fact, the new proposed criterion performs a more accurate analysis on how terms are propagated from the body to the head of rules by taking into account the function symbols occurring in such terms. Moreover, if a logic program \(P\) has only one function symbol, the time complexity of checking whether \(P\) is \(\Gamma\)-acyclic is the same as the one required to check if it is argument-restricted.

**5 Safe programs**

The \(\Gamma\)-acyclicity termination criterion presents some limitations, since it is not able to detect when a rule can be activated only a finite number of times during the bottom-up evaluation of the program. The next example shows a simple terminating program which is not recognized by the \(\Gamma\)-acyclicity termination criterion.
Consider the following logic program $P_8$:

$$r_1 : p(X, X) ← b(X).$$
$$r_2 : p(f(X), g(X)) ← p(X, X).$$

where $b$ is base predicate. As the program is standard, it has a (finite) unique minimal model, which can be derived using the classical bottom-up fixpoint computation algorithm. Moreover, independently from the set of base facts defining $b$, the minimum model of $P_8$ is finite and its computation terminates. □

Observe that the rules of program $P_8$ can be activated at most $n$ times, where $n$ is the cardinality of the active domain of the base predicate $b$. Indeed, the recursive rule $r_2$ cannot activate itself since the newly generated atom is of the form $p(f(a), g(a))$ and does not unify with its body.

As another example consider the recursive rule $q(f(X)) ← q(X), t(X)$ and the strongly linear rule $p(f(X), g(Y)) ← p(Y, X), t(X)$ where $t[1]$ is a limited argument. The activation of these rules is limited by the cardinality of the active domain of $t[1]$.

Thus, in this section, in order to define a more general termination criterion we introduce the safety function which, by detecting rules that can be executed only a finite number of times, derives a larger set of limited arguments of the program. We start by analyzing how rules may activate each other.

**Definition 5 (Activation graph)**

Let $\mathcal{P}$ be a program and let $r_1$ and $r_2$ be (not necessarily distinct) rules of $\mathcal{P}$. We say that $r_1$ activates $r_2$ iff $\text{head}(r_1)$ and an atom in $\text{body}(r_2)$ unify. The activation graph $\Sigma(\mathcal{P}) = (\mathcal{P}, E)$ consists of the set of nodes denoting the rules of $\mathcal{P}$ and the set of edges $(r_i, r_j)$, with $r_i, r_j \in \mathcal{P}$, such that $r_i$ activates $r_j$. □

**Example 9**

Consider program $P_8$ of Example 8. The activation graph of this program contains two nodes $r_1$ and $r_2$ and an edge from $r_1$ to $r_2$. Rule $r_1$ activates rule $r_2$ as the head atom $p(X, X)$ of $r_1$ unifies with the body atom $p(X, X)$ of $r_2$. Intuitively, this means that the execution of the first rule may cause the second rule to be activated. In fact, the execution of $r_1$ starting from the database instance $D = \{b(a)\}$ produces the new atom $p(a, a)$. The presence of this atom allows the second rule to be activated, since the body of $r_2$ can be made true by means of the atom $p(a, a)$, producing the new atom $p(f(a), g(a))$. It is worth noting that the second rule cannot activate itself since $\text{head}(r_2)$ does not unify with the atom $p(X, X)$ in $\text{body}(r_2)$. □

The activation graph shows how rules may activate each other, and, consequently, the possibility to propagate values from one rule to another. Clearly, the active domain of an argument $p[i]$ can be infinite only if $p$ is the head predicate of a rule that may be activated an infinite number of times. A rule may be activated an infinite number of times only if it depends on a cycle of the activation graph. Therefore, a rule not depending on a cycle can only propagate a finite number of values into its head arguments. Another important aspect is the structure of rules and the presence of limited arguments in their body and head atoms. As discussed at the beginning...
Consider the following program $P_{10}$:

$$r_1: \ p(f(X), g(Y)) \leftarrow p(X, Y), b(X).$$
$$r_2: \ q(f(Y)) \leftarrow p(X, Y), q(Y).$$

where $b$ is base predicate. Let $A = \Gamma A(\mathcal{P}) = \{b[1], p[1]\}$. The activation and the propagation graphs of this program are reported in Figure 5. The application of
the safety function to the set of limited arguments $A$ gives $\Psi(A) = \{b[1], p[1], p[2]\}$.

Indeed:

- $b[1] \in \Psi(A)$ since $b$ is a base predicate which does not appear in the head of any rule; consequently all the rules with $b$ in the head (i.e. the empty set) trivially satisfy the conditions of Definition 7.
- $p[1] \in \Psi(A)$ because the unique rule with $p$ in the head (i.e. $r_1$) satisfies the first condition of Definition 6, that is, $r_1$ limits the term $\mathfrak{f}(X)$ w.r.t. $A$ in the head of rule $r_1$ corresponding to argument $p[1]$.
- Since $r_1$ is strongly linear and the second condition of Definition 6 is satisfied, $p[2] \in \Psi(A)$ as well.

The following proposition shows that the safety function can be used to derive further limited arguments.

**Proposition 3**

Let $\mathcal{P}$ be a program and let $A$ be a set of limited arguments of $\mathcal{P}$. Then, all arguments in $\Psi(A)$ are also limited.

**Proof**

Consider an argument $q[i] \in \Psi(A)$, then for every rule $r = q(t_1, \ldots, t_n) \leftarrow \text{body}(r)$ either $r$ does not depend on a cycle of $\Sigma(\mathcal{P})$ or $t_i$ is limited in $r$ w.r.t. $A$.

Clearly, if $r$ does not depend on a cycle of $\Sigma(\mathcal{P})$, it can be activated a finite number of times as it is not ‘effectively recursive’ and does not depend on rules which are effectively recursive.

Moreover, if $t_i$ is limited in $r$ w.r.t. $A$, we have that either:

1. The first condition of Definition 6 is satisfied (i.e. every variable $X$ appearing in $t_i$ also appears in an argument in body($r$) belonging to $A$). This means that variables in $t_i$ can be replaced by a finite number of values.
2. The second condition of Definition 6 is satisfied. Let $p(t_1, \ldots, t_n) = \text{head}(r)$, the condition that all terms $t_1, \ldots, t_n$ must be simple or complex guarantees that, if terms in head($r$) grow, then they grow all together (Conditions 2.a and 2.b). Moreover, if the growth of a term $t_j$ is blocked (Condition 2.c), the growth of all terms (including $t_i$) is blocked too.

Therefore, if one of the two conditions is satisfied for all rules defining $q_i$, the active domain of $q[i]$ is finite.

Unfortunately, as shown in the following example, the relationship $A \subseteq \Psi(A)$ does not always hold for a generic set of arguments $A$, even if the arguments in $A$ are limited.

