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**SEMANTIC FORGETTING IN ANSWER SET  
PROGRAMMING**

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INFSYS RESEARCH REPORT 1843-07-08  
NOVEMBER 2007

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## SEMANTIC FORGETTING IN ANSWER SET PROGRAMMING

Thomas Eiter<sup>1</sup> and Kewen Wang<sup>2</sup>

**Abstract.** The notion of forgetting, also known as variable elimination, has been investigated extensively in the context of classical logic, but less so in (nonmonotonic) logic programming and nonmonotonic reasoning. The few approaches that exist are based on syntactic modifications of a program at hand. In this paper, we establish a declarative theory of forgetting for disjunctive logic programs under answer set semantics that is fully based on semantic grounds. The suitability of this theory is justified by a number of desirable properties. In particular, one of our results shows that our notion of forgetting can be entirely captured by classical forgetting. We present several algorithms for computing a representation of the result of forgetting, and provide a characterization of the computational complexity of reasoning from a logic program under forgetting. As applications of our approach, we present a fairly general framework for resolving conflicts in inconsistent knowledge bases that are represented by disjunctive logic programs, and we show how the semantics of inheritance logic programs and update logic programs from the literature can be characterized through forgetting. The basic idea of the conflict resolution framework is to weaken the preferences of each agent by forgetting certain knowledge that causes inconsistency. In particular, we show how to use the notion of forgetting to provide an elegant solution for preference elicitation in disjunctive logic programming.

**Keywords:** answer set programming, nonmonotonic logic programs, knowledge representation, forgetting, computational complexity.

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**Acknowledgements:** This work has been partially supported by the Austrian Science Funds (FWF) Projects P17212 and P18019, the EC project REVERSE (IST-2003-506779), the Australia Research Council (ARC) Discovery Projects DP0666107 and DP0666540.

Preliminary versions of this paper with some of the results have been presented at AAAI 2006 and at NMR 2006.

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# 1 Introduction

For intelligent agents, the ability to discard irrelevant information has been recognized as an important feature (that is mastered well by humans) and received broad attention in artificial intelligence, both from a cognitive and a computational perspective. In the area of knowledge representation, this ability is often referred to as *forgetting* [Lin and Reiter, 1994] or *variable elimination* [Brown, 2003], but has been studied under many different names including irrelevance, independence, irredundancy, novelty, or separability (see [Subramanian *et al.*, 1997; Lang *et al.*, 2003] for more details).

Forgetting has its root in Boolean Algebra [Boole, 1847] where it is a fundamental reasoning process. C. I. Lewis [1918] has pointed out that, for purposes of application of Boolean logic to commonsense reasoning, the elimination/forgetting is a process more important than solution<sup>1</sup> since most processes of reasoning take place through the elimination of “middle” variables. Boole writes of such middle variables that it “usually happens in commonsense reasoning, and especially when we have more than one premises, that some of the elements [in the premises] are not required to appear in the conclusion.”

Forgetting and its applications have been investigated extensively in the context of classical logic, for example, [Bledsoe and Hines, 1980; Larrosa, 2000; Larrosa *et al.*, 2005; Lin and Reiter, 1994; Lang *et al.*, 2003; Middeldorp *et al.*, 1996; Moinard, 2007; Weber, 1986], but less so in nonmonotonic logic programming and reasoning. In this context, it was first considered in [Zhang *et al.*, 2005; Zhang and Foo, 2006], where two types of forgetting—strong and weak forgetting—have been defined by first transforming a logic program  $P$  into a reduced form and then deleting some rules (and literals) from it. While this approach works well in a number of cases, it has two major drawbacks. First, its semantic underpinning is not fully clear. Specifically, the relationship between the intended semantics of a logic program, in terms of its answer sets, and the result of the syntactic transformations that are carried out by strong and weak forgetting is unclear. Second, this approach does not address desirable properties for a reasonable notion of forgetting in nonmonotonic logic programming. In particular, one may ask what is the difference of these notions of forgetting from traditional approaches to deletion of rules/literals in logic programming and databases.

A further aspect is that both strong and weak forgetting are syntax-sensitive, i.e., programs that are semantically equivalent may have different results after forgetting about the same literal. For example, the programs  $P = \{p \leftarrow . q \leftarrow not\ p\}$  and  $Q = \{p \leftarrow\}$  are equivalent under the answer set semantics. Weak forgetting about  $p$  from  $P$  yields the program  $WForgetLP(P, p) = \{q \leftarrow\}$  and from  $Q$  the program  $WForgetLP(Q, p) = \{\}$ ; clearly, these programs are not equivalent.

While the role of syntax in logic programming is well-acknowledged, one might argue that relative to the semantics of this syntax, equivalent programs should behave in the same way. In particular, in this example the result of forgetting about  $p$  in  $P$  and  $Q$  should yield semantically the same result (note that, under answer set semantics, the second rule in  $P$  is redundant).

A similar phenomenon can be observed for strong forgetting. Consider  $P = \{q \leftarrow not\ p. q \leftarrow not\ q\}$  and  $Q = \{q \leftarrow\}$ . Then these two programs are equivalent under the answer set semantics. However, the results of strong forgetting about  $p$  from  $P$  and  $Q$  are  $SForgetLP(P, p) = \{q \leftarrow not\ q\}$  and  $SForgetLP(Q, p) = \{q \leftarrow\}$ , respectively, which are obviously not equivalent. The discrepancy is here even more noticeable: the result of strong forgetting about an atom from a consistent program can be inconsistent.

Thus, an alternative notion of forgetting for nonmonotonic logic programming is highly desirable. In

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<sup>1</sup>In [Lewis, 1918] a problem is formulated as a Boolean equation such that a solution of the Boolean equation corresponds to a solution of the given problem. In particular, solving a Boolean equation is treated as a process of eliminating/forgetting variables that represent unknowns.

this paper, we choose answer set programming (ASP) [Lifschitz, 2002] as the underlying nonmonotonic logic. ASP is a new paradigm of logic programming under the answer set semantics [Gelfond and Lifschitz, 1990], which is becoming a major tool for knowledge representation and reasoning due to its simplicity, expressive power, and connection to major nonmonotonic logics. A number of efficient ASP solvers, such as DLV, Smodels, ASSAT, Cmodels, or Clasp are available (see [Asparagus homepage, Since 2005]), which can handle large problem instances.

Prior to defining a notion of forgetting for nonmonotonic logic programming, we may pose the question what desirable properties a reasonable theory of forgetting should have. The following ones appear to be natural candidates for us. Let  $P$  be a logic program and let  $P'$  be the result of forgetting about a literal  $l$  in  $P$ .

- (F1) The proposed notion of forgetting should be a “natural” generalization of, and relate to, forgetting in classical logic.
- (F2) No new symbols are introduced in  $P'$ , i.e., the vocabulary stays the same.
- (F3) The reasoning under  $P'$  is equivalent to the reasoning under  $P$  if  $l$  is ignored.
- (F4) The result of forgetting is not sensitive to syntax in that the results of forgetting about  $l$  in semantically equivalent programs should also be semantically equivalent.
- (F5) The semantic notion of forgetting is coupled with a syntactic counterpart, i.e., there is effective constructible syntax for representing the result of forgetting.

(F1) specifies the major intuition behind forgetting and clarifies the difference of forgetting from deletion. (F2) is necessary because the forgetting is to eliminate redundant symbols. This is a difference between forgetting and some approaches to revision, update, and deletion, such as [Alferes *et al.*, 1998; Buccafurri *et al.*, 1999; Eiter *et al.*, 2002; Grant *et al.*, 1993; Waal and Gallagher, 1993]; note that to combine forgetting with other approaches to adding new information is a different issue. (F3) provides a semantic justification for the forgetting. Note that  $P'$  and  $P$  may have different answer sets in general (see Proposition 1); (F4) guarantees that the notion of forgetting is semantically well-defined. Finally, (F5) is useful for applications of forgetting in knowledge representation.

To the best of our knowledge, *there is no theory of forgetting in nonmonotonic reasoning or logic programming which is based on the above criteria*. However, the definition of forgetting in classical logic cannot be directly adapted to logic programming (cf. Section 3.1). The main contributions of the present paper are as follows.

- We establish a declarative, semantically defined notion of forgetting for disjunctive logic programs under answer set semantics called *semantic forgetting*. The suitability of semantic forgetting is justified by a number of desirable properties, including the ones given above.
- As one of them, we show that our notion of forgetting naturally captures classical forgetting. As we show, this can be exploited for reasoning under forgetting about a literal from a logic program by resorting to representations of a nonmonotonic logic program in terms of classical logic [Lin and Zhao, 2002; Lin and Zhao, 2004; Lee and Lifschitz, 2003].
- As another such property, for every consistent disjunctive program  $P$  and literal  $l$ , a syntactic representation  $\text{forget}(P, l)$  for forgetting about  $l$  in  $P$  in terms of a nonmonotonic logic program always

exists. Besides two semantics-based algorithms for computing such a representation, we also present a transformation-based algorithm. This algorithm allows to obtain the result of forgetting about a literal  $l$  in  $P$  via a series of program transformations and other rewritings.

- In connection with these algorithms, we characterize the computational complexity of the major reasoning tasks from logic programs under forgetting about a literal. As it turns out, model checking and credulous reasoning under forgetting about a literal from a logic program are more complex than in the standard setting (by one level in the polynomial hierarchy), while skeptical reasoning has the same complexity. These results provide useful insights into feasible representations of forgetting, and suggest that a polynomial-size result of forgetting may not always be feasible. This (and stronger results) can be established by applying the theory of Cadoli et al. [2000; 2002].
- As an application of our approach, we present a fairly general framework for resolving conflicts in inconsistent knowledge bases. The basic idea of this framework is to weaken the preferences of each agent by forgetting certain knowledge that causes inconsistency. In particular, we show how to use the notion of forgetting to provide an elegant solution for preference elicitation in ASP.
- Furthermore, we show that inheritance programs [Buccafurri et al., 1999], update programs [Eiter et al., 2000; Eiter et al., 2002] and fragments of dynamic programs [Alferes et al., 1998; Alferes et al., 2000] can be characterized in terms of the semantic forgetting.

While in this paper, we focus on nonmonotonic logic programs, the basic ideas underlying our approach to semantic forgetting may be applied to other well-known formalisms of nonmonotonic reasoning, such as default logic [Reiter, 1980] or autoepistemic logic [Moore, 1985], as well, of which nonmonotonic logic programs under answer set semantics can be seen as particular fragments. In fact, these formalisms extend classical logic, and a notion of forgetting that complies with classical forgetting (which is based on semantics) seems needed there. Our results thus also provide a benchmark for approaches to forgetting in other formalisms of nonmonotonic reasoning, which remain to be developed.

The rest of the paper is organized as follows. Section 2 briefly recalls some basics of disjunctive logic programs and the answer sets. Section 3 defines the notion of forgetting in ASP, shows some important properties, and relates it to classical forgetting and independence [Lang et al., 2003]. Thereafter, Section 4 presents algorithms for computing the result of forgetting in ASP, while Section 5 studies some complexity issues. Section 6 then presents some applications, namely to conflict resolution in multi-agent systems, to inheritance logic programs, and to logic program updates. The final Section 7 concludes the work.

## 2 Preliminaries

We briefly review some basic definitions and notation in answer set programming that will be used throughout this paper.

A *disjunctive logic program* (simply, *logic program*) is a finite set of rules of the form

$$a_1 \vee \dots \vee a_s \leftarrow b_1, \dots, b_m, \text{not } c_1, \dots, \text{not } c_n, \quad (1)$$

where  $s, m, n \geq 0$ , and all  $a_i, b_j$ , and  $c_k$  are from a set *Lit* of classical literals in a propositional language<sup>2</sup>. We assume here that all  $a_i$  are pairwise distinct, and similarly all  $b_j$  and all  $c_k$ . A *literal* is a *positive literal*  $p$

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<sup>2</sup>The results of this paper may be lifted to the predicate case as usual.

or a *negative literal*  $\neg p$  for some atom  $p$ . For an atom  $p$ ,  $p$  and  $\neg p$  are called *complementary*. For any literal  $l$ , its complementary literal is denoted by  $\neg.l$ .

Given a rule  $r$  of form (1),  $head(r) = a_1 \vee \dots \vee a_s$  and  $body(r) = body^+(r) \cup not\ body^-(r)$  where  $body^+(r) = \{b_1, \dots, b_m\}$ ,  $body^-(r) = \{c_1, \dots, c_n\}$ , and  $not\ body^-(r) = \{not\ q \mid q \in body^-(r)\}$ . Occasionally, in abuse of notation we view  $head(r)$  also as set  $\{a_1, \dots, a_s\}$ .

A rule  $r$  of the form (1) is *normal* or *non-disjunctive*, if  $s \leq 1$ ; *positive*, if  $n = 0$ ; *negative*, if  $m = 0$ ; *constraint*, if  $s = 0$ ; *fact*, if  $m = 0$  and  $n = 0$ . The rule with  $s = n = m = 0$  is the constant *false*. A logic program  $P$  is called *normal* (resp. *positive*, *negative*), if every rule in  $P$  is normal (resp. positive, negative).

We denote by  $Lit_P \subseteq Lit$  the *literal base* of logic program  $P$ , that is, the set of all literals occurring in  $P$ . Unless stated otherwise or clear from the context,  $Lit$  will be implicitly given by  $Lit_P$ . An *interpretation* is a set of literals  $X \subseteq Lit$  that contains no pair of complementary literals. A disjunction  $a_1 \vee \dots \vee a_s$  is satisfied by  $X$ , denoted  $X \models a_1 \vee \dots \vee a_s$  if  $a_i \in X$  for some  $i$  with  $1 \leq i \leq s$ . A rule  $r$  is satisfied by  $X$ , denoted  $X \models r$ , if  $X \models head(r)$  whenever  $body^+(r) \subseteq X$  and  $body^-(r) \cap X = \emptyset$  hold. Furthermore,  $X$  is a model of  $P$ , denoted  $X \models P$ , if  $X \models r$  for every rule  $r \in P$ .

The semantics of a logic program  $P$  is defined in terms of its *answer sets* [Gelfond and Lifschitz, 1991] as follows. Given an interpretation  $X$ , the *reduct* of  $P$  on  $X$  is defined as  $P^X = \{head(r) \leftarrow body^+(r) \mid r \in P, body^-(r) \cap X = \emptyset\}$ . Then  $X$  is an *answer set* of  $P$ , if  $X$  is a minimal model of  $P^X$ . By  $AS(P)$  we denote the collection of all answer sets of  $P$ .

A logic program  $P$  may have zero, one or multiple answer sets.  $P$  is *consistent*, if it has at least one answer set. It is well-known that the answer sets of a logic program  $P$  are incomparable: for any  $X$  and  $X'$  in  $AS(P)$ ,  $X \subseteq X'$  implies  $X = X'$ .

**Example 1** Let  $P$  be the logic program consisting the following rules:

$$\begin{aligned} a \vee b &\leftarrow not\ c \\ d &\leftarrow a \\ d &\leftarrow b \end{aligned}$$

Then  $P$  has two answer sets  $X_1 = \{a, d\}$  and  $X_2 = \{b, d\}$ . Obviously,  $X_1$  and  $X_2$  are incomparable.

Two logic programs  $P$  and  $P'$  are *equivalent*, denoted  $P \equiv P'$ , if  $AS(P) = AS(P')$ , i.e.,  $P$  and  $P'$  have the same answer sets.

By  $P \models_s l$  and  $P \models_c l$  we denote skeptical and credulous consequence of a literal  $l$  from a logic program  $P$ , respectively; that is,  $P \models_s l$  iff  $l \in S$  for every  $S \in AS(P)$  and  $P \models_c l$  iff  $l \in S$  for some  $S \in AS(P)$ .

### 3 Forgetting in Logic Programming

In this section, we define what it means to forget about a literal  $l$  in a logic program  $P$ . The idea is to obtain a logic program which does not contain  $l$  and is equivalent to the original logic program if we ignore the existence of the literal  $l$ . We believe that forgetting should go beyond syntactic removal of rules/literals and be close to classical forgetting and answer set semantics (keeping its spirit) at the same time. Thus, the definition of forgetting in this section is given in semantics terms, i.e., based on answer sets.



### 3.1 Definition of Semantic Forgetting

In classical propositional logic, the result of forgetting  $\text{forget}(T, p)$  about a proposition  $p$  in a finite theory  $T$  is conveniently defined as  $T(p/\text{true}) \vee T(p/\text{false})$ , where  $T(p/\text{true})$  and  $T(p/\text{false})$  are obtained by taking the conjunction of all sentences in  $T$  and replacing all occurrences of  $p$  with  $\text{true}$  and  $\text{false}$ , respectively. This method cannot be directly generalized to logic programming, since there is no notion of the “disjunction” of two logic programs. However, if we look at forgetting in from a model-theoretic perspective, then we can obtain the models of  $\text{forget}(T, p)$  as follows: compute first all (2-valued) models of  $T$  and then remove  $p$  from each model if it contains  $p$ . The resulting collection of sets  $\{M \setminus \{p\} \mid M \models T\}$  is exactly the set of all models of  $\text{forget}(T, p)$ .

Similarly, given a consistent logic program  $P$  and a literal  $l$ , we could naively define the result of forgetting about  $l$  in  $P$  as a logic program  $P'$  whose answer sets are exactly  $\mathcal{AS}(P) \setminus l = \{X \setminus \{l\} \mid X \in \mathcal{AS}(P)\}$ . However, this notion of forgetting cannot guarantee the existence of  $P'$  for even simple programs. For example, consider  $P = \{a \leftarrow . p \vee q \leftarrow\}$ . Here  $\mathcal{AS}(P) = \{\{a, p\}, \{a, q\}\}$  and thus  $\mathcal{AS}(P) \setminus p = \{\{a\}, \{a, q\}\}$ . Since  $\{a\} \subset \{a, q\}$  and, as well-known, answer sets are incomparable under set inclusion,  $\mathcal{AS}(P) \setminus p$  cannot be the set of answer sets of any logic program.

A solution to this problem is a *suitable notion of minimal answer set* such that the definition of answer sets, minimality, and forgetting can be fruitfully combined. To this end, we call a set  $X'$  an *l-subset of a set*  $X$ , denoted  $X' \subseteq_l X$ , if  $X' \setminus \{l\} \subseteq X \setminus \{l\}$ . Similarly, a set  $X'$  is a *strict l-subset of*  $X$ , denoted  $X' \subset_l X$ , if  $X' \setminus \{l\} \subset X \setminus \{l\}$ . Two sets  $X$  and  $X'$  of literals are *l-equivalent*, denoted  $X \sim_l X'$ , if  $X' \subseteq_l X$  and  $X \subseteq_l X'$ .