**Example 11**

Consider the following program $P_{11}$:

\[
\begin{align*}
  r_1 : & \quad p(f(X), Y) \leftarrow q(X), r(Y). \\
  r_2 : & \quad q(X) \leftarrow p(X, Y). \\
  r_3 : & \quad t(Y) \leftarrow r(Y). \\
  r_4 : & \quad s(Y) \leftarrow t(Y). \\
  r_5 : & \quad r(Y) \leftarrow s(Y).
\end{align*}
\]

For any logic program $\mathcal{P}$:

1. $\Gamma A(\mathcal{P}) \subseteq \Psi(\Gamma A(\mathcal{P}))$;
2. $\Psi(\Gamma A(\mathcal{P})) \subseteq \Psi^{i+1}(\Gamma A(\mathcal{P}))$ for $i > 0$.

**Proof**

(1) Suppose that $q[k] \in \Gamma A(\mathcal{P})$. Then $q[k] \in AR(\mathcal{P})$ or $q[k]$ does not depend on a cycle in $A(\mathcal{P})$ spelling a string of $2(\Gamma A)$. In both cases $q[k]$ can depend only on arguments in $\Gamma A(\mathcal{P})$. If $q[k]$ does not depend on any argument, then it does not appear in the head of any rule and, consequently, $q[k] \in \Psi(\Gamma A(\mathcal{P}))$. Otherwise, the first condition of Definition 6 is satisfied and $q[k] \in \Psi(\Gamma A(\mathcal{P}))$.

(2) We prove that $\Psi^i(\Gamma A(\mathcal{P})) \subseteq \Psi^{i+1}(\Gamma A(\mathcal{P}))$ for $i > 0$ by induction. We start by showing that $\Psi^i(\Gamma A(\mathcal{P})) \subseteq \Psi^{i+1}(\Gamma A(\mathcal{P}))$ for $i = 1$, i.e. that the relation $\Psi(\Gamma A(\mathcal{P})) \subseteq \Psi(\Psi(\Gamma A(\mathcal{P})))$ holds. In order to show this relation we must show that for every argument $q[k] \in \mathcal{P}$ if $q[k] \in \Psi(\Gamma A(\mathcal{P}))$, then $q[k] \in \Psi(\Psi(\Gamma A(\mathcal{P})))$.

Consider $q[k] \in \Psi(\Gamma A(\mathcal{P}))$. Then, $q[k]$ satisfies Definition 7 w.r.t. $A = \Gamma A(\mathcal{P})$.

From comma one of this proof it follows that $\Gamma A(\mathcal{P}) \subseteq \Psi(\Gamma A(\mathcal{P}))$, consequently $q[k]$ satisfies Definition 7 w.r.t. $A = \Psi(\Gamma A(\mathcal{P}))$.

Suppose that $\Psi^i(\Gamma A(\mathcal{P})) \subseteq \Psi^{i+1}(\Gamma A(\mathcal{P}))$ for $k > 0$. In order to show that $\Psi^{k+1}(\Gamma A(\mathcal{P})) \subseteq \Psi^{k+2}(\Gamma A(\mathcal{P}))$ we must show that for every argument $q[k] \in \mathcal{P}$ if $q[k] \in \Psi^{k+1}(\Gamma A(\mathcal{P})))$, then $q[k] \in \Psi^{k+2}(\Gamma A(\mathcal{P})))$.

Consider $q[k] \in \Psi^{k+1}(\Gamma A(\mathcal{P}))$. Then $q[k]$ satisfies Definition 7 w.r.t. $A = \Psi^k(\Gamma A(\mathcal{P}))$. Since $\Psi^k(\Gamma A(\mathcal{P})) \subseteq \Psi^{k+1}(\Gamma A(\mathcal{P}))$, $q[k]$ satisfies Definition 7 w.r.t. $A = \Psi^{k+1}(\Gamma A(\mathcal{P}))$ too. Consequently, $q[k] \in \Psi^{k+2}(\Gamma A(\mathcal{P}))$.

Observe that we can prove in a similar way that $AR(\mathcal{P}) \subseteq \Psi(AR(\mathcal{P}))$ and that $\Psi^i(AR(\mathcal{P})) \subseteq \Psi^{i+1}(AR(\mathcal{P}))$ for $i > 0$. 

[Fig. 6. (Colour online) Activation graph of program $P_{11}$.]
Definition 8 (Safe arguments and safe programs)

For any program $\mathcal{P}$, $\text{safe}(\mathcal{P}) = \Psi^\infty(\Gamma_A(\mathcal{P}))$ denotes the set of safe arguments of $\mathcal{P}$. A program $\mathcal{P}$ is said to be safe if all arguments are safe. The class of safe programs will be denoted by $\mathcal{SP}$. □

Clearly, for any set of arguments $A \subseteq \Gamma_A(\mathcal{P})$, $\Psi^{i}(A) \subseteq \Psi^{i}(\Gamma_A(\mathcal{P}))$. Moreover, as shown in Proposition 4, when the starting set is $\Gamma_A(\mathcal{P})$, the sequence $\Gamma_A(\mathcal{P}), \Psi(\Gamma_A(\mathcal{P})), \Psi^2(\Gamma_A(\mathcal{P})), \ldots$ is monotone and there is a finite $n = O(|\text{args}(\mathcal{P})|)$ such that $\Psi^n(\Gamma_A(\mathcal{P})) = \Psi^\infty(\Gamma_A(\mathcal{P}))$. We can also define the inflactionary version of $\Psi$ as $\hat{\Psi}(A) = A \cup \Psi(A)$, obtaining that $\hat{\Psi}(\Gamma_A(\mathcal{P})) = \hat{\Psi}(\Gamma_A(\mathcal{P}))$, for all natural numbers $i$. The introduction of the inflactionary version guarantees that the sequence $A, \hat{\Psi}(A), \hat{\Psi}^2(A), \ldots$ is monotone for every set $A$ of limited arguments. This would allow us to derive a (possibly) larger set of limited arguments starting from any set of limited arguments.

Example 12

Consider again program $P_8$ of Example 8. Although $AR(P_8) = \emptyset$, the program $P_8$ is safe as $\Sigma(P_8)$ is acyclic.

Consider now the program $P_{10}$ of Example 10. As already shown in Example 10, the first application of the safety function to the set of $\Gamma$-acyclic arguments of $P_{10}$ gives $\Psi(\Gamma_A(P_{10})) = \{b[1], p[1], p[2]\}$. The application of the safety function to the obtained set gives $\Psi(\Psi(\Gamma_A(P_{10}))) = \{b[1], p[1], p[2], q[1]\}$. In fact, in the unique rule defining $q$, term $\epsilon(Y)$, corresponding to the argument $q[1]$, is limited in $r$ w.r.t. $\{b[1], p[1], p[2]\}$ (i.e. the variable $Y$ appears in $\text{body}(r)$ in a term corresponding to argument $p[2]$ and argument $p[2]$, belonging to the input set, is limited). At this point, all arguments of $P_{10}$ belong to the resulting set. Thus, $\text{safe}(P_{10}) = \text{args}(P_{10})$, and we have that program $P_{10}$ is safe. □

We now show results on the expressivity of the class $\mathcal{SP}$ of safe programs.

Theorem 3

The class $\mathcal{SP}$ of safe programs strictly includes the class $\mathcal{A}$ of $\Gamma$-acyclic programs and is strictly contained in the class $\mathcal{FG}$ of finitely ground programs.

Proof

($\mathcal{A} \subseteq \mathcal{SP}$). From Proposition 4 it follows that $\Gamma_A \subseteq \mathcal{SP}$. Moreover, $\Gamma_A \subseteq \mathcal{SP}$ as program $P_{10}$ is safe but not $\Gamma$-acyclic.

($\mathcal{SP} \subseteq \mathcal{FG}$). From Proposition 3 it follows that every argument in the safe program can take values only from a finite domain. Consequently, the set of all possible ground terms derived during the grounding process is finite and the program is finitely ground. Moreover, we have that the program $P_{16}$ of Example 16 is finitely ground, but not safe. □

As a consequence of Theorem 3, every safe program admits a finite minimum model.

Complexity. We start by introducing a bound on the complexity of constructing the activation graph.
Proposition 5
For any program \( \mathcal{P} \), the activation graph \( \Sigma(\mathcal{P}) \) can be constructed in time \( O(n_r^2 \times n_b \times (a_p \times a_f)^2) \), where \( n_r \) is the number of rules of \( \mathcal{P} \), \( n_b \) is the maximum number of body atoms in a rule, \( a_p \) is the maximum arity of predicate symbols and \( a_f \) is the maximum arity of function symbols.

Proof
To check whether a rule \( r_i \) activates a rule \( r_j \) we have to determine if an atom \( B \) in body(\( r_j \)) unifies with the head-atom \( A \) of \( r_i \). This can be done in time \( O(n_b \times u) \), where \( u \) is the cost of deciding whether two atoms unify, which is quadratic in the size of the two atoms (Venturini Zilli 1975), that is \( u = O((a_p \times a_f)^2) \) as the size of atoms is bounded by \( a_p \times a_f \) (recall that the maximum depth of terms is 1). In order to construct the activation graph we have to consider all pairs of rules and for each pair we have to check if the first rule activates the second one. Therefore, the global complexity is \( O(n_r^2 \times n_b \times u) = O(n_r^2 \times n_b \times (a_p \times a_f)^2) \).

We recall that given two atoms \( A \) and \( B \), the size of a mgu \( \theta \) for \( \{ A, B \} \) can be, in the worst case, exponential in the size of \( A \) and \( B \), but the complexity of deciding whether a unifier for \( A \) and \( B \) exists is quadratic in the size of \( A \) and \( B \) (Venturini Zilli 1975).

Proposition 6
The complexity of deciding whether a program \( \mathcal{P} \) is safe is \( O((\text{size}(\mathcal{P}))^2 + |\text{args}(\mathcal{P})|^3 \times |F_p|) \).

Proof
The construction of the activation graph \( \Sigma(\mathcal{P}) \) can be done in time \( O(n_r^2 \times n_b \times (a_p \times a_f)^2) \), where \( n_r \) is the number of rules of \( \mathcal{P} \), \( n_b \) is the maximum number of body atoms in a rule, \( a_p \) is the maximum arity of predicate symbols, and \( a_f \) is the maximum arity of function symbols (cf. Proposition 5).

The complexity of computing \( \Gamma A(\mathcal{P}) \) is bounded by \( O(|\text{args}(\mathcal{P})|^3 \times |F_p|) \) (cf. Theorem 2).

From Definition 7 and Proposition 4 it follows that the sequence \( \Gamma A(\mathcal{P}), \Psi(\Gamma A(\mathcal{P})), \Psi^2(\Gamma A(\mathcal{P})), \ldots \) is monotone and converges in a finite number of steps bounded by the cardinality of the set \( \text{args}(\mathcal{P}) \). The complexity of determining rules not depending on cycles in the activation graph \( \Sigma(\mathcal{P}) \) is bounded by \( O(n_r^2) \), as it can be done by means of a depth-first traversal of \( \Sigma(\mathcal{P}) \), which is linear in the number of its edges. Since checking whether the conditions of Definition 6 hold for all arguments in \( \mathcal{P} \) is in \( O(\text{size}(\mathcal{P})) \), checking such conditions for at most \( |\text{args}(\mathcal{P})| \) steps is \( O(|\text{args}(\mathcal{P})| \times \text{size}(\mathcal{P})) \). Thus, the complexity of checking all the conditions of Definition 7 for all steps is \( O(n_r^2 + |\text{args}(\mathcal{P})| \times \text{size}(\mathcal{P})) \).

Since, \( n_r^2 \times n_b \times (a_p \times a_f)^2 = O((\text{size}(\mathcal{P}))^2) \), \( |\text{args}(\mathcal{P})| = O(\text{size}(\mathcal{P})) \) and \( n_r^2 = O((\text{size}(\mathcal{P}))^2) \), the complexity of deciding whether \( \mathcal{P} \) is safe is \( O((\text{size}(\mathcal{P}))^2 + |\text{args}(\mathcal{P})|^3 \times |F_p|) \).
6 Bound queries and examples

In this section we consider the extension of our framework to queries. This is an
important aspect as in many cases, the answer to a query is finite, although the
models may have infinite cardinality. This happens very often when the query goal
contains ground terms.

6.1 Bound queries

Rewriting techniques, such as magic-set, allow bottom-up evaluators to efficiently
compute (partially) ground queries, that is queries whose query goal contains ground
terms. These techniques rewrite queries (consisting of a query goal and a program)
such that the top-down evaluation is emulated (Beeri and Ramakrishnan 1991;
Greco 2003; Greco et al. 2005; Alviano et al. 2010). Labelling techniques similar to
magic-set have been also studied in the context of term rewriting (Zantema 1995).

Before presenting the rewriting technique, let us introduce some notations.