**Definition 1 (*l-Answer Set*)** Let  $P$  be a consistent logic program, let  $l$  be a literal in  $\text{Lit}_P$ , and let  $X \subseteq \text{Lit}_P$  be a set of literals.

1. For a collection  $\mathcal{S}$  of sets of literals,  $X \in \mathcal{S}$  is *l-minimal* if there is no  $X' \in \mathcal{S}$  such that  $X' \subset_l X$ . By  $\text{min}_l(\mathcal{S})$  we denote the collection of all *l-minimal* elements in  $\mathcal{S}$ .
2. An answer set  $X$  of logic program  $P$  is an *l-answer set* if  $X$  is *l-minimal* in  $\mathcal{AS}(P)$ . By  $\mathcal{AS}_l(P)$  we denote the set of all *l-answer sets* of  $P$ .

For example,  $P = \{a \leftarrow . p \vee q \leftarrow\}$  has two answer sets  $X = \{a, p\}$  and  $X' = \{a, q\}$ .  $X$  is a *p-answer set* of  $P$ , but  $X'$  is not. This example shows that, for a logic program  $P$  and a literal  $l$ , not every answer set is an *l-answer set*.

The sets in  $\mathcal{AS}_l(P) \setminus l = \{X \setminus \{l\} \mid X \in \mathcal{AS}_l(P)\}$  are incomparable, and so we can find a logic program which has this collection as its answer sets. Note that to achieve incomparability, one could select other answer sets than those which are minimal in  $\mathcal{AS}(P) \setminus l$  (e.g., the maximal ones). However, selecting minimal answer sets is in line with the guiding principle of logic programming and nonmonotonic reasoning to minimize positive information.

Note that in the above definition,  $P$  was assumed to be consistent (i.e.,  $\mathcal{AS}(P) \neq \emptyset$ ). If a logic program is inconsistent, the result of forgetting seems not to be clear, since the possibility of removing inconsistency from the logic program might have to be considered. For example, a logic program  $P$  may have partial stable models [Sacca and Zaniolo, 1991] while it is inconsistent under the answer set semantics (i.e. the stable model semantics). Forgetting from inconsistent programs is an interesting issue, but we do not consider it here. In the rest of this paper, we always assume that  $P$  is a consistent logic program.

The following proposition collects some easy properties of *l-answer sets*.

**Proposition 1** For every consistent program  $P$  and every literal  $l$  in  $Lit_P$ , the following holds:

1. Every  $l$ -answer set  $X$  of  $P$  is an answer set of  $P$ .
2. For every answer set  $X$  of  $P$ , there exists an  $l$ -answer set  $X'$  of  $P$  such that  $X' \subseteq_l X$ .
3. Every answer set  $X$  of  $P$  with  $l \in X$  is an  $l$ -answer set of  $P$ .
4. If an answer set  $X$  of  $P$  is not an  $l$ -answer set of  $P$ , then  $l \notin X$  and there exists an  $l$ -answer set  $Y$  of  $P$  such that  $l \in Y$  and  $Y \subset_l X$ .
5. If  $l \notin Lit_P$ , then  $X$  is an  $l$ -answer set of  $P$  iff  $X$  is an answer set of  $P$ .

Having the notion of minimality about forgetting a literal, we are now in a position to define the result of forgetting about a literal in a logic program.

**Definition 2 (Semantic Forgetting)** Let  $P$  be a consistent logic program and  $l$  be a literal. A logic program  $P'$  represents the result of forgetting about  $l$  in  $P$ , if

1.  $B_{P'} \subseteq B_P \setminus \{l\}$ , i.e.,  $l$  does not occur in  $P'$ , and
2.  $\mathcal{AS}(P') = \mathcal{AS}_l(P) \setminus l$ , i.e., for every set  $X'$  of literals such that  $l \notin X'$ ,  $X'$  is an answer set of  $P'$  iff there exists an  $l$ -answer set  $X$  of  $P$  such that  $X' \sim_l X$ .

We use  $\text{forget}(P, l)$  as a generic notation for a logic program representing the result of forgetting about  $l$  in  $P$ .

An important difference of the notion of forgetting here from existing approaches to updating and merging logic programs, cf. [Alferes *et al.*, 1998; Buccafurri *et al.*, 1999; Eiter *et al.*, 2002; Grant *et al.*, 1993; Waal and Gallagher, 1993], is that merely  $l$  and possibly some other literals are removed. However, no new symbols are introduced in  $P'$ .

For a consistent logic program  $P$ , some program  $P'$  as in the above definition always exists (cf. Algorithm 2 for details). However, different such programs  $P'$  exist. It follows from the definition that they are all equivalent under answer set semantics.

**Proposition 2** Let  $P$  be a consistent logic program and let  $l \in Lit_P$  be a literal. If  $P'$  and  $P''$  are two results for forgetting about  $l$  in  $P$ , then  $P'$  and  $P''$  are equivalent.

Before further properties of forgetting are explored in Section 3.2, let us look at some example programs.

**Example 2** 1. If  $P_1 = \{q \leftarrow \text{not } p\}$ , then  $\text{forget}(P_1, q) = \emptyset$  and  $\text{forget}(P_1, p) = \{q \leftarrow\}$ . Here,  $\text{forget}(P_1, p)$  is obtained from  $P_1$  by removing  $\text{not } p$  in the rule  $q \leftarrow \text{not } p$  while  $\text{forget}(P_1, q)$  is obtained by removing the whole rule  $q \leftarrow \text{not } p$ .

2. If  $P_2 = \{q \leftarrow \text{not } p, p \leftarrow \text{not } q\}$ , then  $\text{forget}(P_2, p) = \emptyset$ . The reason is that  $P_2$  has two answer sets  $\{p\}$  and  $\{q\}$  but only  $\{p\}$  is a  $p$ -answer set of  $P_2$ . Thus  $\text{forget}(P_2, p)$  has the single answer set  $\{\}$ . Similarly,  $\text{forget}(P_2, q) = \{\}$ .

3. Consider  $P_3 = \{q \leftarrow \text{not } p. p \leftarrow\}$ , which has the single answer  $\{p\}$ . Thus  $\text{forget}(P_3, p) = \{\}$  rather than  $\{q \leftarrow\}$ . This is intuitive, because we are forgetting all impacts of  $p$  on  $P_3$ . In particular, “forgetting about  $p$ ” is different from “assuming not  $p$ ”.

From the above examples, one might guess that some program  $\text{forget}(P, p)$  can be obtained by simply removing rules and/or literals with head  $p$  and/or positive body literal  $p$ , and by removing not  $p$  in the remaining rules. However, as the next example shows, this is not true in general.

4. Let  $P_4 = \{a \leftarrow \text{not } b. b \leftarrow \text{not } a. p \leftarrow \text{not } a. c \leftarrow \text{not } p\}$ . According to [Zhang et al., 2005], the result of weak forgetting about  $p$  in  $P_4$  is the program  $\text{WForgetLP}(P_4, p) = \{a \leftarrow \text{not } b. b \leftarrow \text{not } a. c \leftarrow\}$ , while the result of strong forgetting about  $p$  in  $P_4$  is the program  $\text{SForgetLP}(P_4, p) = \{a \leftarrow \text{not } b. b \leftarrow \text{not } a\}$ . Neither is fully intuitive:  $c$  depends on  $a$  (by means of double negation), but in both  $\text{WForgetLP}(P_4, p)$  and  $\text{SForgetLP}(P_4, p)$  any connection between  $c$  and  $a$  is lost. In contrast,  $\text{forget}(P_4, p) = \{a \leftarrow \text{not } b. b \leftarrow \text{not } a. c \leftarrow a\}$  where the connection between  $c$  and  $a$  is maintained via the rule  $c \leftarrow a$ .
5.  $P_5 = \{p \vee q \leftarrow \text{not } p. c \leftarrow q\}$  has the single answer set  $\{q, c\}$ . Here,  $p$  is involved in unstratified negation and must be false in every answer set. We obtain  $\text{forget}(P_5, p) = \{q \leftarrow . c \leftarrow\}$ , which intuitively results by pushing through the only possible value for  $p$  and simplifying the program.
6. Let  $P_6 = \{a \vee p \leftarrow \text{not } b. c \leftarrow \text{not } p. b \leftarrow\}$ . This program has the single answer set  $\{b, c\}$ , and no atom  $p$  is involved in cyclic negation. Forgetting about  $p$  in  $P_6$  yields  $\text{forget}(P_6, p) = \{c \leftarrow . b \leftarrow\}$ , which again corresponds to a simplified version of the program  $P_6$  the results after pushing through the value of  $p$ ; note that the first rule in  $P_6$  is never applicable.

We will discuss how to obtain a concrete program  $\text{forget}(P, l)$  in the next section.

### 3.2 Basic Properties of Forgetting

In this subsection, we present some further properties of forgetting. First, the number of answer sets can never increase.

**Proposition 3** *Let  $P$  be a consistent logic program. Then, for every literal  $l$  in  $\text{Lit}_P$ , it holds that  $|\mathcal{AS}(\text{forget}(P, l))| \leq |\mathcal{AS}(P)|$ .*

This is a simple consequence of the fact that only some, but not all answer sets of  $P$  are  $l$ -answer sets. Note that this property is compliant with the principle of closing the world, and eliminating possibilities in favor of a default case.

The following proposition generalizes Proposition 2.

**Proposition 4** *Let  $P$  and  $P'$  be two consistent logic programs and  $l$  a literal in  $P$ . If  $P$  and  $P'$  are equivalent, then  $\text{forget}(P, l)$  and  $\text{forget}(P', l)$  are also equivalent.*

However, forgetting here does not preserve some special equivalences of logic programs stronger than ordinary equivalence like strong equivalence [Lifschitz et al., 1999] or uniform equivalence [Eiter and Fink, 2003].

We say that forgetting *preserves* an equivalence  $\equiv_X$  on logic programs, if for every logic programs  $Q$  and  $Q'$  and for every literal  $l$ ,  $Q \equiv_X Q'$  implies  $\text{forget}(Q, l) \equiv_X \text{forget}(Q', l)$ . Here  $X$  can be strong

equivalence, uniform equivalence, or any other equivalence relation on the collection of (disjunctive) logic programs.

An equivalence relation  $\equiv_X$  for logic programs on  $Lit$  is *invariant* under literal extensions, if the following holds: whenever  $P$  and  $P'$  are programs such that  $Lit_P, Lit_{P'} \subseteq Lit$  and  $l \notin Lit$  is a new literal, then  $P \equiv_X P'$  w.r.t.  $Lit$  iff  $P \equiv_X P'$  w.r.t.  $Lit \cup \{l\}$ .

An equivalence  $\equiv_X$  is *stronger* than ordinary equivalence  $\equiv$  if the following conditions are satisfied:

1. For any programs  $P$  and  $P'$ ,  $P \equiv_X P'$  implies  $P \equiv P'$ .
2. There exist two programs  $P$  and  $P'$  such that  $P \equiv P'$  and  $P \not\equiv_X P'$ .

**Proposition 5** *Let  $\equiv_X$  be a equivalence relation on a collection of logic programs on  $Lit$  that is stronger than ordinary equivalence and invariant under literal extensions. Then forgetting does not preserve  $\equiv_X$ .*

Both strong and uniform equivalence [Lifschitz *et al.*, 2001; Eiter and Fink, 2003] are clearly stronger than ordinary equivalence, and clearly also invariant under literal extensions. Hence, none of them is preserved by the definition of forgetting introduced. This, however, is a consequence of the freedom to arbitrarily instantiate the generic program  $\text{forget}(P, l)$ . For specific realizations of  $\text{forget}(P, l)$ , both strong and uniform equivalence may be preserved under forgetting; for example, the realizations  $\text{forget}_1(P, l)$  and  $\text{forget}_2(P, l)$  in Section 4 have this property. A suitable notion of forgetting which preserves strong equivalence is interesting for some applications, but beyond the scope of this paper.

**Proposition 6** *For every consistent program  $P$  and every literal  $l$  in  $Lit_P$ , the following holds:*

1.  $\mathcal{AS}(\text{forget}(P, l)) = \{X \setminus \{l\} \mid X \in \mathcal{AS}_l(P)\}$ .
2. If  $X \in \mathcal{AS}_l(P)$  with  $l \notin X$ , then  $X \in \mathcal{AS}(\text{forget}(P, l))$ .
3. For every  $X \in \mathcal{AS}(P)$  such that  $l \in X$ ,  $X \setminus \{l\} \in \mathcal{AS}(\text{forget}(P, l))$ .
4. For every  $X' \in \mathcal{AS}(\text{forget}(P, l))$ , either  $X'$  or  $X' \cup \{l\}$  is in  $\mathcal{AS}(P)$ .
5. For every  $X \in \mathcal{AS}(P)$ , there exists  $X' \in \mathcal{AS}(\text{forget}(P, l))$  such that  $X' \subseteq X$ .
6. If  $l$  does not appear in  $P$ , then  $\text{forget}(P, l) \equiv P$ .

The next proposition says that, after forgetting about a literal in a logic program, the resulting program is equivalent to the original one under skeptical reasoning, but weaker under credulous reasoning (i.e., inferences are lost).

**Proposition 7** *Let  $P$  be a consistent logic program and let  $l, l'$  be literals in  $Lit_P$  such that  $l' \neq l$ . Then,*

1.  $P \models_s l'$  iff  $\text{forget}(P, l) \models_s l'$ , and
2.  $P \models_c l'$  if  $\text{forget}(P, l) \models_c l'$ .

The above definition of forgetting about a single literal  $l$  in a logic program  $P$  can be straightforwardly extended to a set  $F$  of literals. We can similarly define  $X_1 \subseteq_F X_2$ ,  $X_1 \sim_F X_2$ , and  $F$ -answer sets of a logic program, and the properties of forgetting about a single literal can be generalized to this setting. Furthermore, the result of forgetting about a set  $F$  can be obtained by forgetting the literals in  $F$  one by one.

**Proposition 8** *Let  $P$  be a consistent logic program and let  $F = \{l_1, \dots, l_m\}$  be a set of literals. Then  $\text{forget}(P, F) \equiv \text{forget}(\text{forget}(\text{forget}(P, l_1), l_2), \dots), l_m)$ .*

Notice that the particular indexing of the literals in  $F$  does not matter. This result, which allows to reduce forgetting to an base operation, is quite useful, but its proof (which is given in Appendix) requires some technicalities.

We remark that for removing a proposition  $p$  entirely from a program  $P$ , it is suggestive to remove both the literals  $p$  and  $\neg p$  in  $P$  (i.e., all positive and negative information about  $p$ ). This can be easily accomplished by  $\text{forget}(P, \{p, \neg p\})$ .

Let us consider a simple logic program which contains a pair of complementary literals.

**Example 3** *Let  $P$  be the following logic program:*

$$\begin{aligned} \text{flies}(\text{Tweety}) &\leftarrow \text{pigeon}(\text{Tweety}). \\ \neg \text{flies}(\text{Tweety}) &\leftarrow \text{penguin}(\text{Tweety}). \\ \text{pigeon}(\text{Tweety}) \vee \text{penguin}(\text{Tweety}) &\leftarrow . \end{aligned}$$

*If we forget about only one of  $\text{flies}$  and  $\neg \text{flies}$ , then the complementary literal must still be in the result of forgetting. For instance, the program  $\text{forget}(P, \text{flies})$  still contains  $\neg \text{flies}$ :*

$$\begin{aligned} \neg \text{flies}(\text{Tweety}) &\leftarrow \text{not penguin}(\text{Tweety}). \\ \text{pigeon}(\text{Tweety}) \vee \text{penguin}(\text{Tweety}) &\leftarrow . \end{aligned}$$

*However,  $\text{forget}(\text{forget}(P, \text{flies}), \neg \text{flies})$  and  $\text{forget}(\text{forget}(P, \neg \text{flies}), \text{flies})$  are by Proposition 8 equivalent and amount to*

$$\text{pigeon}(\text{Tweety}) \vee \text{penguin}(\text{Tweety}) \leftarrow .$$

### 3.3 Relation to Classical Forgetting

We now consider the relationship between classical forgetting  $\text{forget}(T, p)$  and logic programming forgetting  $\text{forget}(P, p)$ . Besides the stable and the answer set semantics [Gelfond and Lifschitz, 1988b; Gelfond and Lifschitz, 1991], another influential semantics for nonmonotonic logic programming is the Clarke's completion semantics [Clark, 1978], which defines the semantics of logic programs in terms of classical logic. It is well-known that answer set and completion semantics are different in general. For example, the logic program  $P = \{p \leftarrow q; q \leftarrow p\}$  has a unique answer set  $\{\}$ . However, Clarke's completion for  $P$  gives  $\{p \equiv q\}$ , which has two models  $\{\}$  and  $\{p, q\}$ . Lin and Zhao [2002; 2004] showed that the answer set semantics for a logic program can be characterized by a simple extension of Clarke's program completion by adding so called *loop formulas*. They consider normal logic programs that may contain constraints. This approach allows to compute the answer sets of a normal logic program using a classical SAT solver. Lee and Lifschitz [2003] extended this characterization to the class of disjunctive logic programs. For simplicity, we assume in this subsection that programs have no strong negation, which can be compiled away in the standard way as usual.

For every logic program  $P$ , its *completion*  $\text{comp}(P)$  is the set of propositional formulas containing

- $\text{body}(r) \rightarrow \text{head}(r)$  for every rule  $r$  in  $P$ ,
- and the formula  $a \rightarrow \bigvee_{r \in P, a \in \text{head}(r)} (\text{body}(r) \wedge \bigwedge_{p \in (\text{head}(r) - \{a\})} \neg p)$ , for every atom  $a$ .

Here  $head(r) - L$  is, for any the set of atoms  $L$ , the head set of atoms that occur in  $head(r)$  but not in  $L$ , and “not ” and comma “;” in rule bodies are translated into the negation “ $\neg$ ” and conjunction “ $\wedge$ ” in classical propositional logic, respectively. An empty head  $head(r)$  is translated into  $\perp$ .

Given a logic program  $P$ , its *positive dependency graph*  $G_P$  is the directed graph whose vertices are the atoms occurring in  $P$  and where there is an edge from  $p$  to  $q$  iff there exists some rule  $r$  in  $P$  such that  $p \in head(r)$  and  $q \in body^+(r)$ . A nonempty set  $L = \{p_1, \dots, p_k\}$  of atoms is a *loop* of  $P$ , if for every distinct  $p_i, p_j \in L$  there exists a path of nonzero length from  $p_i$  to  $p_j$  in  $G_P$  such that all vertices in this path belong to  $L$ . The *conjunctive loop formula* for  $L$  is

$$CLF(L) = (p_1 \wedge \dots \wedge p_k) \rightarrow (\bigvee_{\phi \in R(L)} \phi)$$

where  $R(L)$  is the set of formulas  $body(r) \wedge \bigwedge_{p \in head(r) - L} \neg p$  for all rules  $r \in P$  with  $head(r) \cap L \neq \emptyset$  and  $body^+(r) \cap L = \emptyset$ .