A query is a pair \( Q = \langle q(u_1, \ldots, u_n), P \rangle \), where \( q(u_1, \ldots, u_n) \) is an atom called query
goal and \( P \) is a program. We recall that an adornment of a predicate symbol \( p \) with arity \( n \) is a string \( \alpha \in \{b, f\}^n \) such that \(|\alpha| = n^1\). The symbols \( b \) and \( f \) denote, respectively, bound and free arguments. Given a query \( Q = \langle q(u_1, \ldots, u_n), P \rangle \),
Magic\( S(Q) = \langle q^\alpha(u_1, \ldots, u_n), \text{Magic}\( S(q(u_1, \ldots, u_n), P) \rangle \) indicates the rewriting of \( Q \), where
Magic\( S(q(u_1, \ldots, u_n), P) \) denotes the rewriting of rules in \( P \) w.r.t. the query goal
\( q(u_1, \ldots, u_n) \) and \( \alpha \) is the adornment associated with the query goal.

We assume that our queries \( \langle G, P \rangle \) are positive, as the rewriting technique is here
applied to \( \langle G, \text{st}(P) \rangle \) to generate the positive program which is used to restrict the
source program (see Section 8).

Definition 9
A query \( Q = \langle G, P \rangle \) is safe if \( P \) or Magic\( S(G, P) \) is safe.

It is worth noting that it is possible to have a query \( Q = \langle G, P \rangle \) such that \( P \) is safe,
but the rewritten program Magic\( S(G, P) \) is not safe and vice versa.

Example 13
Consider the query \( Q = \langle p(f(f(a))), P_{13} \rangle \), where \( P_{13} \) is defined below:

\[
\begin{align*}
p(a), \\
p(f(X)) &\leftarrow p(X).
\end{align*}
\]

\( P_{13} \) is not safe, but if we rewrite the program using the magic-set method, we obtain
the safe program:

\[
\begin{align*}
\text{magic}_p^b(f(f(a))). \\
\text{magic}_p^b(X) &\leftarrow \text{magic}_p^b(f(X)). \\
p^b(a) &\leftarrow \text{magic}_p^b(a). \\
p^b(f(X)) &\leftarrow \text{magic}_p^b(f(X)), p^b(X).
\end{align*}
\]

1 Adornments of predicates, introduced to optimize the bottom-up computation of logic queries, are
similar to mode of usage defined in logic programming to describe how the arguments of a predicate \( p \)
must be restricted when an atom with predicate symbol \( p \) is called.
Consider now the query \( Q = \langle p(a), P'_{13} \rangle \), where \( P'_{13} \) is defined as follows:

\[
\begin{align*}
p(f(f(a))). \\
p(X) & \leftarrow p(f(X)).
\end{align*}
\]

The program is safe, but after the magic-set rewriting we obtain the following program:

\[
\begin{align*}
magic.p^b(a). \\
magic.p^b(f(X)) & \leftarrow magic.p^b(X). \\
p^b(f(f(a))) & \leftarrow magic.p^b(f(f(a))). \\
p^b(X) & \leftarrow magic.p^b(X), p^b(f(X)).
\end{align*}
\]

which is not recognized as safe because it is not terminating. 

Thus, we propose to first check if the input program is safe and, if it does not satisfy the safety criterion, to check the property on the rewritten program, which is query-equivalent to the original one.

We recall that for each predicate symbol \( p \) with arity \( n \), the number of adorned predicates \( p^{α_1...α_n} \) could be exponential and bounded by \( O(2^n) \). However, in practical cases only few adornments are generated for each predicate symbol. Indeed, rewriting techniques are well consolidated and widely used to compute bound queries.

### 6.2 Examples

Let us now consider the application of the technique described above to some practical examples. Since each predicate in the rewritten query has a unique adornment, we shall omit them.

**Example 14**

Consider the query \( \langle \text{reverse}([a, b, c, d], L), P_{14} \rangle \), where \( P_{14} \) is defined by the following rules:

\[
\begin{align*}
r_0 : & \quad \text{reverse}([], []). \\
r_1 : & \quad \text{reverse}([X|Y], [X|Z]) \leftarrow \text{reverse}(Y, Z).
\end{align*}
\]

The equivalent program \( P'_{14} \), rewritten to be computed by means of a bottom-up evaluator, is:

\[
\begin{align*}
\rho_0 : & \quad \text{m.reverse}([a, b, c, d]). \\
\rho_1 : & \quad \text{m.reverse}(Y) \leftarrow \text{m.reverse}([X|Y]). \\
\rho_2 : & \quad \text{reverse}([], []). \\
\rho_3 : & \quad \text{reverse}([X|Y], [X|Z]) \leftarrow \text{m.reverse}([X|Y]), \text{reverse}(Y, Z).
\end{align*}
\]

Observe that \( P'_{14} \) is not argument-restricted. In order to check \( Γ \)-acyclicity and safety criteria, we have to rewrite rule \( \rho_3 \) having complex terms in both the head and the body. Thus we add an additional predicate \( b_1 \) defined by rule \( \rho_4 \) and replace \( \rho_3 \) by \( \rho'_3 \).

\[
\begin{align*}
\rho'_3 : & \quad \text{reverse}([X|Y], [X|Z]) \leftarrow b_1(X, Y, Z). \\
\rho_4 : & \quad b_1(X, Y, Z) \leftarrow \text{m.reverse}([X|Y]), \text{reverse}(Y, Z).
\end{align*}
\]

The obtained program, denoted \( P''_{14} \), is safe but not \( Γ \)-acyclic.
Example 15

Consider the query \( \langle \text{length}([a, b, c, d], L), P_{15} \rangle \), where \( P_{15} \) is defined by the following rules:

\[ r_0 : \text{length}([], 0). \]
\[ r_1 : \text{length}([X|T], I + 1) \leftarrow \text{length}(T, I). \]

The equivalent program \( P'_{15} \), is rewritten to be computed by means of a bottom-up evaluator as follows:

\[ \rho_0 : \text{m}\text{.length}([a, b, c, d]). \]
\[ \rho_1 : \text{m}\text{.length}(T) \leftarrow \text{m}\text{.length}([X|T]). \]
\[ \rho_2 : \text{length}([], 0) \leftarrow \text{m}\text{.length}([]). \]
\[ \rho_3 : \text{length}([X|T], I + 1) \leftarrow \text{m}\text{.length}([X|T]), \text{length}(T, I). \]

Also in this case, it is necessary to split rule \( \rho_3 \) into two rules to avoid having function symbols in both the head and the body, as shown below:

\[ \rho'_3 : \text{length}([X|T], I + 1) \leftarrow \text{b1}(X, T, I). \]
\[ \rho_4 : \text{b1}(X, T, I) \leftarrow \text{m}\text{.length1}(X, T), \text{length}(T, I). \]

The obtained program, denoted \( P''_{15} \), is safe but not \( \Gamma \)-acyclic.

We conclude this section pointing out that the queries in the two examples above are not recognized as terminating by most of the previously proposed techniques, including AR. We also observe that many programs follow the structure of programs presented in the examples above. For instance, programs whose aim is the verification of a given property on the elements of a given list, have the following structure:

\[ \text{verify}([], []). \]
\[ \text{verify}([X|L_1], [Y|L_2]) \leftarrow \text{property}(X, Y), \text{verify}(L_1, L_2). \]

Consequently, queries having a ground argument in the query goal are terminating.

7 Further improvements

The safety criterion can be improved further as it is not able to detect that in the activation graph, there may be cyclic paths that are not effective or can only be activated a finite number of times. The next example shows a program which is finitely ground, but recognized as terminating by the safety criterion.

Example 16

Consider the following logic program \( P_{16} \) obtained from \( P_8 \) by adding an auxiliary predicate \( q \):

\[ r_1 : p(X, X) \leftarrow b(X). \]
\[ r_2 : q(f(X), g(X)) \leftarrow p(X, X). \]
\[ r_3 : p(X, Y) \leftarrow q(X, Y). \]

\( P_{16} \) is equivalent to \( P_8 \) w.r.t. predicate \( p \). \( \square \)

Observe that program \( P'_{15} \) is equivalent to program \( P_1 \) presented in the Introduction, assuming that the base predicate \( \text{input} \) is defined by a fact \( \text{input}([a, b, c, d]). \)
Although the activation graph $\Sigma(P)$ contains a cycle, the rules occurring in the cycle cannot be activated an infinite number of times. Therefore, in this section we introduce the notion of active paths and extend the definitions of activation graphs and safe programs.

**Definition 10 (Active path)**

Let $P$ be a program and $k \geq 1$ be a natural number. The path $(r_1, r_2, \ldots, r_k, r_{k+1})$ is an active path in the activation graph $\Sigma(P)$ iff there is a set of unifiers $\theta_1, \ldots, \theta_k$, such that

- head($r_1$) unifies with an atom from body($r_2$) with unifier $\theta_1$;
- head($r_i\theta_{i-1}$) unifies with an atom from body($r_{i+1}$) with unifier $\theta_i$ for $i \in [2, \ldots, k]$.

We write $r_1 \xrightarrow{k} r_{k+1}$ if there is an active path of length $k$ from $r_1$ to $r_{k+1}$ in $\Sigma(P)$. $\blacksquare$

Intuitively, $(r_1, r_2, \ldots, r_k, r_{k+1})$ is an active path if $r_1$ transitive activates rule $r_{k+1}$, that is if the head of $r_1$ unifies with some body atom of $r_2$ with mgu $\theta_1$, then the head of the rule $r_2\theta_1$ unifies with some body atom of $r_3$ with mgu $\theta_2$, then the head of the rule $r_3\theta_2$ unifies with some body atom of $r_4$ with mgu $\theta_3$, and so on until the head of the rule $r_k\theta_{k-1}$ unifies with some body atom of $r_{k+1}$ with mgu $\theta_k$.

**Definition 11 (k-restricted activation graph)**

Let $P$ be a program and $k \geq 1$ be a natural number, the $k$-restricted activation graph $\Sigma_k(P) = (\mathcal{P}, E)$ consists of a set of nodes denoting the rules of $P$ and a set of edges $E$ defined as follows: there is an edge $(r_i, r_j)$ from $r_i$ to $r_j$ iff $r_i \xrightarrow{k} r_j$, i.e. iff there is an active path of length $k$ from $r_i$ to $r_j$. $\blacksquare$

**Example 17**

The $k$-restricted activation graphs for the program of Example 16, with $k \in [1, \ldots, 3]$, are reported in Figure 7. $\blacksquare$

Obviously, the activation graph presented in Definition 5 is 1-restricted. We next extend the definition of safe function by referring to $k$-restricted activation graphs, instead of the (1-restricted) activation graph.

**Definition 12 (k-safety function)**

For any program $\mathcal{P}$ and natural number $k \geq 1$, let $A$ be a set of limited arguments of $\mathcal{P}$. The $k$-safety function $\Psi_k(A)$ denotes the set of arguments $q[i] \in \text{args}(\mathcal{P})$ such that for all rules $r = q(t_1, \ldots, t_m) \leftarrow \text{body}(r) \in \mathcal{P}$, either $r$ does not depend on a cycle $\pi$ of $\Sigma_j(\mathcal{P})$, for some $1 \leq j \leq k$, or $t_i$ is limited in $r$ w.r.t. $A$. $\blacksquare$
Observe that the \( k \)-safety function \( \Psi_k \) is defined as a natural extension of the safety function \( \Psi \) by considering all the \( j \)-restricted activation graphs, for \( 1 \leq j \leq k \).

Note that the 1-restricted activation graph coincides with the standard activation graph and, consequently, \( \Psi_1 \) coincides with \( \Psi \).

**Definition 13 (k-safe arguments)**

For any program \( \mathcal{P} \), \( \text{safe}_k(\mathcal{P}) = \Psi_k(\Gamma_A(\mathcal{P})) \) denotes the set of \( k \)-safe arguments of \( \mathcal{P} \). A program \( \mathcal{P} \) is said to be \( k \)-safe if all arguments are \( k \)-safe.

**Example 18**

Consider again the logic program \( P_{16} \) from Example 16. \( \Sigma_2(P_{16}) \) contains the unique cycle \((r_2, r_1)\); consequently, \( q[1] \) and \( q[2] \) appearing only in the head of rule \( r_2 \) are 2-safe. By applying iteratively operator \( \Psi_2 \) to the set of limited arguments \( \{b[1], q[1], q[2]\} \), we derive that also \( p[1] \) and \( p[2] \) are 2-safe. Since \( \text{safe}_2(P_{16}) = \text{args}(P_{16}) \), we have that \( P_{16} \) is 2-safe. Observe also that \( \Sigma_3(P_{16}) \) does not contain any edge and, therefore, all arguments are 3-safe.

For any natural number \( k > 0 \), \( \mathcal{I}_k \) denotes the class of \( k \)-safe logic programs, that is the set of programs \( \mathcal{P} \) such that \( \text{safe}_k(\mathcal{P}) = \text{args}(\mathcal{P}) \). The following proposition states that the classes of \( k \)-safe programs define a hierarchy where \( \mathcal{I}_k \subseteq \mathcal{I}_{k+1} \).

**Proposition 7**

The class \( \mathcal{I}_{k+1} \) of \((k + 1)\)-safe programs strictly extends the class \( \mathcal{I}_k \) of \( k \)-safe programs, for any \( k \geq 1 \).

**Proof**

\((\mathcal{I}_k \subseteq \mathcal{I}_{k+1})\) It follows straightforwardly from the definition of \( k \)-safe function.