For example,  $L = \{p, q\}$  is a loop of the logic program  $P = \{r \vee p \leftarrow q; q \leftarrow p\}$ . In this example,  $R(L) = \{\}$  and thus the right hand side of  $CLF(L)$  is an empty disjunction. Thus  $CLF(L)$  is the propositional formula  $(p \wedge q) \rightarrow \perp$ .

Let  $lcomp(P) = comp(P) \cup CLF(P)$ , where  $CLF(P)$  is the set of all loop formulas. A fundamental result established by Lin and Zhao [2002] shows that the generalized completion  $lcomp(P)$  exactly characterizes the answer set semantics. This was extended to disjunctive logic programs by Lee and Lifschitz [2003] as follows.

**Theorem 1 ([Lee and Lifschitz, 2003])** *Let  $P$  be a (disjunctive) logic program and let  $X \subseteq Lit_P$  be a set of atoms. Then  $X$  is an answer set of  $P$  iff  $X$  is a model of  $lcomp(P)$ .*

Since  $L = \{p, q\}$  is the only loop of  $P = \{r \vee p \leftarrow q; q \leftarrow p\}$ , we have  $CLF(P) = \{(p \wedge q) \rightarrow \perp\}$ . Since we have  $comp(P) = \{r \rightarrow q \wedge \neg q, p \rightarrow q \wedge \neg r, q \rightarrow p\}$ , we obtain (after some simplifications) that  $lcomp(P) = \{p \equiv q, r \rightarrow \perp, p \rightarrow \neg r, q \rightarrow r, (p \wedge q) \rightarrow \perp\}$ . This theory has exactly one model, namely  $\{\}$ , which is the unique answer set of  $P$ .

It is easy to see that  $lcomp(P) = comp(P)$  if  $P$  is negative. Hence,

**Corollary 2** *Let  $Q$  be a negative program (without strong negation) and let  $X \subseteq Lit_P$  be a set of literals. Then  $X$  is an answer set of  $Q$  if and only if  $X$  is a model of  $comp(Q)$ .*

This corollary is a special case of some previous results in [Ben-Eliyahu and Dechter, 1994; Erdem and Lifschitz, 2003].

Since for a logic program  $P$  and an atom  $p$ , the two classical theories  $lcomp(\text{forget}(P, p))$  and  $\text{forget}(lcomp(P), p)$  can be formed, where in the former logic programming forgetting is applied to  $P$  and in the latter classical forgetting to the theory  $lcomp(P)$ , the natural question is how these two theories are related. Intuitively, the models of the first theory are all incomparable, while the models of the second theory may be not.

For example, let  $P = \{p \leftarrow \text{not } q. q \leftarrow \text{not } p\}$ . Then  $lcomp(\text{forget}(P, p)) = \{\neg q\}$ , which has the single model  $\{\}$ , while  $\text{forget}(lcomp(P), p) = \{\top \leftrightarrow \neg q\} \vee \{F \leftrightarrow \neg q\} \equiv \top$ , which has two comparable models  $\{q\}$  and  $\emptyset$ . However, the *minimal models* of  $\text{forget}(lcomp(P), p)$  are the same as the *models* of  $lcomp(\text{forget}(P, p))$ . In fact, this holds in general.

**Theorem 3** *Let  $P$  be a consistent (disjunctive) logic program and let  $p \in Lit_P$  be an atom. Then  $X \subseteq Lit_P$  is an answer set of  $\text{forget}(P, p)$  iff  $X$  is a minimal model of  $\text{forget}(lcomp(P), p)$ . That is,*

$$AS(\text{forget}(P, p)) = MMod(\text{forget}(lcomp(P), p)),$$

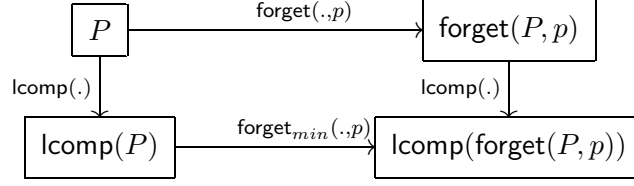


Figure 1: Commuting logic program completion and forgetting

where  $MMod(T)$  denotes the set of all minimal models (w.r.t.  $\subseteq$ ) of a theory  $T$  in classical logic.

The proof of this theorem is given in the Appendix. This result means that  $\text{forget}(P, p)$  can be characterized by forgetting in classical logic. If we use  $\text{forget}_{min}(T, p)$  to denote a set of classical formulas whose models are the minimal models of the classical forgetting  $\text{forget}(T, p)$ , then the equation in Theorem 3 can be reformulated as

$$\text{lcomp}(\text{forget}(P, p)) \equiv \text{forget}_{min}(\text{lcomp}(P), p),$$

where  $\equiv$  denotes classical equivalence. This result is graphically represented in the commutative diagram in Figure 1. It is quite useful, since it implies that one can “bypass” the use of a logic programming engine entirely and represent the answer sets of  $\text{forget}(P, p)$  in the frameworks of circumscription and closed world reasoning [McCarthy, 1980; Lifschitz, 1994; Gelfond and Lifschitz, 1988a]. This can be done by applying circumscription to  $\text{lcomp}(P)$ , which we explain in more detail.

In circumscription, minimality is understood as the impossibility of making, in the context of predicate logic, the extent of the predicates  $p_1, \dots, p_k$  which are circumscribed in a theory  $T$  smaller without changing the extent of the other predicates. Furthermore, some predicates  $z_1, \dots, z_l$  among them may be allowed to vary in the process of minimizing  $p_1, \dots, p_k$ , which is needed in many applications. As for our concern of propositional logic, all  $p_i$  and  $z_j$  are atoms and  $\text{Circ}(T; \vec{p}, \vec{z})$  is a propositional formula with quantifiers on atoms which semantically captures the circumscription of a finite theory  $T$  with respect to  $\vec{p} = (p_1, \dots, p_k)$  and  $\vec{z} = (z_1, \dots, z_l)$ . As well known,  $\text{Circ}(T; \vec{p}, \vec{z})$  is logically equivalent to Gelfond et al.’s extended closed world assumption  $\text{ECWA}(T; \vec{p}, \vec{z})$  [Gelfond and Lifschitz, 1988a], which augments  $T$  by additional formulas.

Let us write  $\text{Circ}(T, z)$  respectively  $\text{ECWA}(T, z)$  for the case where  $\vec{z}$  contains a single atom  $z$  and  $\vec{p}$  all other atoms. Then,  $M$  is a model of  $\text{Circ}(T; z)$  respectively  $\text{ECWA}(T; z)$ , exactly if  $M$  is a  $z$ -minimal model of  $T$  in the sense that  $M' \subseteq_z M$  implies  $M' \sim_z M$  for every model  $M'$  of  $T$ . We have the following result.

**Theorem 4** *Let  $P$  be a consistent (disjunctive) logic program and let  $p \in \text{Lit}_P$  be an atom. Then  $X \subseteq \text{Lit}_P \setminus \{p\}$  is an answer set of  $\text{forget}(P, p)$  if and only if either  $X$  or  $X \cup \{p\}$  is a model of  $\text{Circ}(\text{lcomp}(P), p)$  (resp.,  $\text{ECWA}(\text{lcomp}(P), p)$ ).*

(A proof is given in the appendix.) By this rather intuitive result, we may exploit also circumscription engines for reasoning from  $\text{forget}(P, p)$ , into which we feed the loop completion  $\text{lcomp}(P)$  of  $P$ . Note that the latter is exploited by some ASP solvers (viz. ASSAT, Cmodels) as a stepping stone to compute answer sets, and that classical forgetting of  $p$  from  $\text{lcomp}(P)$  can be performed efficiently.

### 3.4 Forgetting vs. Independence

As it is argued in [Lang *et al.*, 2003], the notion of *independence* is important in automated deduction, query answering, and belief revision. For example, an intelligent agent must possess the ability of determining and discarding irrelevant information efficiently. When reasoning is involved, the issue of independence (or irrelevance) becomes more delicate and complex. The idea of forgetting about independent literals can be useful in improving reasoning procedures.

Informally, if a literal  $l$  is independent of a logic program  $P$ , then the answer set semantics of  $P$  should be unchanged if we forget about  $l$  from  $P$ . So, it is natural and reasonable to formally define the notion of *semantic independence* as follows.

**Definition 3** *Let  $P$  be a consistent logic program. A literal  $l \in Lit_P$  is semantically independent of  $P$ , if  $AS(\text{forget}(P, l)) = AS(P)$ .*

Obviously, a literal  $l$  which is semantically independent of  $P$  can be safely “forgotten.”

**Example 4** *Consider the program  $P = \{p \leftarrow \text{not } q. s \leftarrow s\}$ . Then,  $\text{forget}(P, q) = \{p \leftarrow . s \leftarrow s\}$  and  $\text{forget}(P, s) = \{p \leftarrow \text{not } q\}$ . Clearly,  $P$ ,  $\text{forget}(P, q)$ , and  $\text{forget}(P, s)$  all have the single answer set  $\{p\}$ . Hence, both  $s$  and  $q$  are semantically independent of  $P$ .*

The following proposition provides an intuitive characterization for semantic independence.

**Proposition 9** *Let  $P$  be a logic program. A literal  $l \in Lit_P$  is semantically independent of  $P$  if and only if  $l \notin S$  for every answer set  $S$  of  $P$ .*

This result, which is straightforward from Definition 2, is intuitive: a literal is independent of  $P$  if and only if  $l$  is false with respect to every answer set of  $P$ , that is,  $P \not\models_c l$ . In some cases, semantic independence of literals can be verified syntactically, as shown by the next result. Let  $\mathbf{T}^*$  be the program transformations introduced in Section 4.3.1 and let  $\mathbf{T}^*(P)$  be the respective canonical form of program  $P$ .

**Proposition 10** *Let  $P$  be a logic program and let  $l \in Lit_P$  be a literal. Suppose  $P'$  is any negative program obtained from  $P$  by transformations from  $\mathbf{T}^*$  such that  $l$  does not occur in  $P'$ . Then  $l$  is semantically independent of  $P$ . Furthermore, such a program  $P'$  exists iff  $l$  does not occur in  $\mathbf{T}^*(P)$ .*

The converse of Proposition 10 is not true in general. For example, consider the program  $P = \{p \leftarrow \text{not } q. q \leftarrow \text{not } p. q \leftarrow p.\}$ . Then  $p$  is semantically independent of  $P$ . On the other hand,  $p$  occurs in every program  $P'$  resulting from  $P$  by (repeated) transformations in  $\mathbf{T}^*$ , and in particular in the canonical form  $\mathbf{T}^*(P) = \{p \leftarrow \text{not } q. q \leftarrow \text{not } p. q \leftarrow \text{not } q.\}$ . It remains as an interesting issue whether there is a set of program transformations that is strong enough to syntactically characterize the notion of semantic independence of literals.

## 4 Computation of Forgetting

As we have noted,  $\text{forget}(P, l)$  exists for any consistent logic program  $P$  and literal  $l$ . In this section, we discuss some issues on computing the result of forgetting.



**Algorithm**  $\text{forget}_1(P, l)$ *Input:* Consistent (disjunctive) logic program  $P$  and a literal  $l$  in  $P$ .*Output:* A normal logic program representing  $\text{forget}(P, l)$ .*Method:**Step 1.* Compute  $\mathcal{AS}(P)$  using an ASP solver (e.g., DLV or gnT).*Step 2.* Remove  $l$  from every set in  $\mathcal{AS}(P)$  and denote the resulting collection as  $\mathcal{A}'$ .*Step 3.* Obtain  $\mathcal{A}''$  by removing all non-minimal sets from  $\mathcal{A}'$ .*Step 4.* Construct  $P'$  whose answer sets are exactly  $\mathcal{A}'' = \{A_1, \dots, A_m\}$ :

- For each  $A_i$ , let  $P_i = \{l' \leftarrow \text{not } \bar{A}_i \mid l' \in A_i\}$ , where  $\bar{A}_i = \text{Lit}_P \setminus A_i$ .
- Let  $P' = P_1 \cup \dots \cup P_m$ .

*Step 5.* Output  $P'$  as  $\text{forget}(P, l)$ .Figure 2: Algorithm  $\text{forget}_1(P, l)$ 

## 4.1 Naive Algorithm

By Definition 2, we can easily obtain a naive algorithm for computing  $\text{forget}(P, l)$  using an ASP solver for logic programs, like DLV [Leone *et al.*, 2006] or GnT [Janhunen *et al.*, 2000], which is shown in Figure 2.

It is well-known that any collection  $\mathcal{S}$  of sets of consistent literals which are pairwise incomparable, can be represented by some logic program  $P$  such that  $\mathcal{AS}(P) = \mathcal{S}$ . In fact, such  $P$  can be constructed from  $\mathcal{S}$  in polynomial time.

Algorithm  $\text{forget}_1(P, l)$  is sound and complete w.r.t. forgetting as in Definition 2.

**Theorem 5** *Given a consistent (disjunctive) logic program  $P$  and a literal  $l$ , Algorithm  $\text{forget}_1(P, l)$  outputs a correct representation of  $\text{forget}(P, l)$ .*

The above Step 2 may return collect many answer sets in  $\mathcal{A}'$  that are not minimal; in fact, it is easy to find examples where  $\mathcal{A}'$  contains exponentially many sets but only few of them are minimal. For example, let  $P = \{p \vee q \leftarrow \cdot a_i \vee b_i \leftarrow q, 1 \leq i \leq n\}$ . Then  $P$  has the single  $p$ -minimal answer set  $\{p\}$ , but exponentially many other answer sets  $\{q, l_1, \dots, l_n\}$ , where  $l_i$  is either  $a_i$  or  $b_i$ , which all lead to non-minimal sets in  $\mathcal{A}'$ .

To avoid this problem, we present an improved version of Algorithm  $\text{forget}_1(P, l)$  in the next subsection.

## 4.2 Improved Algorithm

An improved version of Algorithm  $\text{forget}(P, l)$ , which is shown in Figure 3, pushes the task of minimality checking for candidate  $l$ -answer sets of  $P$  into constraint satisfiability of an augmented program. It exploits in this way the constraint solving capabilities offered by some ASP solvers.

Since the answer sets of  $P'$  are exactly  $\mathcal{A}$ , Algorithm  $\text{forget}_2(P, l)$  is sound and complete w.r.t. semantic forgetting.

**Theorem 6** *For every consistent (disjunctive) logic program  $P$  and a literal  $l$ , Algorithm  $\text{forget}_2(P, l)$  outputs a correct representation of  $\text{forget}(P, l)$ .*

The advantage of Algorithm  $\text{forget}_2(P, L)$  is that the strategy in the Steps 1 and 2 makes  $l$ -minimization obsolete and thus no blowup into a large intermediate result with respect to the number of answer sets as

**Algorithm**  $\text{forget}_2(P, l)$ 

*Input:* Consistent (disjunctive) logic program  $P$  and a literal  $l$  in  $P$ .

*Output:* A normal logic program representing  $\text{forget}(P, l)$ .

*Method:*

*Step 1.* Let  $P_1 = P \cup \{ \leftarrow \text{not } l \}$ .

- Compute the answer sets of  $P_1$  (e.g., using DLV or gnT).
- Remove  $l$  from each obtained answer set. Let  $\mathcal{A}_1$  be the resulting collection of sets. (Each set in  $\mathcal{A}_1$  is an answer set of forgetting about  $l$  from  $P$ .)

*Step 2.* Let  $P_2 = P \cup \{ \leftarrow l \} \cup \{ \leftarrow a_1, \dots, a_k \mid \{a_1, \dots, a_k\} \in \mathcal{A}_1 \}$ .

- Compute the answer sets of  $P_2$  (e.g., using DLV or gnT), and let  $\mathcal{A}_2$  be the result.
- Set then  $\mathcal{A} = \mathcal{A}_1 \cup \mathcal{A}_2$ .

*Step 3.* Construct  $P'$  whose answer sets are exactly  $\mathcal{A} = \{A_1, \dots, A_m\}$ :

- For each  $A_i$ , let  $P_i = \{l' \leftarrow \text{not } \bar{A}_i \mid l' \in A_i\}$ , where  $\bar{A}_i = \text{Lit}_P \setminus A_i$ .
- Let  $P' = P_1 \cup \dots \cup P_n$ .

*Step 4.* Output  $P'$  as  $\text{forget}(P, l)$ .

Figure 3: Improvement to Algorithm  $\text{forget}_1(P, l)$

in Algorithm  $\text{forget}_1(P, l)$  can happen. Still, however, the resulting program  $P'$  may be large compared to a small program representing  $\text{forget}(P, l)$ , which might be constructed from  $P$  by other means. In the next subsection, we discuss how to construct a representation of  $\text{forget}(P, l)$  in a more syntactic manner by program transformations.

However, the “semantic” constructions by  $\text{forget}_1(P, l)$  and  $\text{forget}_2(P, l)$  also have an advantage: they clearly preserve the equivalence of logic programs under any notion of equivalence  $\equiv_X$  between logic programs that is stronger than ordinary equivalence, and thus in particular under strong and weak equivalence. This is a simple consequence of the fact that for all programs  $P$  which are ordinarily equivalent the output of  $\text{forget}_1(P, l)$  is the same, and similarly for  $\text{forget}_2(P, l)$ .

### 4.3 Transformation-Based Algorithm

The algorithm  $\text{forget}_1(P, l)$  and  $\text{forget}_2(P, l)$  are based on the semantic view of forgetting, and do not aim at computing the result of forgetting in a more syntax-oriented manner, by modifying the rules in  $P$ . In this subsection, we present an algorithm  $\text{forget}_3(P, l)$  of this kind that is based on program transformations. This algorithm outputs, differently from  $\text{forget}_1(P, l)$  and  $\text{forget}_2(P, l)$ , not always ordinary logic programs but sometime logic programs in which some literals are under double negation as failure. They inherit their semantics from the more general class of nested logic programs [Lifschitz *et al.*, 1999].

Before presenting algorithm  $\text{forget}_3(P, l)$ , we need further preliminaries on program transformations and programs with double negation as failure.