\((\mathcal{I}_k \neq \mathcal{I}_{k+1})\) To show that the containment is strict, consider the program \( P_{16} \) from Example 16 for \( k = 1 \) and the following program \( \mathcal{P}_k \) for \( k > 1 \):

\[
\begin{align*}
    r_0 & : q_1(f(X), g(X)) \leftarrow p(X, X). \\
    r_1 & : q_2(X, Y) \leftarrow q_1(X, Y). \\
    & \vdots \\
    r_{k-1} & : q_k(X, Y) \leftarrow q_{k-1}(X, Y). \\
    r_k & : p(X, Y) \leftarrow q_k(X, Y).
\end{align*}
\]

It is easy to see that \( \mathcal{P}_k \) is in \( \mathcal{I}_{k+1} \), but not in \( \mathcal{I}_k \).

Recall that the minimal model of a standard program \( \mathcal{P} \) can be characterized in terms of the classical immediate consequence operator \( \mathcal{T}_\mathcal{P} \) defined as follows. Given a set \( I \) of ground atoms, then

\[
\mathcal{T}_\mathcal{P}(I) = \{A\theta \mid \exists r : A \leftarrow A_1, \ldots, A_n \in \mathcal{P} \text{ and } \exists \theta \text{ s.t. } A\theta \in I \text{ for every } 1 \leq i \leq n\}
\]

where \( \theta \) is a substitution replacing variables with constants. Thus, \( \mathcal{T}_\mathcal{P} \) takes as input a set of ground atoms and returns as output a set of ground atoms; clearly, \( \mathcal{T}_\mathcal{P} \) is monotonic. The \( i \)th iteration of \( \mathcal{T}_\mathcal{P} \) (\( i \geq 1 \)) is defined as follows: \( \mathcal{T}^1_\mathcal{P}(I) = \mathcal{T}_\mathcal{P}(I) \) and \( \mathcal{T}^{i+1}_\mathcal{P}(I) = \mathcal{T}_\mathcal{P}(\mathcal{T}^i_\mathcal{P}(I)) \) for \( i > 1 \). It is well known that the minimum model of \( \mathcal{P} \) is equal to the fixed point \( \mathcal{T}^\infty_\mathcal{P}(\theta) \).

A rule \( r \) is fired at run-time with a substitution \( \theta \) at step \( i \) if head\((r)\theta \in T^i_\mathcal{P}(\theta) - T^{i-1}_\mathcal{P}(\theta) \). Moreover, we say that \( r \) is fired (at run-time) by a rule \( s \) if \( r \) is fired with a
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1020 substitution $\theta$ at step $i$, $s$ is fired with a substitution $\sigma$ at step $i - 1$, and head($s$)$\sigma \in$
1021 body($r$)$\theta$. Let $\mathcal{P}$ be a program whose minimum model is $M = \mathcal{M}(\mathcal{P}) = T^\mathcal{M}_0(\emptyset)$,
1022 $M[r]$ denotes the set of facts which have been inferred during the application of
1023 the immediate consequence operator using rule $r$, that is the set of facts head($r$)$\theta$
1024 such that, for some natural number $i$, head($r$)$\theta \in T^\mathcal{M}_i(\emptyset) - T^\mathcal{M}_{i-1}(\emptyset)$; $M[r]$ if infinite
1025 iff $r$ is fired an infinite number of times. Clearly, if a rule $s$ fires at run-time a rule
1026 $r$, then the activation graph contains an edge $(s, r)$. An active sequence of rules is a
1027 sequence of rules $r_1, \ldots, r_n$ such that $r_i$ fires at run-time rule $r_{i+1}$ for $i \in [1, \ldots, n - 1]$.

1028 \textbf{Theorem 4}
1029 Let $\mathcal{P}$ be a logic program and let $r$ be a rule of $\mathcal{P}$. If $M[r]$ is infinite, then, for
1030 every natural number $k$, $r$ depends on a cycle of $\Sigma_k(\mathcal{P})$.
1031 \textbf{Proof}
1032 Let $n_r$ be the number of rules of $\mathcal{P}$ and let $N = n_r \times k$. If $M[r]$ is infinite we have
1033 that there is an active sequence of rules $r_0, r_1, \ldots, r_n$ such that $r_n$ coincides
1034 with $r$. This means that
1035 \[ r_0 \xrightarrow{k} r_k \xrightarrow{k} r'_k, \ldots, r_j \xrightarrow{k} r'_j, \ldots, r_{j=k} \xrightarrow{k} r'_{j=k} \xrightarrow{k} r''_{j=k} \xrightarrow{k} r_X, \]
1036 i.e. that the $k$-restricted activation graph $\Sigma_k(\mathcal{P})$ contains path $\pi = (r_0, r'_k)$,
1037 $(r'_k, r'_2), \ldots, (r'_j, r'_{j=k}), \ldots, (r'_X, r)$, Observe that the number of rules involved
1038 in $\pi$ is $n_r + 1$ and is greater than the number of rules of $\mathcal{P}$. Consequently, there is a
1039 rule occurring more than once in $\pi$, i.e. $\pi$ contains a cycle. Therefore, $r$ depends on
1040 a cycle of $\Sigma_k(\mathcal{P})$. \hfill $\square$
1041
1042 As shown in Example 18, in some cases the analysis of the $k$-restricted
1043 activation graph is enough to determine the termination of a program. Indeed,
1044 let cyclicR($\Sigma_k(\mathcal{P})$) be the set of rules $r$ in $\mathcal{P}$ s.t. $r$ depends on a cycle in $\Sigma_k(\mathcal{P})$, the
1045 following results hold.
1046
1047 \textbf{Corollary 2}
1048 A program $\mathcal{P}$ is terminating if $\forall r \in \mathcal{P}, \exists k$ s.t. $r \notin$ cyclicR($\Sigma_k(\mathcal{P})$).
1049 \textbf{Proof}
1050 Straightforward from Theorem 4. \hfill $\square$
1051
1052 Obviously, if there is a $k$ such that for all rules $r \in \mathcal{P}$ \( r \notin \text{cyclicR}(\Sigma_k(\mathcal{P})) \), $\mathcal{P}$ is
1053 terminating. We conclude this section showing that the improvements here discussed
1054 increase the complexity of the technique which is not polynomial anymore.
1055
1056 \textbf{Proposition 8}
1057 For any program $\mathcal{P}$ and natural number $k > 1$, the activation graph $\Sigma_k(\mathcal{P})$ can be
1058 constructed in time exponential in the size of $\mathcal{P}$ and $k$.
1059 \textbf{Proof}
1060 Let $(r_1, r_2), \ldots, (r_k, r_{k+1})$ be an active path of length $k$ in $\Sigma(\mathcal{P})$. Consider a pair $(r_i, r_{i+1})$
1061 and two unifying atoms $A_i = \text{head}(r_i)$ and $B_{i+1} = \text{body}(r_{i+1})$ (with $1 \leq i \leq k$), the
1062 size of an mgu $\theta$ for $A_i$ and $B_{i+1}$, represented in the standard way (cf. Section 2),
1063 can be exponential in the size of the two atoms. Clearly, the size of $A_i\theta$ and $B_{i+1}\theta$
can also be exponential. Consequently, the size of $A_{i+1} \theta$ which is used for the next step, can grow exponentially as well. Moreover, since in the computation of an active path of length $k$ we apply $k$ mgu’s, the size of terms can grow exponentially with $k$.