### 4.3.1 Basic program transformations

We start with recalling program transformations that were discussed in [Brass and Dix, 1999; Wang and Zhou, 2005]. More precisely, we consider the following collection  $\mathbf{T}^*$  of program transformations:

**Elimination of Tautologies**  $P'$  is obtained from  $P$  by the elimination of tautologies if there is a rule  $r$ :  $head(r) \leftarrow body^+(r), not\ body^-(r)$  in  $P$  such that  $head(r) \cap body^+(r) \neq \emptyset$  and  $P' = P \setminus \{r\}$ .

**Example 5** Let  $P_1$  consist of the following rules:

$$\begin{aligned} r_1 : & \quad p \vee p_1 \leftarrow not\ p_1 \\ r_2 : & \quad p \leftarrow p, not\ q_1 \\ r_3 : & \quad p \leftarrow p_1, not\ q_1 \\ r_4 : & \quad p_1 \leftarrow not\ q, not\ p_2 \\ r_5 : & \quad q_1 \leftarrow p_2, not\ q \\ r_6 : & \quad p \vee q_1 \leftarrow \\ r_7 : & \quad p_1 \vee p_3 \leftarrow \\ r_8 : & \quad p_3 \leftarrow p, not\ p. \end{aligned}$$

Then  $r_2$  is a tautology and thus  $P_2 = \{r_1, r_3, r_4, r_5, r_6, r_7, r_8\}$  can be obtained from  $P$  by the elimination of tautologies.

**Elimination of Head Redundancy**  $P'$  is obtained from  $P$  by the elimination of head redundancy if there is a rule  $r$  in  $P$  such that literal  $l$  is in both  $head(r)$  and  $body^-(r)$  and  $P' = P \setminus \{r\} \cup \{(head(r) - l) \leftarrow body(r)\}$ . Here  $head(r) - l$  is the disjunction obtained by removing  $l$  from  $head(r)$ .

By the elimination of head redundancy,  $r_1$  is simplified into  $r'_1 : p \leftarrow not\ p_1$  and thus  $P_2$  is transformed into  $P_3 = \{r'_1, r_3, r_4, r_5, r_6, r_7, r_8\}$ .

The above two transformations guarantee that those rules whose head and body have common literals are removed.

**Positive Reduction**  $P'$  is obtained from  $P$  by positive reduction if there is a rule  $r$ :  $head(r) \leftarrow body^+(r), not\ body^-(r)$  in  $P$  and  $c \in body^-(r)$  such that  $c \notin head(P)$  and  $P'$  is obtained from  $P$  by removing  $not\ c$  from  $r$ . That is,  $P' = P \setminus \{r\} \cup \{head(r) \leftarrow body^+(r), not\ (body^-(r) \setminus \{c\})\}$ . Here  $head(P) = \cup_{r \in P} head(r)$ .

$P_4 = \{r'_1, r_3, r'_4, r'_5, r_6, r_7, r_8\}$  is obtained from  $P_3 = \{r'_1, r_3, r_4, r_5, r_6, r_7, r_8\}$  by the positive reduction, where  $r'_4$  is the rule  $p_1 \leftarrow$  and  $r'_5$  is  $q_1 \leftarrow p_2$ .

**Negative Reduction**  $P'$  is obtained from  $P$  by negative reduction if there are two rules  $r$ :  $head(r) \leftarrow body^+(r), not\ body^-(r)$  and  $r'$ :  $head(r') \leftarrow$  in  $P$  such that  $head(r') \subseteq body^-(r)$  and  $P' = P \setminus \{r\}$ .

In our example,  $P_5 = \{r_3, r'_4, r'_5, r_6, r_7, r_8\}$  is obtained from  $P_4$  by negative reduction, where  $r = r'_1$  and  $r' = r'_4$ .

For defining the next program transformation, we need the notion of *implications* for rules defined in [Brass and Dix, 1999]. We say  $r'$  is an implication of  $r$  if  $head(r) \subseteq head(r')$ ,  $body(r) \subseteq body(r')$  and at least one of the inclusions is strict.

**Elimination of Implications**  $P'$  is obtained from  $P$  by the elimination of implications if there are two distinct rules  $r$  and  $r'$  of  $P$  such that  $r'$  is an implication of  $r$  and  $P' = P \setminus \{r'\}$ .

In the above example,  $r_7$  is an implication of  $r'_4$  and thus  $P_6 = \{r_3, r'_4, r'_5, r_6, r_8\}$  is obtained from  $P_5$  by the elimination of implications.

**Elimination of Contradictions**  $P'$  is obtained from  $P$  by elimination of contradictions if there is a rule  $r$  in  $P$  such that  $body^+(r) \cap body^-(r) \neq \emptyset$  and  $P' = P \setminus \{r\}$ .

By the elimination of contradictions,  $r_8$  can be removed from  $P_6$  and thus we obtain  $P_7 = \{r_3, r'_4, r'_5, r_6\}$ .

**Unfolding**  $P'$  is obtained from  $P$  by unfolding if there is a rule  $r$  with  $body^+(r) \neq \emptyset$  such that

$$\begin{aligned} P' &= P \setminus \{r\} \\ &\cup \{ H(r, r', b) \leftarrow B(r, r', b) \mid b \in body^+(r), r' \in P, b \in head(r'), \\ &\quad H(r, r', b) = head(r) \cup (head(r') - \{b\}), \\ &\quad B(r, r') = body^+(r) \setminus \{b\}, not\ body^-(r), body(r') \}. \end{aligned}$$

$P_7$  is further transformed into  $P_8 = \{r'_3, r'_4, r_6\}$  where  $r'_3 : p \leftarrow not\ q_1$  is obtained from  $P_7$  by unfolding  $r_3$  and  $r'_4$ , and in particular,  $r'_5$  is removed as a special case of the unfolding since there is no rule in  $P_7$  whose head can be unfolded with the positive body literal  $p_2$  of  $r'_5$ .

Note that the above example is only used to illustrate the definitions of program transformations. In practice, one could make the process of simplifying  $P_1$  more efficient by choosing different transformations and different orderings.

As shown by [Brass and Dix, 1997], the program transformations from  $\mathbf{T}^*$  considered there, i.e., all except the elimination of head redundancy, preserve the answer set semantics. The same also holds for the latter transformation.

**Proposition 11** *Let  $P$  be a disjunctive program. If  $P'$  is obtained by the elimination of head redundancy from  $P$ , then  $P$  and  $P'$  have the same answer sets.*

Furthermore, by [Brass and Dix, 1999] every disjunctive program can be converted with the transformations in  $\mathbf{T}^*$  considered there into an equivalent negative disjunctive program. Moreover, by using the elimination of tautologies and the elimination of head redundancy (which preserves negative rules), every literal that occurs both in the head and the body of a rule can be removed. Thus, we have the following result.

**Lemma 1** *Every logic program  $P$  can be transformed into an equivalent negative program  $N$  via  $\mathbf{T}^*$  such that every rule  $r$  in  $N$  fulfills  $head(r) \cap body(r) = \emptyset$ .*

In fact, transformations can be applied in arbitrary manner and no backtracking is necessary to construct such a negative program  $N$  (i.e., choices for transformations are “don’t care”). In addition, we introduce only basic program transformations here. One can introduce some other program transformations, such as the elimination of s-implications in [Wang and Zhou, 2005], to further simplify the negative program obtained from basic transformations.

### 4.3.2 Logic programs with double negation

A *disjunctive logic program with double negation as failure (DDLDP)* is a finite set of rules  $r$  of the form

$$a_1 \vee \cdots \vee a_s \leftarrow b_1, \dots, b_m, \text{not } c_1, \dots, \text{not } c_n, \text{not not } d_1, \dots, \text{not not } d_t \quad (2)$$

where  $s, m, n, t \geq 0$  and all  $a_i, b_j, c_k,$  and  $d_l$  are from a set  $Lit$  of classical literals; as above, we assume that all  $a_i,$  and similarly all  $b_j,$  all  $c_k,$  and all  $d_l,$  are pairwise distinct. (Note that form (1) results for  $t = 0.$ ) The definition of  $head(r), body^+(r),$  and  $body^-(r)$  is analogous to ordinary rules (1), and  $body^{--}(r) = \{d_1, \dots, d_t\}.$  Thus,  $r$  can be denoted  $head(r) \leftarrow body^+(r), \text{not } body^-(r), \text{not not } body^{--}(r),$  where  $\text{not not } body^{--}(r) = \{\text{not not } l \mid l \in body^{--}(r)\}.$

Every DDLP  $P$  is a nested logic program [Lifschitz *et al.*, 1999], and inherits answer set semantics from such programs. Formally, the *reduct* of  $P$  w.r.t. an interpretation  $X$  is defined as  $P^X = \{head(r) \leftarrow body^+(r) \mid r \in P, body^-(r) \cap X = \emptyset, body^{--}(r) \subseteq X\}.$  As usual,  $X$  is an *answer set* of  $P$  iff  $X$  is a minimal model of  $P^X.$

Different from ordinary logic programs, the answer sets of DDLP may be comparable. For example, the DDLP  $P = \{p \leftarrow \text{not not } p\}$  has the two answer sets  $\emptyset$  and  $\{p\}.$  For our purposes, we will use those answer sets of  $P$  which are minimal.

Clearly, all answer sets of  $P$  are minimal iff they are all incomparable, and the latter is equivalent to the fact that  $P$  can be rewritten to an ordinary logic program. While this is difficult and expensive to check in general, there is an attractive syntactic class for which a simple rewriting exists.

Similar as for logic programs with default negation in rule heads [Inoue and Sakama, 1998], double negations can be safely eliminated from a DDLP without changing its semantics, if there is no cycle through positive and double negated dependencies.

**Definition 4** A DDLP  $P$  is *N-acyclic*, if there is a level mapping  $L : Lit_P \rightarrow \{0, 1, \dots\}$  of the literals in  $P$  to the non-negative integers such that for every rule  $r$  in  $P,$  the following two conditions hold:

- (i)  $L(l) \geq L(l')$  for all  $l \in head(r)$  and  $l' \in body^+(r).$
- (ii)  $L(l) > L(l')$  for all  $l \in head(r)$  and  $l' \in body^{--}(r).$

Note that there are no conditions  $L(l) > L(l'),$  for all  $l \in head(r)$  and  $l' \in body^-(r),$  and  $L(l) = L(l'),$  for all literals  $l, l' \in head(r)$  as in the familiar definition of stratified logic programs.

Given a DDLP  $P,$  let  $\mathcal{T}(P)$  be the logic program obtained from  $P$  by canceling every double negation *not not* in it. For example, if  $P_0 = \{p \leftarrow q, \text{not } q', \text{not not } q''\},$  then  $\mathcal{T}(P_0) = \{p \leftarrow q, \text{not } q', q''\}.$  We have the following result.

**Theorem 7** For every N-acyclic DDLP  $P,$  it holds that  $AS(P) = AS(\mathcal{T}(P)).$

(See Appendix for the proof.) Note that since every ordinary logic program is trivially N-acyclic, we can view N-acyclic DDLPs as a syntactic extension of ordinary logic programs. The fact that N-acyclic DDLPs can be easily cast to ordinary logic programs will be used for transformation-based forgetting.

### 4.3.3 The algorithm

We are now in a position to present a syntax-based algorithm for computing forgetting in a logic program. The algorithm  $\text{forget}_3(P, l),$  which is shown in Figure 4, first translates the input program  $P$  into a negative program  $N$  (Step 1) and then separates  $l$  from head disjunction (Step 2). After that,  $l$  is eliminated from rule

**Algorithm**  $\text{forget}_3(P, l)$ 

*Input:* Consistent (disjunctive) logic program  $P$  and a literal  $l$  in  $P$ .

*Output:* DDLP logic program  $N'$  whose minimal answer sets are  $\mathcal{AS}(\text{forget}(P, l))$ .

*Method:*

*Step 1.* Apply program transformations in  $\mathbf{T}^*$  on  $P$  to obtain a negative program  $N_0$ .

*Step 2.* Separate  $l$  from head disjunction via semi-shifting:

- Replace each rule  $r \in N_0$  such that  $\text{head}(r) = l \vee l_1 \vee \dots \vee l_k$ , where  $k \geq 1$ , by the two rules  $l \leftarrow \text{not } l_1, \dots, \text{not } l_k, \text{body}(r)$  and  $l_1 \vee \dots \vee l_k \leftarrow \text{not } l, \text{body}(r)$ .

Let  $N$  be the resulting logic program.

*Step 3.* Suppose that  $r_1, \dots, r_n$  are the rules in  $N$  with head  $l$ , where  $r_j : l \leftarrow \text{not } l_{j1}, \dots, \text{not } l_{jm_j}$  and  $m_j \geq 0$  for  $1 \leq j \leq n$ . Distinguish three cases:

- 3.1 If  $n = 0$ , then obtain the program  $Q$  by removing in  $N$  all literals  $\text{not } l$ .
- 3.2 If  $n > 0$  and  $m_j = 0$  for some  $1 \leq j \leq n$  (i.e.,  $l \leftarrow$  is a rule in  $N$ ), then obtain the program  $Q$  by removing from  $N$  all rules whose bodies contain  $\text{not } l$ .
- 3.3 If  $n > 0$  and  $m_j > 0$  for all  $1 \leq j \leq n$ , let  $D_1, \dots, D_s$  be all possible conjunctions of form  $\text{not } \text{not } l_{1k_1}, \dots, \text{not } \text{not } l_{nk_n}$  where  $0 \leq k_1 \leq m_j, 1 \leq j \leq n$ .  
Obtain the program  $Q$  by replacing in  $N$  each  $\text{not } l$  by all  $D_i$  (one at a time).

*Step 4.* Remove all rules with  $l$  in the head from  $Q$  and output the resulting program  $N'$ .

Figure 4: Syntax-based algorithm to compute forgetting

bodies (Step 3), and finally from rule heads (Step 4). The resulting output program is, in general, a logic program with double negation as failure.

**Example 6** Consider  $P_4 = \{c \leftarrow \text{not } q. p \leftarrow \text{not } q. q \leftarrow \text{not } p\}$ . Then, in Step 1 we have  $N_0 = P_4$  since  $P_4$  is already negative, and in Step 2  $N = P_4$  since  $P_4$  is normal. In Step 3, we have  $n = 1$  and  $r_1 = p \leftarrow \text{not } q$ . Thus, case three applies and we have  $D_1 = \text{not } \text{not } q$ ; we obtain  $Q = \{c \leftarrow \text{not } q. p \leftarrow \text{not } q. q \leftarrow \text{not } \text{not } q\}$ . In Step 4, the program  $N' = \{c \leftarrow \text{not } q. q \leftarrow \text{not } \text{not } q\}$  is output. This program has the answer sets  $\{c\}$  and  $\{q\}$ , which are both minimal. They are the same as the answer sets of  $\text{forget}(P_4, p)$ .

Note that Algorithm 1 in [Wang *et al.*, 2005] outputs on the input of  $P_4$  and  $p$  the program  $N' = \{c \leftarrow \text{not } q. q \leftarrow q\}$ , which has the single answer set  $\{c\}$ . However, the semantic result of forgetting about  $p$  in  $P_4$ , as defined [Wang *et al.*, 2005] and in this paper, has another answer set, viz.  $\{q\}$ . This shows that Algorithm 1 in [Wang *et al.*, 2005] is incomplete, i.e., outputs in general a logic program that represents only a subset of all answer sets after forgetting.

Several remarks on Algorithm  $\text{forget}_3(P, l)$  are in order. (1) As formulated here, the algorithm is stated in a very general form. A number of refinements and improvements can be made in order to make it more efficient and the result more compact. For example, in Step 1 some program transformations could be omitted for special programs and various heuristics could also be employed. In Step 3, only those  $D_i$  need to be considered which, after removal of duplicate literals, are not properly contained in some other  $D_j$ . To compute them, one can use efficient hypergraph transversal algorithms (see [Eiter *et al.*, 2007]).

(2) In the construction of  $D_i$ , *not not*  $l_{ij}$  cannot be replaced with  $l_{ij}$  (even for a normal logic program). As shown by Example 6, the resulting output program  $\{c \leftarrow \text{not } q. q \leftarrow q\}$  would only represent a subset of  $\mathcal{AS}_l(P)$ , and thus would be incorrect. The use of double negation as failure, which remedies this problem, seems to be intuitive. It remains as an interesting issue whether this can be avoided in a similar transformation based algorithm.

(3) The running time of algorithm  $\text{forget}_3(P, l)$  is worst case exponential, and the output program may be exponentially large. As follows from complexity considerations in Section 5, there is no ordinary or nested logic program  $P'$  representing  $\text{forget}(P, l)$  which can be constructed in polynomial time, even if auxiliary literals might be used which are projected from the answer sets of  $P'$ .

(4) In essence, algorithm  $\text{forget}_3(P, l)$  improves the corresponding Algorithm 1 in [Wang *et al.*, 2005] in at least two ways: it works for the more expressive class of disjunctive logic programs, and importantly, its output correctly represents the result of forgetting. This is shown formally by the following result.

**Theorem 8** *Let  $P$  be a consistent disjunctive logic program and let  $l \in \text{Lit}_P$  be a literal. Then  $\text{forget}_3(P, l)$  correctly represents  $\text{forget}(P, l)$ , i.e.,  $X$  is an answer set of  $\text{forget}(P, l)$  iff  $X$  is a minimal answer set of  $N'$ .*

(For a proof, see the Appendix.) While the output program of  $\text{forget}_3(P, l)$  generally contains double negation as failure and its minimal answer sets have to be considered, Theorem 7 provides a simple condition for transforming the output of Algorithm  $\text{forget}_3(P, l)$  into an ordinary logic program in some cases.

**Proposition 12** *Let  $P$  be a consistent logic program, let  $l \in \text{Lit}_P$  be a literal, and let  $N'$  be the output of Algorithm  $\text{forget}_3(P, l)$ . If  $N'$  is  $N$ -acyclic, then  $\mathcal{AS}_l(P) = \mathcal{AS}(T(P'))$ , i.e., the ordinary logic program  $T(P')$  correctly represents  $\text{forget}(P, l)$ .*

For arbitrary inputs, it is not clear whether the output of  $\text{forget}_3(P, l)$  is an  $N$ -acyclic program. We note here a relevant subclass which has this property, given by a simple syntactic condition that can be efficiently recognized.

Recall that the standard dependency graph of a program  $P$ , denoted  $DG_P$ , has  $\text{Lit}_P$  as vertices and a positive edge from literal  $l$  to literal  $l'$  if  $l \in \text{head}(r)$  and  $l' \in \text{body}^+(r) \cup (\text{head}(r) \setminus \{l\})$  for some rule  $r \in P$  and a negative edge from  $l$  to  $l'$  if  $l \in \text{head}(r)$  and  $l' \in \text{body}^-(r)$  for some rule  $r \in P$ . A (directed) cycle in the graph  $DG_P$  is negative, if it contains at least one negative edge.