Observe that for the computation of the $1$-restricted argument graph it is sufficient to determine if two atoms unify (without computing the mgu), whereas for the computation of the $k$-restricted argument graphs, with $k > 1$, it is necessary to construct all the mgu’s and to apply them to atoms.

### 8 Computing stable models for disjunctive programs

In this section we discuss how termination criteria, defined for standard programs, can be applied to general disjunctive logic programs. First, observe that we have assumed that whenever the same variable $X$ appears in two terms occurring, respectively, in the head and body of a rule, at most one of the two terms is a complex term and that the nesting level of complex terms is at most one. There is no real restriction in such an assumption as every program could be rewritten into an equivalent program satisfying such a condition. For instance, a rule $r'$ of the form

$$p(f(g(X)), h(Y, Z)) \leftarrow p(f(X), Y), q(h(g(X), 1(Z)))$$

is rewritten into the set of ‘flatten’ rules below:

- $p(f(A), h(Y, Z)) \leftarrow b_1(A, Y, Z)$
- $b_1(g(X), Y, Z) \leftarrow b_2(X, Y, Z)$
- $b_2(X, Y, Z) \leftarrow b_3(X, Y, g(X), 1(Z))$
- $b_3(X, Y, B, C) \leftarrow p(f(X), Y), q(h(B, C))$

where $b_1$, $b_2$ and $b_3$ are new predicate symbols, whereas $A$, $B$ and $C$ are new variables introduced to flat terms with depth greater than 1.

More specifically, let $d(p(t_1, \ldots, t_n)) = \max\{d(t_1), \ldots, d(t_n)\}$ be the depth of atom $p(t_1, \ldots, t_n)$ and $d(A_1, \ldots, A_n) = \max\{d(A_1), \ldots, d(A_n)\}$ be the depth of a conjunction of atoms $A_1, \ldots, A_n$, for each standard rule $r$ we generate a set of ‘flatten’ rules, denoted by $\text{flat}(r)$ whose cardinality is bounded by $O(d(\text{head}(r)) + d(\text{body}(r)))$.

Therefore, given a standard program $\mathcal{P}$, the number of rules of the rewritten program is polynomial in the size of $\mathcal{P}$ and bounded by

$$O \left( \sum_{r \in \mathcal{P}} d(\text{head}(r)) + d(\text{body}(r)) \right) .$$

Concerning the number of arguments in the rewritten program, for a given rule $r$ we denote with $nl(r, h, i)$ (resp. $nl(r, b, i)$) the number of occurrences of function symbols occurring at the same nesting level $i$ in the head (resp. body) of $r$ and with $nf(r) = \max\{nl(r, t, i) | t \in \{h, b\} \wedge i > 1\}$. For instance, considering the above rule $r'$, we have that $nl(r', h, 1) = 2$ (function symbols $f$ and $h$ occur at nesting level 1 in the head), $nl(r', h, 2) = 1$ (function symbol $g$ occurs at nesting level 2 in the head), $nl(r', b, 1) = 2$ (function symbols $f$ and $h$ occur at nesting level 1 in the
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31

head), \(nl(r', b, 2) = 2\) (function symbols \(g\) and \(l\) occur at nesting level 2 in the head).

Consequently, \(nf(r') = 2\).

The rewriting of the source program results in a ‘flattened’ program with \(|\text{flat}(r)| - 1\) new predicate symbols. The arity of every new predicate in \(\text{flat}(r)\) is bounded by \(|\text{var}(r)| + nf(r)\). Therefore, the global number of arguments in the flattened program is bounded by

\[
O \left( \sum_{r \in \mathcal{P}} (|\text{head}(r)| + nf(r)) \right).
\]

The termination of a disjunctive program \(\mathcal{P}\) with negative literals can be determined by rewriting it into a standard logic program \(st(\mathcal{P})\) such that every stable model of \(\mathcal{P}\) is contained in the (unique) minimum model of \(st(\mathcal{P})\), and then by checking \(st(\mathcal{P})\) for termination.

**Definition 14 (Standard version)**

Given a program \(\mathcal{P}\), \(st(\mathcal{P})\) denotes the standard program, called standard version, obtained by replacing every disjunctive rule \(r = a_1 \vee \cdots \vee a_m \leftarrow \text{body}(r)\) with \(m\) standard rules of the form \(a_i \leftarrow \text{body}^+(r)\), for \(1 \leq i \leq m\).

Moreover, we denote with \(ST(\mathcal{P})\) the program derived from \(st(\mathcal{P})\) by replacing every derived predicate symbol \(q\) with a new derived predicate symbol \(Q\).

The number of rules in the standard program \(st(\mathcal{P})\) is equal to \(\sum_{r \in \mathcal{P}} |\text{head}(r)|\) where \(|\text{head}(r)|\) denotes the number of atoms in the head of \(r\).

**Example 19**

Consider program \(P_{19}\) consisting of the two rules

\[
\begin{align*}
p(X) \vee q(X) & \leftarrow r(X), \neg a(X) \\
r(X) & \leftarrow b(X), \neg q(X).
\end{align*}
\]

where \(p, q\) and \(r\) are derived (mutually recursive) predicates, whereas \(a\) and \(b\) are base predicates. The derived standard program \(st(P_{19})\) is as follows:

\[
\begin{align*}
p(X) & \leftarrow r(X) \\
q(X) & \leftarrow r(X) \\
r(X) & \leftarrow b(X).
\end{align*}
\]

**Lemma 2**

For every program \(\mathcal{P}\), every stable model \(M \in \mathcal{SM}(\mathcal{P})\) is contained in the minimum model \(\mathcal{M}.M(st(\mathcal{P}))\).

**Proof**

From the definition of stable models we have that every \(M \in \mathcal{SM}(\mathcal{P})\) is the minimal model of the ground positive program \(\mathcal{P}^M\). Consider now the standard program \(\mathcal{P}'\) derived from \(\mathcal{P}^M\) by replacing every ground disjunctive rule \(r = a_1 \vee \cdots \vee a_n \leftarrow \text{body}(r)\) with \(m\) ground normal rules \(a_i \leftarrow \text{body}(r)\). Clearly, \(M \leq \mathcal{M}.M(st(\mathcal{P}'))\).