**Proposition 13** *Let  $P$  be a consistent normal logic program and let  $l \in \text{Lit}_P$  be a literal. If no negative cycle of  $DG_P$  contains  $l$ , then  $\text{forget}_3(P, l)$  outputs an  $N$ -acyclic program.*

In Example 2, for  $l = p$  the programs  $P_1$ ,  $P_3$ , and  $P_4$  have this property (as well as  $P_6$  after removal of the redundant disjunctive rule). Hence, the outputs of the respective calls  $\text{forget}_3(P_i, p)$  can be cast to ordinary logic programs.

Unfortunately, the extension of Proposition 13 from normal to disjunctive logic programs fails, even for the simple case where  $P = \{p \vee q \leftarrow \cdot\}$  and  $l = p$ . However, it is possible to single out fragments for which this is possible, which we leave for future work.

Computing a representation of  $\text{forget}(P, l)$  can be refined in different directions. One is to localize the computation, such that only a relevant part of the program  $P$  is subject to modification when forgetting a literal  $l$ , while the rest of  $P$  remains untouched. This is, for example, easy if  $P$  is a normal logic program; in that case, merely the value of  $l$  in the single answer set of  $P$  has to be plugged in for all occurrences of  $l$  in  $P$  and the resulting program be simplified.

Another case is if  $P$  splits into separate components  $P_1 \cup \dots \cup P_n$  such that the answer sets of  $P$  can be obtained by combining the local answer sets of each  $P_i$ .

To this end, we call a literal  $l'$  in a program  $P$  *unaffected* by the forgetting of literal  $l$ , if there is no path between  $l$  and  $l'$  in the undirected version of the dependency graph  $DG_P$ . Suppose  $U(P, l)$  is the set of all rules in  $P$  that only involve such literals  $l'$ . Then the following easy property holds.

**Proposition 14** *For every (consistent) program  $P$  and literal  $l$ , it holds that  $\text{forget}(P, l) \equiv U(P) \cup \text{forget}(P \setminus U(P), l)$ .*

This property is in line with  $N$ -acyclicity: disconnected components cannot destroy the  $N$ -acyclicity of the (rewritten) program. Furthermore, it can be combined with Proposition 13 to enlarge the classes of programs  $P$  for which  $\text{forget}(P, l)$  is representable by an ordinary logic program.

Note that Proposition 14 is independent of the concrete syntactic form of  $\text{forget}(P, l)$ ; it may well be generalized for specific such forms and/or classes of programs. For example, if  $P$  can be split into programs  $P_1$  and  $P_2$  such that  $l$  occurs only in  $P_1$ , program  $P_2$  has a single answer set  $S$  (e.g. if  $P_2$  is normal and stratified), and no head of a rule in  $P_1$  occurs in  $P_2$ , then  $\text{forget}_3(P, l) \equiv \text{forget}_3(P_1 \cup \{l' \leftarrow . \mid l' \in S\}, l) \cup P_2$ . A detailed study of this issue remains for future work.

## 5 Computational Complexity

In this section, we address the computational complexity of forgetting for different classes of logic programs. Our main complexity results for forgetting are compactly summarized in Table 1. They show that for general logic programs, (1) model checking under forgetting is  $\Pi_2^P$ -complete; (2) credulous reasoning under forgetting is  $\Sigma_3^P$ -complete; and (3) skeptical reasoning under forgetting is  $\Pi_2^P$ -complete.

Intuitively, this complexity is explained by two respectively three intermingled sources of complexity (i)–(iii): For problem (1), given an answer set  $S$  of a program  $P$  and a literal  $l$ , (i) the number of candidate answer sets  $S'$  such that  $S' \subset_l S$  and (ii) the test whether a given such  $S'$  is in fact an answer set of  $P$ ; for problem (2), in addition (iii) the number of candidate  $l$ -answer sets  $S$  containing the query literal  $l'$ . Note, however, that for problem (3) source (i) is absent (by Proposition 7) and only (ii) and (iii) (in dual form) are present, causing the same complexity as for standard skeptical reasoning (without forgetting).

For normal programs and negative logic programs, the complexity of all problems is lowered by one level of the Polynomial Hierarchy. Intuitively, the reason is that source (ii), i.e., model checking for such programs, is polynomial in both cases.

In the rest of this section, we state and develop the complexity results formally, and also argue that space-efficient representations of  $\text{forget}(P, l)$  in terms of ordinary (disjunctive) logic programs are unlikely to exist. The design of Algorithm  $\text{forget}_3(P, l)$  in Section 4 is heavily influenced by the complexity analysis.

**Theorem 9** *Given a consistent (disjunctive) logic program  $P$ , a literal  $l$ , and a set of literals  $X$ , deciding whether  $X$  is an  $l$ -answer set of  $P$  is  $\Pi_2^P$ -complete.*

Intuitively, in order to show that  $X$  is an  $l$ -answer set, we have to witness that  $X$  is an answer set (which is coNP-complete to test), and that there is no answer set  $X'$  of  $P$  such that  $X' \subset_l X$ . Any  $X'$  disproving this can be guessed and checked using an NP-oracle in polynomial time. Thus,  $l$ -answer set checking is in  $\Pi_2^P$ , as stated in Theorem 9. The hardness result is shown by a reduction from deciding whether a given logic program  $P$  (without strong negations) has no answer set, which is  $\Pi_2^P$ -complete [Eiter and Gottlob, 1995b].

If  $P$  is either negative or normal,  $l$ -answer checking is co-NP-complete.



program $P$	disjunctive	negative	normal
$X \in \mathcal{AS}_l(P)?$ (model checking)	$\Pi_2^p$	co-NP	co-NP
$\text{forget}(P, l) \models_c l'?$	$\Sigma_3^p$	$\Sigma_2^p$	$\Sigma_2^p$
$\text{forget}(P, l) \models_s l'?$	$\Pi_2^p$	co-NP	co-NP

Table 1: Complexity of forgetting (entries are completeness results;  $l, l'$  are literals,  $P$  is a logic program,  $X$  is a set of literals)

**Theorem 10** *Given a consistent normal logic program  $P$ , a literal  $l$ , and a set of literals  $X$ , deciding whether  $X$  is an  $l$ -answer set of  $P$  is co-NP-complete.*

The proof of this theorem exploits that the reduction in the proof of Theorem 9 still works for normal programs, and that deciding whether a normal logic program has an answer set is well-known to be NP-complete [Marek and Truszczyński, 1991; Ben-Eliyahu and Dechter, 1994].

Using a minor modification of the reduction in the proof of Theorem 9, we can show the co-NP-completeness for negative programs. Notice that, as already mentioned, deciding whether a given set of literals is an answer set of negative program is feasible in polynomial time, which explains the complexity drop.

**Theorem 11** *Given a consistent negative program  $P$ , a literal  $l$ , and a set of literals  $X$ , deciding whether  $X$  is an  $l$ -answer set of  $P$  is co-NP-complete.*

The following theorem shows that credulous reasoning with forgetting has a higher complexity.

**Theorem 12** *Given a consistent (disjunctive) logic program  $P$  and literals  $l$  and  $l'$ , deciding whether  $\text{forget}(P, l) \models_c l'$  is  $\Sigma_3^p$ -complete.*

In Theorem 12, a suitable  $l$ -answer set containing  $l'$  can be guessed and checked, by Theorem 9 using  $\Sigma_2^p$ -oracle. Hence, credulous inference  $\text{forget}(P, l) \models_c l'$  is in  $\Sigma_3^p$ . The  $\Sigma_3^p$ -hardness is shown by an encoding of quantified Boolean formulas (QBFs) of the form  $\exists Z \forall X \exists Y \phi$ .

The construction in the proof of Theorem 11 can be lifted to show that credulous inference with forgetting is  $\Sigma_2^p$ -complete for negative programs.

**Theorem 13** *Given a consistent negative program  $P$  and literals  $l$  and  $l'$ , deciding whether  $\text{forget}(P, l) \models_c l'$  is  $\Sigma_2^p$ -complete.*

In fact, the program constructed to show the hardness part of this result is normal. Therefore, we easily derive the following result.

**Theorem 14** *Given a consistent normal program  $N$  and literals  $l$  and  $l'$ , deciding whether  $\text{forget}(N, l) \models_c l'$  is  $\Sigma_2^p$ -complete.*

The complexity results for skeptical reasoning with forgetting are straightforward from Proposition 7 and well-known results about the complexity of normal logic programs (see [Eiter and Gottlob, 1995b; Marek and Truszczyński, 1991; Dantsin *et al.*, 2001]).

**Theorem 15** *Given a consistent logic program  $P$  and literals  $l$  and  $l'$ , deciding whether  $\text{forget}(P, l) \models_s l'$  is (i)  $\Pi_2^p$ -complete for arbitrary disjunctive logic programs  $P$ , and (ii) co-NP-complete for normal logic programs and for negative logic programs  $P$ .*

By applying techniques that build on non-uniform complexity classes from [Cadoli *et al.*, 2000], one can show that for a given (disjunctive) program  $P$  and literal  $l$  there is generally no ordinary disjunctive program  $P'$  representing  $\text{forget}(P, l)$  that has size polynomial in the size of  $P$ , unless the Polynomial Hierarchy collapses. This remains true even if auxiliary literals might be used in  $P'$  for the representation which are projected off the models of  $P'$  to obtain the models of  $\text{forget}(P, l)$ . This means that the exponential blow up of  $\text{forget}(P, l)$  is, in a formal sense, unavoidable in general.

More precisely, it can be shown that the model checking problem for forgetting is complete for the complexity class  $\|\sim\!\Pi_2^p$  defined in [Cadoli *et al.*, 2000; Cadoli *et al.*, 2002]. Informally, this means that problem is among the hardest in  $\|\sim\!\Pi_2^p$ , which contains those problems that are decidable in  $\Pi_2^p$  with preprocessing of the input (which depends on a “fixed” part and the size of the input). The preprocessing can resort to precompiled knowledge in polynomial-size data structures, where the compilation cost does not count. Technically, the  $\Pi_2^p$ -completeness of model checking for forgetting implies that (a syntactic variant of) the problem belongs to  $\|\sim\!\Pi_2^p$ . On the other hand, the problem is  $\|\sim\!\Pi_2^p$ -hard, as evaluating a QBF of the form  $\forall X \exists Y \phi$  can be reduced (under the suitable notion of reduction) to model checking for forgetting. The proof is similar in spirit to the one of Theorem 3.2 in [Cadoli *et al.*, 2002], which shows  $\|\sim\!\Pi_2^p$  hardness of clause inference from the minimal models of a propositional CNF, but uses the encoding of QBFs into model checking for forgetting given in the proof of Theorem 9 via [Eiter and Gottlob, 1995b]. We refrain here from further details.

Now while model checking for forgetting is  $\|\sim\!\Pi_2^p$ -hard, model checking for ordinary disjunctive programs is well-known to be in co-NP (cf. [Eiter and Gottlob, 1995a; Dantsin *et al.*, 2001]). From Theorem 5 in [Cadoli *et al.*, 2000], it follows that for arbitrary logic programs  $P$  and literals  $l$  there exists no representation of  $\text{forget}(P, l)$  by an ordinary (disjunctive) logic program of size polynomial in the size of  $P$  unless the Polynomial Hierarchy collapses (which is considered to be unlikely).

Analogously, one can show that for normal programs  $P$ ,  $\text{forget}(P, l)$  is not representable by normal programs of polynomial size in the size of  $P$ , unless the Polynomial Hierarchy collapses, again even if auxiliary literals might be used as above.

However, we point out that if auxiliary literals would be allowed, then we can represent  $\text{forget}(P, l)$  in terms of the minimal answer sets of a polynomial-size logic program with double negation as failure. More precisely, let  $s$  and  $p_l$  be fresh auxiliary literals, and use  $l$  itself as an auxiliary literal. Let

- $P_1$  be the program resulting from  $P$  by adding *not*  $s$  in each rule body, let
- $P_2$  be the program resulting from  $P$  by replacing each occurrence of  $l$  with  $p_l$  and by adding  $s$  in each rule body, and let
- $Q = P_1 \cup P_2 \cup \{s \leftarrow \text{not not } s. \quad l \leftarrow s. \quad \leftarrow \text{not } l, \text{not } s. \quad \leftarrow p_l, s\}$ .

Note that  $Q$  is easily constructed from  $P$  and  $l$  (in linear time). Informally,  $s$  is a switch between  $P_1$  and  $P_2$ , to compute the answer sets of  $P$  where  $l$  is true (via  $P_1$  and the constraint  $\leftarrow \text{not } l, \text{not } s$ , when  $s$  is false) respectively false (via  $P_2$ , when  $s$  is true). In order to make these answer sets  $l$ -comparable, in the computations of  $P_2$  the literal  $p_l$  replaces  $l$  and  $l$  is artificially included (by  $l \leftarrow s$ ) while  $p_l$  is excluded (through  $\leftarrow p_l, s$ ). The answer sets  $S$  of  $Q$  (which all contain  $l$  and but not  $p_l$ ) correspond then one-to-one to the answer sets of  $P$ . Now every  $S$  such that  $s \notin S$  is a minimal answer set of  $Q$ , while if  $s \in S$ , then

$S$  is minimal iff it contains no answer set  $S'$  of  $Q$  such that  $s \notin S'$  properly. Consequently, the minimal answer sets  $S$  of  $Q$  encode the  $l$ -answer sets of  $P$ , and thus the answer sets of  $\text{forget}(P, l)$ , which are given by  $S \setminus \{l, s\}$ .

## 6 Applications

In this section, we present some applications of the results on forgetting from logic programs in the previous sections. In particular, we consider applications to conflict resolution in logic-based multi-agent systems, which is obviously an important task, to inheritance logic programs, which model objects and classes with inheritance of properties, and to logic program updates. A further application in the area of ontology merging and alignment (in an extended framework), is described in [Eiter *et al.*, 2006]. The applications show the usefulness of our results about forgetting in different respects: On the one hand, for finding novel solutions to problems (like in conflict resolution and ontology merging), and on the other hand, to obtain novel characterizations (and thus interpretations) of existing concepts (like for inheritance logic programs and logic program updates).

### 6.1 Resolving Conflicts in Multi-Agent Systems

As the first application, we present a general framework for resolving conflicts in multi-agents systems, which is inspired from the *preference recovery* problem [Lang and Marquis, 2002]. In particular, an example is given to show the elegance of using the semantic forgetting in answer set programming to solve the problem of preference recovery for multi-agents.

Suppose that there are  $n$  agents who may have different preferences on the same issue. In many cases, these preferences (or constraints) have conflicts and thus cannot be satisfied at the same time. It is an important issue in constraint reasoning to find intuitive criteria such that preferences with higher priorities are satisfied. Consider the following example.

**Example 7 ([Lang and Marquis, 2002])** *Suppose that a group of four residents in a complex tries to reach an agreement on building a swimming pool and/or a tennis court. The preferences and constraints are as follows.*

1. *Building a tennis court or a swimming pool costs each one unit of money.*
2. *A swimming pool can be either red or blue.*
3. *The first resident would not like to spend more than one money unit, and prefers a red swimming pool.*
4. *The second resident would like to build at least one of tennis court and swimming pool. If a swimming pool is built, he would prefer a blue one.*
5. *The third resident would prefer a swimming pool but either colour is fine with him.*
6. *The fourth resident would like both tennis court and swimming pool to be built. He does not care about the colour of the pool.*

*Obviously, the preferences of the group are jointly inconsistent and thus it is impossible to satisfy them at the same time.*

In the following, we will show how to resolve this kind of preference conflicts and find possible agreements with minimal costs using the theory of forgetting.

An  $n$ -agent system  $\mathcal{S}$  is an  $n$ -tuple  $(P_1, P_2, \dots, P_n)$  of logic programs,  $n > 0$ , where  $P_i$  represents agent  $i$ 's knowledge (including preferences, constraints).

As shown in Example 7,  $P_1 \cup P_2 \cup \dots \cup P_n$  may be inconsistent. The basic idea in our approach is to forget some literals for each agent so that conflicts can be resolved.

**Definition 5** Let  $\mathcal{S} = (P_1, P_2, \dots, P_n)$  be an  $n$ -agent system. A compromise of  $\mathcal{S}$  is a sequence  $C = (F_1, F_2, \dots, F_n)$  where each  $F_i$  is a set of literals. An agreement of  $\mathcal{S}$  on  $C$  is an answer set of the logic program  $\text{forget}(\mathcal{S}, C)$  where  $\text{forget}(\mathcal{S}, C) = \text{forget}(P_1, F_1) \cup \text{forget}(P_2, F_2) \cup \dots \cup \text{forget}(P_n, F_n)$ .

Intuitively, the set  $F_i$  in a compromise contains those aspects which agents  $i$  does not care much about. For a specific application, we may need to impose certain conditions on each  $F_i$ .

**Example 8** (Example 7 continued) The scenario can be encoded as a collection of five disjunctive programs ( $P_0$  stands for general constraints):  $\mathcal{S} = (P_0, P_1, P_2, P_3, P_4)$  where

$$\begin{aligned} P_0 &= \{ \text{red} \vee \text{blue} \leftarrow s. \quad \leftarrow \text{red}, \text{blue}. \quad u_0 \leftarrow \text{not } s, \text{not } t. \\ &\quad u_1 \leftarrow \text{not } s, t. \quad u_1 \leftarrow s, \text{not } t. \quad u_2 \leftarrow s, t. \quad \}; \\ P_1 &= \{ u_0 \vee u_1 \leftarrow . \quad \text{red} \leftarrow s \}; \\ P_2 &= \{ s \vee t \leftarrow . \quad \text{blue} \leftarrow s \}; \\ P_3 &= \{ s \leftarrow \}; \\ P_4 &= \{ s \leftarrow . \quad t \leftarrow \}. \end{aligned}$$

Since this knowledge base is jointly inconsistent, each resident may have to weaken some of her preferences so that an agreement is reached. Some possible compromises are:

1.  $C_1 = (\emptyset, F, F, F, F)$  where  $F = \{s, \text{blue}, \text{red}\}$ : Every resident would be willing to weaken her preferences on the swimming pool and its colour. Since  $\text{forget}(\mathcal{S}, C_1) = P_0 \cup \{u_0 \vee u_1 \leftarrow . \quad t \leftarrow\}$ ,  $\mathcal{S}$  has a unique agreement  $\{t, u_1\}$  on  $C_1$ . That is, only a tennis court is built.
2.  $C_2 = (\emptyset, F, F, F, F)$  where  $F = \{u_0, u_1, u_2, \text{blue}, \text{red}\}$ : Every resident can weaken her preferences on the price and the pool colour. Since  $\text{forget}(\mathcal{S}, C_2) = P_0 \cup \{s \vee t \leftarrow . \quad s \leftarrow . \quad t \leftarrow\}$ ,  $\mathcal{S}$  has two possible agreements  $\{s, t, \text{red}\}$  and  $\{s, t, \text{blue}\}$  on  $C_2$ . That is, both a tennis court and a swimming pool will be built but the pool colour can be either red or blue.
3.  $C_3 = (\emptyset, \{\text{blue}, \text{red}\}, \emptyset, \emptyset, \{t\})$ : The first resident can weaken her preference on pool colour and the fourth resident can weaken her preference on tennis court. Since  $\text{forget}(\mathcal{S}, C_3) = P_0 \cup P_2 \cup P_3 \cup \{u_0 \vee u_1 \leftarrow . \quad s \vee t \leftarrow . \quad s \leftarrow\}$ ,  $\mathcal{S}$  has a unique agreement  $\{s, \text{blue}, u_1\}$  on  $C_3$ . That is, only a swimming pool will be built and its colour is blue.
4.  $C_4 = (\emptyset, \{\text{blue}, \text{red}\}, \{\text{blue}, \text{red}\}, \{s, t\}, \{s, t\})$ : The first and second residents can weaken her preference on pool colour; the third and fourth residents would not mind if tennis court or swimming pool is built. Since  $\text{forget}(\mathcal{S}, C_4) = P_0 \cup \{u_0 \vee u_1 \leftarrow . \quad s \vee t \leftarrow\}$ ,  $\mathcal{S}$  has three possible agreements  $\{u_1, t\}$ ,  $\{u_1, s, \text{blue}\}$ ,  $\{u_1, s, \text{red}\}$  on  $C_4$ .

It should be noted that a solution to this problem is also provided in [Lang and Marquis, 2002] where the forgetting for propositional logic is used where a theory in propositional logic rather than possible agreements are produced. A model of that theory may not represent an agreement in the sense of Definition 5. For example, the solution for  $C_4$  given in [Lang and Marquis, 2002] is the theory  $T_4 = \{s \rightarrow red \vee blue, (red \wedge blue) \rightarrow \perp, (\neg s \wedge t) \rightarrow u_1, (s \wedge \neg t) \rightarrow u_1, (s \wedge t) \rightarrow u_2, (\neg s \wedge \neg t) \rightarrow u_0\}$  and its models also include some agreements that have non-minimal costs (e.g.  $\{u_2, s, t, red\}$  is a model of  $T_4$ ).

However, in our approach each answer set corresponds to exactly one agreement with minimal cost. In addition, the issue of resolving conflicts in multi-agent systems has been challenging and numerous proposals have been suggested for different systems (for example, see [Tessier *et al.*, 2001]). Thus it would be interesting to explore applications of our technique in practical multi-agent systems.

## 6.2 Inheritance Logic Programs

In this section, we investigate relationships between forgetting in logic programs and inheritance logic programs [Buccafurri *et al.*, 2002]. As we show, the semantics of such programs can be expressed by forgetting from a logic program.

Let  $P$  be a logic program with classical negation and each rule  $r$  of  $P$  is labeled with either the symbol ‘.’ or the symbol ‘!’. The symbol ‘.’ means that  $r$  is a *defeasible* rule and the symbol ‘!’ means that  $r$  is a *strict* rule. In the approach proposed in [Buccafurri *et al.*, 2002],  $P$  is an *inheritance program* if  $P$  is classified into different objects and an object may have higher priority than another object. For any two objects  $o_1$  and  $o_2$ ,  $o_1 < o_2$  denotes that  $o_1$  has higher priority over  $o_2$ . This priority relation naturally defines a priority for rules in these two objects:  $r_1 < r_2$  if  $r_1 \in o_1, r_2 \in o_2$  and  $o_1 < o_2$ .

**Example 9** Let  $(P, <)$  be an inheritance logic program that consists of three objects:  $o_1, o_2, o_3$  where

$$\begin{aligned} o_1 &= \{ penguin(Tweety) \leftarrow! \} \\ o_2 &= \{ bird(x) \leftarrow penguin(x)!, \neg flies(x) \leftarrow penguin(x). \} \\ o_3 &= \{ flies(x) \leftarrow bird(x). \}. \end{aligned}$$

$o_1 < o_2 < o_3$  since more specific rules have higher priority.

In the rest of this section, we view in accordance with [Buccafurri *et al.*, 2002] inheritance programs as pairs  $(P, <)$ , where  $P$  is a ground (propositional) logic program with classical negation and  $<$  is a strict preorder (irreflexive and transitive relation) on the rules in  $P$ , such that  $r < r'$  iff  $r$  has higher priority than  $r'$ .<sup>3</sup>

The semantics of inheritance programs is defined in terms of *inheritance answer sets*, which are based on the notion of *models* of an inheritance program. The notion of satisfiability of rules for inheritance programs encodes priority information and thus is quite different from the traditional notion.

Given two (ground) rules  $r_1$  and  $r_2$ , we say  $r_1$  *threatens*  $r_2$  on a literal  $l$  if (1)  $\neg.l \in head(r_1)$ , (2)  $r_1 < r_2$ , and (3)  $r_2$  is defeasible (recall that  $\neg.l$  denotes the complement of literal  $l$ ).

**Definition 6** Given an inheritance program  $(P, <)$  and an interpretation  $S$ , a rule  $r_1$  overrides  $r_2$  in  $S$  if (1)  $r_1$  threatens  $r_2$  on a literal  $l$ , (2)  $\neg.l \in S$  and (3)  $S \models body(r_2)$ . A rule  $r$  is overridden in  $S$  if for each  $l \in head(r)$  there exists a rule  $r'$  in  $P$  such that  $r'$  overrides  $r$  on  $l$  in  $S$ .

<sup>3</sup>Cf. [Buccafurri *et al.*, 2002, p.297] for technical assumptions to ensure this.

Informally, a rule  $r$  is overridden by another rule  $r'$ , if it has lower priority than  $r'$  and is in conflict with  $r'$ . Obviously, a strict rule cannot be overridden.

An interpretation  $S$  is a model of  $(P, <)$ , if every rule in  $P$  is either satisfied or overridden in  $S$ . The Gelfond-Lifschitz reduct is extended to inheritance programs as follows.

**Definition 7** *Given an inheritance program  $(P, <)$  and an interpretation  $S$ , the reduction of  $(P, <)$  w.r.t.  $S$ , denoted  $(P, <)^S$ , is the set of rules obtained from  $P$  by removing (1) every rule overridden in  $S$ , (2) every rule  $r$  such that  $\text{body}^-(r) \cap S \neq \emptyset$ , and (3)  $\text{body}^-(r)$  from each remaining rule  $r$ .*

*An interpretation  $S$  is an inheritance answer set of  $(P, <)$ , if  $S$  is a minimal model of  $(P, <)^S$ .*

**Example 10** *The ground version of the inheritance program in Example 9,  $(\text{ground}(P), <)$ , has the single inheritance answer set  $S = \{\text{penguin}(\text{Tweety}), \text{bird}(\text{Tweety}), \neg \text{flies}(\text{Tweety})\}$ . Indeed, the instance of the rule in  $o_3$ ,  $\text{flies}(\text{Tweety}) \leftarrow \text{bird}(\text{Tweety})$ , is overridden by the rule instance  $\neg \text{flies}(\text{Tweety}) \leftarrow \text{penguin}(\text{Tweety})$  from  $o_2$  in  $S$ ;  $(\text{ground}(P), <)^S$  consists of the three rules  $\text{penguin}(\text{Tweety}) \leftarrow$ ,  $\text{bird}(\text{Tweety}) \leftarrow \text{penguin}(\text{Tweety})$ , and  $\neg \text{flies}(\text{Tweety}) \leftarrow \text{penguin}(\text{Tweety})$ . Clearly,  $S$  is their unique answer set, which means that  $S$  is an inheritance answer set of  $(\text{ground}(P), <)$ .*

Let  $(P, <)$  be an inheritance program and  $S$  be a set of literals. We introduce a new literal  $l'$  for each literal  $l$  in  $P$ . For each rule  $r$  in  $P$ , if  $r$  is overridden in  $S$ , then every literal  $l$  in  $\text{head}(r)$  is replaced with  $l'$ . The resulting program from  $P$  is denoted  $P'$ . Let  $F = \{l' \mid l' \text{ is a new literal and } l' \in \text{head}(r') \text{ for some } r' \in P'\}$ .

The following theorem provides a semantic characterization of inheritance programs in terms of semantic forgetting.

**Theorem 16** *Let  $(P, <)$  be an inheritance program and let  $S$  be a set of literals. Then  $S$  is an inheritance answer set of  $(P, <)$  iff  $S$  is an answer set of  $\text{forget}(P', F)$  where  $P'$  is obtained as above.*

**Example 11** *Continuing our birds example, for  $S = \{\text{penguin}(\text{Tweety}), \text{bird}(\text{Tweety}), \neg \text{flies}(\text{Tweety})\}$  the corresponding ordinary logic program  $\text{ground}(P)'$  consists of the following rules:*

$$\begin{aligned} \text{penguin}(\text{Tweety}) &\leftarrow , \\ \text{bird}(\text{Tweety}) &\leftarrow \text{penguin}(\text{Tweety}), \\ \neg \text{flies}(\text{Tweety}) &\leftarrow \text{penguin}(\text{Tweety}), \\ \text{flies}(\text{Tweety})' &\leftarrow \text{bird}(\text{Tweety}). \end{aligned}$$

*For  $F = \{\text{flies}(\text{Tweety})'\}$ , we obtain that  $\text{forget}(\text{ground}(P)', F)$  is represented by the first three rules above. This program has the unique answer set  $S$ , as stated by Theorem 16.*

The proof of Theorem 16 is based on the following result, which is of independent interest.

**Proposition 15** *Let  $P$  be a logic program and let  $F$  be a consistent set of literals. Suppose that (1) no literal in  $F$  occurs in a rule body in  $P$ , and (2) for each rule  $r$ , either no or every literal in  $\text{head}(r)$  is in  $F$ . Then  $\text{forget}(P, F) = P \setminus R(F)$ , where  $R(F) = \{r \in P \mid r \text{ contains a literal of } F\}$ .*

Note that the conclusion of Proposition 15 may not be true if  $F$  contains opposite literals. For example, consider the logic program  $P = \{p \leftarrow a; \neg p \leftarrow a; a \leftarrow \text{not } b; b \leftarrow \text{not } a\}$  and  $F = \{p, \neg p\}$ . Then  $P \setminus R(F) = \{b \leftarrow \text{not } a; a \leftarrow \text{not } b\}$ , which has two answer sets  $\{a\}$  and  $\{b\}$ , while  $\text{forget}(P, F)$  has only one answer set  $\{b\}$ .

### 6.3 Update Logic Programs

Update programs [Eiter *et al.*, 2000; Eiter *et al.*, 2002] and dynamic logic programs [Alferes *et al.*, 1998; Alferes *et al.*, 2000] are besides [Marek and Truszczyński, 1998; Sakama and Inoue, 2003; Zhang and Foo, 1997] major approaches to updating nonmonotonic logic programs, and are in particular geared towards modeling sequences of updates (see [Zhang, 2006] for a recent survey and comparison of these and other approaches).

An *update program* is a sequence  $\mathcal{P} = [P_1, P_2, \dots, P_t]$ ,  $t \geq 1$ , where each  $P_i$  is a normal logic program for  $1 \leq i \leq t$ . Informally,  $P_{i+1}$  is assumed to update the information represented by  $[P_1, \dots, P_i]$ . So  $P_{i+1}$  represents more recent information than  $P_i$ , and the rules in  $P_{i+1}$  are assigned higher priority in case of conflicts. The semantics of update programs has been given by means of a translation into an ordinary logic program  $P_{\triangleleft}$ , and can be equivalently defined in terms of the answer sets  $S$  of an ordinary logic program  $\bigcup_{i=1}^t P_i \setminus \text{Rej}(S, \mathcal{P})$ , where  $\text{Rej}(S, \mathcal{P})$  is the set of rules which are rejected w.r.t.  $S$ .

It has been shown in [Eiter *et al.*, 2000; Eiter *et al.*, 2002] that every update program  $\mathcal{P}$  can be naturally translated into an equivalent inheritance logic program  $\text{ihp}(\mathcal{P}) = (P, <)$  where  $r < r'$  if  $r \in P_i$  and  $r' \in P_j$  such that  $1 \leq i < j \leq t$ . More precisely,

**Lemma 2 ([Eiter *et al.*, 2002])** *For every update program  $\mathcal{P} = [P_1, P_2, \dots, P_t]$ , a set  $S$  is an answer set of  $\mathcal{P}$  iff  $S$  is an answer set of the inheritance program  $\text{ihp}(\mathcal{P})$ .*

Combining this lemma and Theorem 16 gives us a characterization of update programs in terms of semantic forgetting from a logic program.

**Corollary 17** *Let  $\mathcal{P}$  be an update program and let  $S$  be a set of literals. Then  $S$  is an answer set of  $\mathcal{P}$  if and only if  $S$  is an answer set of  $\text{forget}(P', F)$ , where  $\text{ihp}(\mathcal{P}) = (P, <)$  and  $P'$  is obtained from  $(P, <)$  as described above.*

Finally, results in [Eiter *et al.*, 2002] show that for certain classes of update programs, the semantics coincides with the one under dynamic logic programming as in [Alferes *et al.*, 1998; Alferes *et al.*, 2000]. Hence, the respective classes of dynamic logic programs can also be characterized by semantic forgetting in the way described.

## 7 Conclusion

While it is widely acknowledged that forgetting about atomic propositions in knowledge bases is an important technique for many AI applications, it has been less clear how this should materialize in the context of nonmonotonic reasoning and logic programming. To the best of our knowledge, this paper is the first attempt towards identifying criteria for this operation in this context. In particular, we have specified some desirable properties for forgetting in nonmonotonic logic programming. Based on these criteria, we have then proposed a semantics-based theory of forgetting literals in (disjunctive) logic programming. Compared to preliminary work, a distinguishing feature of our approach is that forgetting is defined in purely semantic terms. However, we have shown that this declarative approach to forgetting has a syntactic counterpart based on program transformations.

The properties of forgetting show that the approach in this paper extends the classical notion of forgetting and, moreover, satisfies all criteria that we have identified. As we have explained before, it also naturally generalizes the notion of forgetting for normal programs investigated in [Wang *et al.*, 2005]. Furthermore,

we have presented algorithms and analyzed the computational complexity of major reasoning tasks under forgetting.

Another approach to forgetting for normal logic programs was proposed in [Zhang *et al.*, 2005; Zhang and Foo, 2006]. Different from ours, the approach by Zhang and colleagues is procedural. The result of forgetting is obtained by removing some rules and/or literals, but little semantic justification for the removals is provided from a global perspective.

As an application of forgetting, we have also presented a fairly general framework for resolving conflicts in disjunctive logic programming. In particular, this framework provides an elegant solution to the preference recovery problem. Furthermore, our results show that the semantic forgetting has a close relationship with inheritance programs [Buccafurri *et al.*, 1999], update programs [Eiter *et al.*, 2000; Eiter *et al.*, 2002] and fragments of dynamic logic programs [Alferes *et al.*, 1998; Alferes *et al.*, 2000].

In further work, semantic forgetting has been extended [Eiter *et al.*, 2006] to HEX-programs, which allow to combine logic programs with Description Logics [Eiter *et al.*, 2005] and, applied in defining a notion of forgetting for the Web Ontology Language (OWL).<sup>4</sup>

Furthermore, a system prototype for our semantic forgetting, called LPForget, which comprises two modules has been implemented and is available for experiments.<sup>5</sup> The module Forgetting serves for computing the result of forgetting about certain literals in a logic program under the answer set semantics; all algorithms introduced in this paper have been implemented in it. The other module, named CRS, facilitates conflict resolution (or preference recovery) in multi-agent systems along the approach in Section 6.1. In our system, once the constraints for different agents are specified, CRS will first check whether these constraints are consistent. If they are not consistent, the user can make compromises by forgetting about some literals such that an agreement is reached. The system can also make recommendations for the set of literals to be forgotten. However, the current recommendation algorithm is not optimized yet, and it remains to explore more efficient algorithms.

Several interesting issues remain for further research. One issue are more efficient implementations and improved algorithms for computing the result of forgetting. In particular, given a disjunctive logic program  $P$  and a literal  $l$ , Algorithm  $\text{forget}_3(P, l)$  outputs generally a nested logic program which represents the result of forgetting about  $l$  from  $P$  by its minimal answer sets. It is well-known that nested logic programs can be efficiently transformed into equivalent ordinary disjunctive programs if new symbols are allowed [Pearce *et al.*, 2001]. However, this does not carry over to minimal answer sets, and further minimization is needed. It would be interesting to see a syntax-based algorithm (i.e., based merely on program transformations but not on the actual answer sets) that outputs an ordinary disjunctive logic program on the original vocabulary as the result of forgetting.

Another issue is the application of forgetting in various scenarios of conflict resolving, such as ontology merging and alignment in the Semantic Web [Noy and Stuckenschmidt, 2005]. In these applications, both closed and open world reasoning are involved. Exploring a theory of forgetting in such a setting is an interesting issue.

Strong equivalence [Lifschitz *et al.*, 2001] has received a lot of attention in nonmonotonic logic programming because of its importance for program modularity and in applications like information integration. It would be interesting to introduce a notion of semantics forgetting that preserves strong equivalence, in line with the properties and the approach in this paper.

Finally, an extension of the approach in this paper to other semantics of nonmonotonic logic programming and more general formalisms, such as default logic or autoepistemic logic, is an intriguing issue.

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<sup>4</sup><http://www.w3.org/2004/OWL/>

<sup>5</sup>The LPForget website is <http://www.cit.gu.edu.au/~kewen/LPForget>.



## Acknowledgements

The authors would like to thank Fu-Leung Cheng, Esra Erdem, Paolo Ferraris, Fangzhen Lin, Abdul Sattar, Kaile Su, Rodney Topor and Yan Zhang for helpful comments and discussions.