Moreover, since \(\mathcal{P}' \subseteq st(\mathcal{P})\), we have that \(\mathcal{M}.M(st(\mathcal{P})) \leq \mathcal{M}.M(st(\mathcal{P}))\). Therefore, \(M \leq \mathcal{M}.M(st(\mathcal{P}))\).
The above lemma implies that for any logic program \( \mathcal{P} \), if \( \text{st}(\mathcal{P}) \) is finitely ground we can restrict the Herbrand base and only consider head (ground) atoms \( q(r) \) such that \( q(r) \in \mathcal{M}(\text{st}(\mathcal{P})) \). This means that, after having computed the minimum model of \( \text{st}(\mathcal{P}) \), we can derive a finite ground instantiation of \( \mathcal{P} \), equivalent to the original program, by considering only ground atoms contained in \( \mathcal{M}(\text{st}(\mathcal{P})) \).

We next show how the original program \( \mathcal{P} \) can be rewritten so that, after having computed \( \mathcal{M}(\text{st}(\mathcal{P})) \), every grounder tool easily generates an equivalent finitely ground program. The idea consists in generating, for any disjunctive program \( \mathcal{P} \), such that \( \text{st}(\mathcal{P}) \) satisfies some termination criterion (e.g. safety), a new equivalent program \( \text{ext}(\mathcal{P}) \). The computation of the stable models of \( \text{ext}(\mathcal{P}) \) can be carried out by considering the finite ground instantiation of \( \text{ext}(\mathcal{P}) \) (Leone et al. 2002; Simons et al. 2002; Gebser et al. 2007a).

For any disjunctive rule \( r = q_1(u_1) \lor \cdots \lor q_k(u_k) \leftarrow \text{body}(r) \), the conjunction of atoms \( Q_1(u_1), \ldots, Q_k(u_k) \) will be denoted by \( \text{head conj}(r) \).

**Definition 15** (Extended program)

Let \( \mathcal{P} \) be a disjunctive program and let \( r \) be a rule of \( \mathcal{P} \), then, \( \text{ext}(r) \) denotes the (disjunctive) extended rule \( \text{head}(r) \leftarrow \text{head conj}(r), \text{body}(r) \) obtained by extending the body of \( r \), whereas \( \text{ext}(\mathcal{P}) = \{ \text{ext}(r) \mid r \in \mathcal{P} \} \cup \text{ST}(\mathcal{P}) \) denotes the (disjunctive) extended program obtained by extending the rules of \( \mathcal{P} \) and adding (standard) rules defining the new predicates.

**Example 20**

Consider the program \( P_{19} \) of Example 19. The extended program \( \text{ext}(P_{19}) \) is as follows:

\[
\begin{align*}
\text{p}(X) \lor \text{q}(X) & \leftarrow \text{p}(X), \text{q}(X), \text{r}(X), \neg \text{a}(X). \\
\text{r}(X) & \leftarrow \text{r}(X), \text{b}(X), \neg \text{q}(X). \\
\text{p}(X) & \leftarrow \text{r}(X). \\
\text{q}(X) & \leftarrow \text{r}(X). \\
\text{r}(X) & \leftarrow \text{b}(X).
\end{align*}
\]

The following theorem states that \( \mathcal{P} \) and \( \text{ext}(\mathcal{P}) \) are equivalent w.r.t. the set of predicate symbols in \( \mathcal{P} \).

**Theorem 5**

For every program \( \mathcal{P} \), \( \mathcal{M}(\mathcal{P})[S_\mathcal{P}] = \mathcal{M}(\text{ext}(\mathcal{P}))[S_\mathcal{P}] \), where \( S_\mathcal{P} \) is the set of predicate symbols occurring in \( \mathcal{P} \).

**Proof**

First, we recall that \( \text{ST}(\mathcal{P}) \subseteq \text{ext}(\mathcal{P}) \) and assume that \( N \) is the minimum model of \( \text{ST}(\mathcal{P}) \), i.e. \( N = \mathcal{M}(\text{ST}(\mathcal{P})) \).

- We first show that for each \( S \in \mathcal{M}(\text{ext}(\mathcal{P})) \), \( M = S - N \) is a stable model for \( \mathcal{P} \), that is \( M \in \mathcal{M}(\mathcal{P}) \).
- Let us consider the ground program \( \mathcal{P}'' \) obtained from \( \text{ext}(\mathcal{P})[S] \) by first deleting every ground rule \( r = \text{head}(r) \leftarrow \text{head conj}(r), \text{body}(r) \) such that \( N \not\models \text{head conj}(r) \) and then by removing from the remaining rules, the conjunction \( \text{head conj}(r) \). Observe that the sets of minimal models for \( \text{ext}(\mathcal{P})[S] \)
and \( P'' \) coincide, i.e. \( \mathcal{M} \mathcal{M}(\text{ext}(P))^\mathcal{S} = \mathcal{M} \mathcal{M}(P'') \). Indeed, for every \( r \) in \( \text{ext}(P)^\mathcal{S} \), if \( N \not\models \text{headconj}(r) \), then the body of \( r \) is false and thus \( r \) can be removed as it does not contribute to infer head atoms. On the other hand, if \( N \models \text{headconj}(r) \), the conjunction \( \text{headconj}(r) \) is trivially true, and can be safely deleted from the body of \( r \).

Therefore, \( M \cup N \in \mathcal{M} \mathcal{M}(P'') \). Moreover, since \( P'' = (P \cup \text{ST}(P))^\mathcal{S} \), we have that \( M \in \mathcal{M} \mathcal{M}(P(M)) \), that is \( M \in \mathcal{I} \mathcal{M}(P) \).

\[ \square \]

\section{Conclusion}

In this paper we have proposed a new approach for checking, on the basis of structural properties, termination of the bottom-up evaluation of logic programs with function symbols. We have first proposed a technique, called \( \Gamma \)-acyclicity, extending the class of argument-restricted programs by analyzing the propagation of complex terms among arguments using an extended version of the argument graph. Next, we have proposed a further extension, called safety, which also analyzes how rules can activate each other (using the activation graph) and how the presence of some arguments in a rule limits its activation. We have also studied the application of the techniques to partially ground queries and have proposed further improvements which generalize the safety criterion through the introduction of a hierarchy of classes of terminating programs, called \( k \)-safety, where each \( k \)-safe class strictly includes the \((k-1)\)-safe class.

Although our results have been defined for standard programs, we have shown that they can also be applied to disjunctive programs with negative literals, by simply rewriting the source programs. The semantics of the rewritten program is ‘equivalent’ to the semantics of the source one and can be computed by current answer set systems. Even though our framework refers to the model theoretic semantics, we believe that the results presented here go beyond the ASP community and could be of interest also for the (tabled) logic programming community (e.g. tabled Prolog community).

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References