## 8 Appendix

**Proposition 5** *Let  $\equiv_X$  be a equivalence relation on a collection of logic programs on  $Lit$  that is stronger than ordinary equivalence and invariant under literal extensions. Then forgetting does not preserve  $\equiv_X$ .*

*Proof.* Since  $\equiv_X$  is stronger than ordinary equivalence of logic programs, there must exist two programs  $P$  and  $P'$  such that  $P \equiv P'$  but  $P \not\equiv_X P'$ . Let  $l$  be a new literal that appears neither in  $P$  nor  $P'$ . Then  $\text{forget}(P, l) \equiv P$  and thus  $P'$  is also a result of forgetting about  $l$  in  $P$ . Obviously,  $P \equiv_X P'$  but their results of forgetting (i.e.  $P$  and  $P'$ ) are not equivalent under  $\equiv_X$ .  $\square$

**Proposition 8** *Let  $P$  be a consistent logic program and let  $F = \{l_1, \dots, l_m\}$  be a set of literals. Then  $\text{forget}(P, F) \equiv \text{forget}(\text{forget}(\text{forget}(P, l_1), l_2), \dots, l_m)$ .*

*Proof.* Assume that  $m > 1$ . Let  $M = \mathcal{AS}(\text{forget}(P, F))$ ,  $M' = \mathcal{AS}(\text{forget}(P, F'))$  and  $M'' = \mathcal{AS}(\text{forget}(\text{forget}(P, F'), l_m))$  where  $F' = \{l_1, \dots, l_{m-1}\}$ .

We claim that  $M = M''$ . To prove this, we first show that  $M \subseteq M''$ . Consider any set  $X \in M$ . Then, by an analogue of item 5 in Proposition 6, there exists a stable model  $S$  of  $P$  such that  $X = S - F$ . Now in the process of iterative construction, let

$$\begin{aligned} P_0 &= P, \\ P_i &= \text{forget}(P_{i-1}, l_i), \text{ for } i = 1, \dots, m. \end{aligned}$$

By item 5 of Proposition 6 and the fact that  $l_i$  does not occur in  $P_i$ , we can show by induction on  $i = 1, 2, \dots, m$  that there must exist some answer set  $S_i$  of  $P_i$  such that  $S_i \subseteq S_{i-1} - \{l_i\}$ . Consequently,  $S_m \subseteq S - \{l_1, \dots, l_m\} = S - F$ . Hence, there exists an answer set  $S_m$  of  $P_m$  such that  $S_m \subseteq X$ .

Furthermore,  $S_m \subset X$  is impossible. Otherwise,  $S_m \cap X \subset X$  would hold. Thus,  $S_m \cup F'$ , for some  $F' \subseteq F$ , is an answer set of  $P$ , and  $S_m \cup F' \subset S \setminus F$ . Hence,  $S$  is not an  $F$ -answer set of  $P$ , which is a contradiction. Thus, we have  $S_m = X$ . In conclusion,  $M \subseteq M''$ .

Conversely, suppose  $X \in M''$ . That is,  $X$  is an answer set of  $P_m$ . By the definition of one-literal forgetting, this means that there exists an answer set  $S$  of  $P$  such that  $S \setminus F = X$ . Towards a contradiction, suppose that  $X \notin M$ . Then there exists a set  $Y$  in  $M$  such that  $Y \subset X$ . As already shown,  $Y$  is an answer set of  $P_m$ . But this contradicts that  $X$  is an answer set of  $P_m$ , as  $X$  is not minimal. Hence,  $X \in M$ .  $\square$

**Theorem 3** *Let  $P$  be a consistent (disjunctive) logic program and let  $p \in Lit_P$  be an atom. Then  $X \subseteq Lit_P$  is an answer set of  $\text{forget}(P, p)$  iff  $X$  is a minimal model of  $\text{forget}(\text{lcomp}(P), p)$ . That is,*

$$\mathcal{AS}(\text{forget}(P, p)) = \text{MMod}(\text{forget}(\text{lcomp}(P), p)).$$

*Proof.* We use the following lemma.

**Lemma 3** *Let  $X \subseteq Lit_P \setminus \{p\}$  such that  $X \models \text{forget}(\text{lcomp}(P), p)$ . Then either  $X \in \mathcal{AS}(P)$  or  $X \cup \{p\} \in \mathcal{AS}(P)$ .*

**Proof of Lemma 3** By Theorem 1,  $X' \in \mathcal{AS}(P)$  iff  $X' \models \text{lcomp}(P)$  holds for each  $X' \subseteq \text{Lit}_P$ . Since  $\text{forget}(\text{lcomp}(P), p) = \text{lcomp}(P)(p/\text{true}) \vee \text{lcomp}(P)(p/\text{false})$ , for each  $X \subseteq \text{Lit}_P \setminus \{p\}$  such that  $X \models \text{forget}(\text{lcomp}(P), p)$  thus either  $X' = X \in \mathcal{AS}(P)$  (if  $X \models \text{lcomp}(P)(p/\text{false})$ ) or  $X' = X \cup \{p\} \in \mathcal{AS}(P)$  (if  $X \models \text{lcomp}(P)(p/\text{true})$ ) holds.  $\blacksquare$

(i)  $\mathcal{AS}(\text{forget}(P, p)) \subseteq \text{MMod}(\text{forget}(\text{lcomp}(P), p))$ : Let  $X \in \mathcal{AS}(\text{forget}(P, p))$ . Then, there exists some  $S \in \mathcal{AS}_p(P)$  such that  $X \sim_p S$ . By Theorem 1,  $S \models \text{lcomp}(P)$ . Since  $S \setminus \{p\} = X$ , we have  $X \models \text{forget}(\text{lcomp}(P), p)$ . To show that  $X$  is also minimal, assume towards a contradiction that some  $X' \subset X$  exists such that  $X' \models \text{forget}(\text{lcomp}(P), p)$ . By Lemma 3,  $S' \in \mathcal{AS}(P)$  holds for either  $S' = X'$  or  $S' = X' \cup \{p\}$ . In both cases,  $S' \subset_p S$ ; however, this contradicts  $S \in \mathcal{AS}_p(P)$ . This proves  $X \in \text{MMod}(\text{forget}(\text{lcomp}(P), p))$ .

(ii)  $\text{MMod}(\text{forget}(\text{lcomp}(P), p)) \subseteq \mathcal{AS}(\text{forget}(P, p))$ : Let  $X \in \text{MMod}(\text{forget}(\text{lcomp}(P), p))$ . By Lemma 3,  $S \in \mathcal{AS}(P)$  for either  $S = X$  or  $S = X \cup \{p\}$ . We show that in both cases  $S \in \mathcal{AS}_p(P)$ . Towards a contradiction, suppose that  $S \notin \mathcal{AS}_p(P)$ . Then there exists some  $S' \in \mathcal{AS}_p(P)$  such that  $S' \subset_p S$  and, by item 1 of Proposition 6,  $X' := S' \setminus \{p\} \in \mathcal{AS}(\text{forget}(P, p))$ . By part (i), it follows that  $X' \in \text{MMod}(\text{forget}(\text{lcomp}(P), p))$ ; Since  $X' \subset X$ , this contradicts  $X \in \text{MMod}(\text{forget}(\text{lcomp}(P), p))$ . Hence,  $S \in \mathcal{AS}_p(P)$ . By item 1 of Proposition 6,  $S \setminus \{p\} = X \in \mathcal{AS}(\text{forget}(P, p))$ . This proves the result.  $\square$

**Theorem 4** Let  $P$  be a consistent (disjunctive) logic program and let  $p \in \text{Lit}_P$  be an atom. Then  $X \subseteq \text{Lit}_P \setminus \{p\}$  is an answer set of  $\text{forget}(P, p)$  if and only if either  $X$  or  $X \cup \{p\}$  is a model of  $\text{Circ}(\text{lcomp}(P), p)$  (resp.,  $\text{ECWA}(\text{lcomp}(P), p)$ ).

*Proof.* By Theorem 1,  $\mathcal{AS}(P) = \{S \subseteq \text{Lit}_P \mid S \models \text{lcomp}(P)\}$ . Hence, from the characterization of models of  $\text{Circ}(\text{lcomp}(P), p)$  in terms of  $\subseteq_p$  and  $\sim_p$ , it is easily seen that  $\mathcal{AS}_p(P) = \{S \subseteq \text{Lit}_P \mid S \models \text{Circ}(\text{lcomp}(P), p)\}$ .

Consider  $X \subseteq \text{Lit}_P \setminus \{p\}$ . Suppose first that  $X \in \mathcal{AS}(\text{forget}(P, p))$ . Then  $X \sim_p S$  for some  $S \in \mathcal{AS}_p(P)$ , and either  $S = X$  or  $S = X \cup \{p\}$  must hold. As  $S \models \text{Circ}(\text{lcomp}(P), p)$ , the only-if direction of the theorem holds. Conversely, suppose that  $S \models \text{Circ}(\text{lcomp}(P), p)$  for either  $S = X$  or  $S = X \cup \{p\}$ . Then  $S \in \mathcal{AS}_p(P)$  and by item 1 of Proposition 6,  $S \setminus \{p\} = X \in \mathcal{AS}(\text{forget}(P, p))$ . This proves the result.  $\square$

**Theorem 6** For every consistent (disjunctive) logic program  $P$  and a literal  $l$ , Algorithm  $\text{forget}_2(P, l)$  outputs a correct representation of  $\text{forget}(P, l)$ . *Proof.* By the construction of  $P_1$  in Step 1,  $\mathcal{AS}(P_1) = \{X \in \mathcal{AS}(P) \mid l \in X\}$ . By Proposition 1,  $\mathcal{A}_1$  consists thus of all sets  $X$  such that  $l \notin X$  and  $X \cup \{l\} \in \mathcal{AS}_l(P)$ .

In Step 2, the constraint  $\leftarrow l$  guarantees that each answer set  $X$  of  $P_2$  is an answer set of  $P$  such that  $l \notin X$ . The constraint  $\leftarrow a_1, \dots, a_k$ , for  $M = \{a_1, \dots, a_k\} \in \mathcal{A}_1$ , enforces that  $M \cup \{l\} \not\subseteq X$  holds. Consequently,  $\mathcal{A}_2$  consists of all sets  $X \in \mathcal{AS}_l(P)$  such that  $l \notin X$ .

Combining the two cases,  $\mathcal{AS}_l(P) = \{X \cup \{l\} \mid X \in \mathcal{A}_1\} \cup \mathcal{A}_2$ ; thus by item 1 of Proposition 6,  $\mathcal{AS}(\text{forget}(P, l)) = \mathcal{A} = \mathcal{A}_1 \cup \mathcal{A}_2$ . It is easy to see that  $\mathcal{AS}(P') = \mathcal{A}$ ; this proves the result.  $\square$

**Proposition 11** Let  $P$  be a disjunctive program. If  $P'$  is obtained by the elimination of head redundancy from  $P$ , then  $P$  and  $P'$  have the same answer sets. *Proof.* Let  $P = P_0 \cup \{r\}$  with  $l \in \text{head}(r) \cap \text{body}^-(r)$  and let  $P' = P_0 \cup \{(head(r) - l) \leftarrow body(r)\}$ . We show that for every  $S \subseteq \text{Lit}_P$ , the following statement holds: (\*)  $S \in \mathcal{AS}(P)$  iff  $S \in \mathcal{AS}(P')$ . If  $l \in S$ , then  $P^S = (P')^S$  and therefore (\*) holds. Otherwise

(i.e.,  $l \notin S$ ), the rule  $r_1 = \text{head}(r) \leftarrow \text{body}^+(r)$  is in  $P^S$  iff the rule  $r_2 = (\text{head}(r) - l) \leftarrow \text{body}^+(r)$  is in  $(P')^S$ . Furthermore, for each  $S' \subseteq S$  it holds that  $S' \models r_1$  iff  $S' \models r_2$ , and therefore  $S' \models P^S$  iff  $S' \models (P')^S$ . Again it follows that (\*) holds.  $\square$

To prove Theorem 7, we first show a lemma.

**Lemma 4** *Let  $P$  be a DDLP and let  $S$  be an interpretation of  $P$ . If  $S \in \mathcal{AS}(\mathcal{T}(P))$ , then  $S$  is a minimal answer set of  $P$ .*

*Proof.* Let  $S \in \mathcal{AS}(\mathcal{T}(P))$ . We first show that  $S \in \mathcal{AS}(P)$ . Let  $r' \in P^S$  such that  $S \models \text{body}(r')$ . Then  $r'$  must be of the form  $\text{head}(r) \leftarrow \text{body}^+(r)$  for some  $r \in P$  such that  $\text{body}^-(r) \cap S = \emptyset$  and  $\text{body}^{--}(r) \subseteq S$ . Furthermore, the rule  $r'' = \text{head}(r) \leftarrow \text{body}^+(r), \text{not not } \text{body}^{--}(r)$  is in  $\mathcal{T}(P)^S$ . Now if  $\text{body}^+(r) \subseteq S$ , then  $S \models \text{head}(r)$  as  $S \models \text{body}^{--}(r)$ . Hence,  $S \models r'$ . It follows that  $S \models P^S$ .

Consider any  $S' \subseteq S$  such that  $S' \models P^S$ . Then  $S' \models \mathcal{T}(P)^S$ , as each rule in  $\mathcal{T}(P)^S$  results from a rule in  $P^S$  by adding literals in the body. As  $S \in \mathcal{AS}(\mathcal{T}(P))$ , it follows that  $S' = S$ . This proves  $S \in \mathcal{AS}(P)$ . It remains to show that  $S$  is minimal. Consider any  $S' \in \mathcal{AS}(P)$  such that  $S' \subseteq S$ . Then  $S' \models \mathcal{T}(P)^{S'}$ . Since  $\mathcal{T}(P)^S \subseteq \mathcal{T}(P)^{S'}$ ,  $S' \models \mathcal{T}(P)^S$ . As  $S \in \mathcal{AS}(\mathcal{T}(P))$ , it follows that  $S' = S$ .  $\square$

**Theorem 7** *For every N-acyclic DDLP  $P$ , it holds that  $\mathcal{AS}(P) = \mathcal{AS}(\mathcal{T}(P))$ .*

*Proof.*  $\mathcal{AS}(\mathcal{T}(P)) \subseteq \mathcal{AS}(P)$ : Immediate from Lemma 4.

$\mathcal{AS}(P) \subseteq \mathcal{AS}(\mathcal{T}(P))$ : Suppose  $S \in \mathcal{AS}(P)$  but  $S \notin \mathcal{AS}(\mathcal{T}(P))$ . Then  $S$  is not a minimal model of  $\mathcal{T}(P)^S$ . Note that  $S \models \mathcal{T}(P)^S$ . Hence, there exists some interpretation  $S' \subset S$  such that  $S' \models \mathcal{T}(P)^S$ . We show that then some interpretation  $S''$  exists such that  $S \subseteq S'' \subset S$  and  $S'' \models P^S$ ; this means that  $S$  is not a minimal model of  $P^S$  and thus contradicts that  $S \in \mathcal{AS}(P)$ , proving the result.

We first assume that  $P$  is normal, i.e.,  $|\text{head}(r)| \leq 1$  for each  $r \in P$ . Let  $L$  be a level mapping witnessing the N-acyclicity of  $P$  and let  $X = \{l \in S \setminus S' \mid \forall l' \in S \setminus S' : L(l) \leq L(l')\}$  be the set of all literals in  $S \setminus S'$  having the smallest level. Since  $S' \subset S$ ,  $X \neq \emptyset$ . We claim that  $S'' := S \setminus X \models P^S$ .

Towards a contradiction, suppose that  $S'' \not\models r$  for some  $r \in P^S$ . Since  $S \models P^S$ ,  $r$  is of the form  $l \leftarrow \text{body}^+(r)$  where  $l \in X$  and  $\text{body}^+(r) \subseteq S''$ . Since  $r$  stems from a rule  $r' \in P$  such that  $S \models r'$  (i.e.,  $r = \text{head}(r') \leftarrow \text{body}^+(r')$ ) and  $S' \models \mathcal{T}(P)^S$ , it follows that either (a)  $\text{body}^+(r) \setminus S' \neq \emptyset$  or (b)  $\text{body}^{--}(r') \setminus S' \neq \emptyset$ . The minimality of  $L(l)$  implies that (a) is the case and that  $\text{body}^+(r) \cap X \neq \emptyset$ , and thus  $\text{body}^+(r) \not\subseteq S''$ ; this is a contradiction. This proves the claim  $S'' \models P^S$ , which contradicts that  $S \in \mathcal{AS}(P)$ . Hence, the result for normal programs  $P$  is proved.

The result for arbitrary programs is an easy consequence of this result and the following two facts. Call a program  $P'$  a *split* of a DDLP  $P$ , if  $P'$  results from  $P$  by replacing every rule  $r$  of form  $l_1 \vee \dots \vee l_k \leftarrow B$ ,  $k \geq 2$ , by at least one of the normal rules  $l_1 \leftarrow B, \dots, l_k \leftarrow B$ . Now,

*Fact A.* For every  $S \in \mathcal{AS}(P)$  there exists a split  $P'$  of  $P$  such that  $S \in \mathcal{AS}(P')$ .

Indeed, every split  $P'$  of  $P$  such that  $r$  is replaced by all rules  $l_i \leftarrow B$  with  $l_i \in S$  whenever  $\text{head}(r) \cap S \neq \emptyset$ , has  $S$  as an answer set. (This is a simple extension of a folklore result for ordinary logic programs.)

*Fact B.* If a DDLP  $P$  is N-acyclic, then every split  $P'$  of  $P$  is N-acyclic.  $\square$

**Theorem 8** *Let  $P$  be a consistent disjunctive logic program and let  $l \in \text{Lit}_P$  be a literal. Then  $\text{forget}_3(P, l)$  correctly represents  $\text{forget}(P, l)$ , i.e.,  $X$  is an answer set of  $\text{forget}(P, l)$  iff  $X$  is a minimal answer set of  $N'$ .*

*Proof.* By Lemma 1,  $P$  can be transformed into an equivalent negative program  $N_0$ . Furthermore, an easy extension of Corollary 2 implies that  $N_0$  is equivalent to  $N$ . Hence,  $P \equiv N$  and  $\text{forget}(P, l) \equiv \text{forget}(N, l)$ .

Let  $\mathcal{AS}_{min}(N')$  denote the minimal answer sets of program  $N'$  (w.r.t.  $\subseteq$ ). To prove the theorem, it is now by item 1 of Proposition 6 sufficient to prove the following claim:

*Claim.* For every set  $X'$  of literals such that  $l \notin X'$ ,  $X' \in \mathcal{AS}_{min}(N')$  iff  $X' = X \setminus \{l\}$  for some  $X \in \mathcal{AS}_l(N)$ .

Due to the elimination of tautology and the elimination of head redundancy,  $head(r) \cap body(r) = \emptyset$  for each rule  $r \in N$ . Thus the program  $N$  can be split into three disjoint parts:  $N = N_1 \cup N_2 \cup N_3$  where  $N_1$  consists of rules in  $N$  in which  $l$  does not appear;  $N_2 = \{r \in N \mid l \in head(r), l \notin body^-(r)\}$ ; and  $N_3 = \{r \in N \mid l \notin head(r), l \in body^-(r)\}$ . Notice that Step 3 in Algorithm  $forget_3(P, l)$  is performed only on the rules in  $N_3$ . Let the program  $N'_3$  result from  $N_3$  by the transformations in Step 3. Then  $N' = N_1 \cup N'_3$ . Let  $D_1, D_2, \dots, D_s$  denote all possible conjunctions constructed from  $l_i$  in Step 3; note that  $N_2 = \{r_1, \dots, r_n\}$ . We consider the three cases in Step 3.

(3.1) If  $n = 0$ , then no rule in  $N$  has  $l$  in the head. Thus  $l$  is false in every answer set of  $N$ , and hence  $N \equiv N'$  and  $\mathcal{AS}_l(N) = \mathcal{AS}(N)$ . Since  $N'$  is an ordinary program,  $\mathcal{AS}_{min}(N') = \mathcal{AS}(N')$ . It follows that  $\mathcal{AS}_{min}(N') = \mathcal{AS}_l(N)$  and the claim holds.

(3.2) If  $n > 0$  and  $m_j = 0$  for some  $j \in \{1, \dots, n\}$ , then the rule  $l \leftarrow$  is in  $N$ . Hence, every answer set of  $N$  contains  $l$ , and clearly  $X' \in \mathcal{AS}(N')$  iff  $X' \cup \{l\} \in \mathcal{AS}(N)$  holds for every set of literals  $X'$  with  $l \notin X'$ . As in the previous case,  $\mathcal{AS}_l(N) = \mathcal{AS}(N)$  and  $\mathcal{AS}_{min}(N') = \mathcal{AS}(N')$ , and the claim holds.

(3.1)  $n \geq 1$  and  $m_i \geq 1$  for every  $i = 1, \dots, n$ . We use the following lemmas.

**Lemma 5** *If  $l \notin X'$  and  $X' \models D_{i_0}$  for some  $i_0$  with  $1 \leq i_0 \leq s$ , then  $(N_3)^{X'} = (N'_3)^{X'}$ .*

**Lemma 6** *For every  $X' \in \mathcal{AS}_{min}(N')$ ,*

(1) *if  $X' \models D_{i_0}$  for some  $i_0$ ,  $1 \leq i_0 \leq s$ , then  $X' \in \mathcal{AS}(N)$ .*

(2) *if  $X' \not\models D_i$  for all  $i = 1, \dots, s$ , then  $X' \cup \{l\} \in \mathcal{AS}(N)$ .*

**Proof of Lemma 6** (1)  $N^{X'} = (N_1)^{X'} \cup (N_3)^{X'}$  since  $X' \models D_{i_0}$  implies  $l \notin X'$ . By Lemma 5,  $(N_3)^{X'} = (N'_3)^{X'}$  and thus  $N^{X'} = (N')^{X'}$ . So  $X' \models N^{X'}$ .

If  $X'' \subseteq X'$  and  $X'' \models N^{X'}$ , then  $X'' \models (N')^{X'}$ . Since  $X'$  is a minimal model of  $N^{X'}$ ,  $X'' = X'$ . Therefore,  $X' \in \mathcal{AS}(N)$ .

(2) Let  $X = X' \cup \{l\}$ . Since  $l \in X$  and  $l$  does not appear in  $N_1$ , we have  $N^X = (N_1 \cup N_2 \cup N_3)^X = (N_1)^X \cup (N_2)^X = (N_1)^{X'} \cup (N_2)^X = (N_1)^{X'} \cup \{l \leftarrow\}$ . By the assumption,  $X' \models (N_1)^{X'}$  and thus  $X \models (N_1)^{X'}$ . Obviously,  $X \models (N_2)^X$  since  $l \in X$  and every rule in  $N_2$  has head  $l$ . Thus,  $X \models N^X$ .

Now suppose that  $Y \subseteq X$  and  $Y \models N^X$ . Since  $l \leftarrow$  is in  $N^X$ ,  $l \in Y$ . Let  $Y = Y' \cup \{l\}$  where  $l \notin Y'$ . Then  $Y' \models (N_1)^{X'}$ . By  $(N')^{X'} = (N_1)^{X'}$ , we have  $Y' \models (N')^{X'}$ . Thus,  $Y' = X'$  by  $Y' \subseteq X'$  and the minimality of  $X'$ . This implies  $Y = X$ . So  $X$  is a minimal model of  $N^X$ , and thus  $X \in \mathcal{AS}(N)$ . ■

We now prove the claim.

( $\Rightarrow$ ): Let  $X' \in \mathcal{AS}_l(N)$ . Then  $X \in \mathcal{AS}(N)$  for either  $X = X'$  or  $X = X' \cup \{l\}$ , where  $l \notin X$ . Consider two cases:

Case 1.  $X = X'$ : Note that  $X' \models D_{i_0}$  for some  $i_0$ ,  $1 \leq i_0 \leq s$ , since  $l \notin X'$ . By Lemma 5,  $(N_3)^{X'} = (N'_3)^{X'}$ . Therefore,  $(N')^{X'} = (N_1)^{X'} \cup (N'_3)^{X'} = (N_1)^{X'} \cup (N_3)^{X'}$ . This implies  $X' \models (N')^{X'}$ .

Suppose now  $X'' \subseteq X'$  and  $X'' \models (N')^{X'}$ . Then  $X'' \models N^{X'}$ . Since  $X'$  is a minimal model of  $N^{X'}$ , we have  $X'' = X'$ . Thus,  $X' \in \mathcal{AS}(N')$ .

Suppose that  $Y \in \mathcal{AS}_{min}(N')$  and  $Y \subseteq X'$ . If  $Y \models D_{j_0}$  for some  $j_0$  ( $1 \leq j_0 \leq s$ ), then by Lemma 6  $Y \in \mathcal{AS}(N)$  and thus it follows that  $Y = X'$ . So  $X' \in \mathcal{AS}_{min}(N')$ . If  $Y \not\models D_i$  for all  $i$  ( $1 \leq j_0 \leq s$ ), then  $Y \cup \{l\} \in \mathcal{AS}(N)$  again by Lemma 6. This implies  $X' \notin \mathcal{AS}_l(N)$ , a contradiction. Therefore,  $X' \in \mathcal{AS}_{min}(N')$ .

Case 2.  $X = X' \cup \{l\}$ : Then  $N^X = N^{X' \cup \{l\}} = (N_1)^{X'} \cup \{l \leftarrow\}$ . Since  $l \in X$ , we have  $X \not\models D_i$  for every  $i$ ,  $1 \leq i \leq s$ . From  $l \notin X'$ , it follows that  $X' \not\models D_i$  for every  $i$ ,  $1 \leq i \leq s$ . Thus,  $(N')^{X'} = (N_1)^{X'} \cup (N'_3)^{X'} = (N_1)^{X'}$ . Since  $X \models N^X$ ,  $X' \models (N_1)^{X'}$  and thus  $X' \models (N')^{X'}$ .

If  $X'' \subseteq X'$  and  $X'' \models (N')^{X'}$ , then  $X'' \cup \{l\} \models N^X$ . By the minimality of  $X$ ,  $X'' \cup \{l\} = X$ . Since  $l \notin X'$  and  $l \notin X''$ , we have  $X'' = X'$ . Thus,  $X' \in \mathcal{AS}(N')$ .

If  $Y \in \mathcal{AS}_{min}(N')$  and  $Y \subseteq X'$ , then  $Y \not\models D_i$  for all  $i = 1, \dots, s$ . Consequently,  $Y \cup \{l\} \in \mathcal{AS}(N)$ . So  $Y \cup \{l\} \subseteq X' \cup \{l\}$ . This means  $Y = X'$ , since the answer sets of  $N$  are incomparable under  $\subseteq$ . Thus  $X' \in \mathcal{AS}_{min}(N')$ .

( $\Leftarrow$ ): Suppose that  $X' \in \mathcal{AS}_{min}(N')$ . Consider two possible cases:

Case 1.  $X' \not\models D_i$  for all  $i = 1, \dots, s$ : By Lemma 6,  $X = X' \cup \{l\} \in \mathcal{AS}(N)$ . Thus  $X' \in \mathcal{AS}_l(N)$  since  $l \in X$  by item 3 of Proposition 1.

Case 2.  $X' \models D_{i_0}$  for some  $i_0$  with  $1 \leq i_0 \leq s$ : By Lemma 6,  $X' \in \mathcal{AS}(N)$ . Consider any  $Y \in \mathcal{AS}(N)$ ,  $Y \neq X'$ , such that  $Y \setminus \{l\} \subseteq X'$ . Then  $l \in Y$  since  $l \notin X'$  and answer sets of  $N$  are incomparable. Thus  $Y \in \mathcal{AS}_l(N)$  by Proposition 1. From the proof of the only-if part,  $Y \setminus \{l\}$  is a (minimal) answer set of  $N'$ . By the minimality of  $X'$ ,  $Y \setminus \{l\} = X'$ . This proves  $X' \in \mathcal{AS}_l(N)$ .

**Proposition 13** *Let  $P$  be a consistent normal logic program and let  $l \in Lit_P$  be a literal. If no negative cycle of  $DG_P$  contains  $l$ , then  $\text{forget}_3(P, l)$  outputs an  $N$ -acyclic program.*

*Proof.* The result is a consequence of the fact that for a program  $P$  as described, also in the dependency graph  $DG_{N_0}$  for the program  $N_0$  from Step 1 of  $\text{forget}_3(P, l)$  no negative cycle will contain  $l$ . This is because each of the transformations in  $\mathbf{T}^*$  preserves the property that if a literal  $l$  does not occur on a negative cycle of  $DG_{P_0}$  of the original program  $P_0$ , then it does not occur on a negative cycle of  $DG_{P_1}$  in the transformed program  $P_1$ . Indeed, each transformation except Unfolding only removes edges; Unfolding can add only some positive edges from  $l$  to  $l'$  such that positive edges from  $l$  to  $l''$  and from  $l''$  to  $l'$  exist, for some literal  $l''$ , and some negative edges from  $l$  to  $l'$  such that a positive edge from  $l$  to  $l''$  and a negative edge from  $l''$  to  $l'$  exist, for some literal  $l''$ . Thus, occurrence of  $l$  in a negative cycle of  $DG_{P_1}$  implies occurrence of  $l$  in a negative cycle of  $DG_{P_0}$ .

As a consequence, in Step 3 (where  $N = N_0$ ) no replacements of *not*  $l$  by  $D_i$  will be performed that can violate the condition (ii) of  $N$ -acyclicity; condition (i) is vacuously true.  $\square$

**Theorem 9** *Given a consistent (disjunctive) logic program  $P$ , a literal  $l$ , and a set of literals  $X$ , deciding whether  $X$  is an  $l$ -answer set of  $P$  is  $\Pi_2^p$ -complete. *Proof.* Deciding whether  $X \notin \mathcal{AS}_l(P)$  can be done in NP time using an NP oracle: we must show that either (1)  $X \notin \mathcal{AS}(P)$  (which is in co-NP, cf. [Eiter and Gottlob, 1995b]), or that (2) there exists some  $X' \in \mathcal{AS}(P)$  such that  $X' \subset_l X$ ; such an  $X'$  can be guessed and checked using an NP-oracle in polynomial time. Consequently,  $l$ -answer set checking is in  $\text{co-NP}^{\text{NP}} = \Pi_2^p$ .*

The hardness result is shown by a reduction from deciding whether a given logic program  $P$  (without strong negations) has no answer set, which is  $\Pi_2^p$ -complete [Eiter and Gottlob, 1995b].

In fact, given a (disjunctive) logic program  $P$ , construct a logic program  $P' = \{\text{head}(r) \leftarrow p, \text{body}(r) \mid r \in P\} \cup \{q \leftarrow \text{not } p. p \leftarrow \text{not } q\} \cup \{a \leftarrow \text{not } p \mid a \text{ appears in } P\}$ , where  $p$  and  $q$

are two fresh atoms. This program  $P'$  has one answer set  $X_0$  in which  $p$  is false and all other atoms are true; all other answer sets are of the form  $X \cup \{p\}$ , where  $X \in \mathcal{AS}(P)$ . It holds that  $X_0 \in \mathcal{AS}_p(P')$  iff  $P$  has no answer set.  $\square$

**Theorem 10** *Given a consistent normal logic program  $P$ , a literal  $l$ , and a set of literals  $X$ , deciding whether  $X$  is an  $l$ -answer set of  $P$  is co-NP-complete. Proof.* Similar to the proof of Theorem 9, in order to show that  $X \notin \mathcal{AS}_l(P)$ , we must that either (1)  $X \notin \mathcal{AS}(P)$  (which can be tested in polynomial time), or that (2) there exists some  $X' \in \mathcal{AS}(P)$  such that  $X' \subset_l X$ . Such an  $X'$  can be guessed and checked in polynomial time. Hence, deciding  $X \notin \mathcal{AS}_l(P)$  is in NP, which implies that  $l$ -answer set checking is in co-NP.

The hardness result is shown by the reduction in Theorem 9. In fact, note that if we consider only normal programs  $P$  (without strong negation), then  $P'$  is also a normal program. The problem of deciding whether  $P$  has no answer set, which is co-NP-complete, is thus reduced to deciding whether  $X_0$  is a  $p$ -answer set of  $P'$ .  $\square$

**Theorem 11** *Given a consistent negative program  $P$ , a literal  $l$ , and a set of literals  $X$ , deciding whether  $X$  is an  $l$ -answer set of  $P$  is co-NP-complete. Proof.* The co-NP membership follows from that fact that for any given set of literals  $X$  and negative program  $P$ , deciding whether  $X \in \mathcal{AS}(P)$  is polynomial. (Indeed,  $X \in \mathcal{AS}(P)$  iff  $X \models P^X$  and  $X \setminus \{l\} \not\models P^X$ , for every  $l \in X$ .) Thus, testing that (1)  $X \in \mathcal{AS}(P)$  and that (2) there is no  $X' \in \mathcal{AS}(P)$  such that  $X' \subset_l X$  is feasible in co-NP.

As for co-NP-hardness, let  $C = C_1 \wedge \dots \wedge C_k$  be a propositional CNF over atoms  $y_1, \dots, y_m$ , where each  $C_j$  is non-empty. Define

$$N = \bigcup_{i=1}^m (N_i \cup \{ \leftarrow \text{not } C'_i \}) \cup \{ l \leftarrow \text{not } y_1, l \leftarrow \text{not } y'_1 \},$$

where

$$\begin{aligned} N_i &= \{ y_i \leftarrow \text{not } y'_i, y'_i \leftarrow \text{not } y_i, y_i \leftarrow \text{not } l, y'_i \leftarrow \text{not } l \}, \quad 1 \leq i \leq m, \\ C'_j &= \{ y_i \mid y_i \in C_j \} \cup \{ y'_i \mid \neg y_i \in C_j \}, \quad 1 \leq j \leq k. \end{aligned}$$

Clearly, the satisfying assignments of  $C$  correspond one-to-one to the answer sets of  $N$  containing  $l$ . Furthermore, the set  $X = \{ y_i, y'_i \mid 1 \leq i \leq m \}$  is an answer set of  $N$ . It holds that  $X$  is also an  $l$ -answer set of  $N$  iff  $C$  is unsatisfiable, which establishes the co-NP-hardness.  $\square$

**Theorem 12** *Given a consistent (disjunctive) logic program  $P$  and literals  $l$  and  $l'$ , deciding whether  $\text{forget}(P, l) \models_c l'$  is  $\Sigma_3^p$ -complete. Proof.* Given a logic program  $P$  and two literals  $l$  and  $l'$ ,  $\text{forget}(P, l) \models_c l'$  holds iff there exists some  $S \in \mathcal{AS}_l(P)$  such that  $l' \in S$ . Since deciding  $S \in \mathcal{AS}_l(P)$  is in  $\Pi_2^p$ , the problem thus is in  $\Sigma_3^p$ .

The  $\Sigma_3^p$ -hardness can be shown by an encoding of quantified Boolean formulas (QBFs) of the form  $\exists Z \forall X \exists Y C$ . It is well-known that the problem of deciding whether a disjunctive logic program has no answer sets is  $\Pi_2^p$ -complete [Eiter and Gottlob, 1995b]. As shown there, for every QBF  $F = \forall X \exists Y C$  there exists a polynomial-time constructible logic program  $P_F$  which has no answer set iff  $F$  is true. We extend this program to an encoding of a QBF  $F' = \exists Z \forall X \exists Y C$  into credulous inference under forgetting as follows.

First,  $P_F$  can be easily extended to a program  $P_{F[Z]}$  encoding a QBF  $F[Z] = \forall X \exists Y C[Z]$  with free variables (i.e., parameters)  $Z$ , where all new atoms in  $P_{F[Z]}$  are from  $Z$  and do not occur in rule heads, such that for every truth assignment  $\tau$  to  $Z$ , the program  $P_{F[Z/\tau(Z)]} = P_{F[Z]} \cup \{z \leftarrow . \mid z \in Z, \tau(Z) = \text{true}\}$  has no answer set iff the QBF  $F[Z/\tau(Z)]$  evaluates to true, where  $F[Z/\tau(Z)]$  results from  $F[Z]$  by replacing every  $z \in Z$  with  $\top$  if  $\tau(z) = \text{true}$  and with  $\text{F}$  if  $\tau(z) = \text{false}$ .

More in detail, suppose without loss of generality that  $C[Z] = C_1 \wedge \dots \wedge C_r$  where each  $C_i = \ell_{i,1} \vee \ell_{i,2} \vee \ell_{i,3}$  is a disjunction of literals  $\ell_{i,j}$  over atoms  $X \cup Y \cup Z$ . Let  $X' = \{x' \mid x \in X\}$ ,  $Y' = \{y' \mid y \in Y\}$ , and  $w$  be new atoms. Then  $P_{F[Z]}$  consists of the rules

$$\begin{array}{ll} x \vee x' \leftarrow . & \text{for each } x \in X, \\ y \vee y' \leftarrow . \quad y \leftarrow w. \quad y' \leftarrow w. \quad w \leftarrow y, y'. & \text{for each } y \in Y, \\ w \leftarrow \sigma(\neg.l_{k,1}), \sigma(\neg.l_{k,2}), \sigma(\neg.l_{k,3}). & \text{for each } k = 1, \dots, r, \\ w \leftarrow \text{not } w. & \end{array}$$

where  $\sigma$  maps classical literals  $\ell$  to classical and *not* literals as follows:

$$\sigma(\ell) = \begin{cases} x' & \text{if } \ell = \neg x \text{ for some } x \in X \\ y' & \text{if } \ell = \neg y \text{ for some } y \in Y \\ \text{not } z & \text{if } \ell = \neg z \text{ for some } z \in Z \\ \ell & \text{otherwise} \end{cases}$$

Notice that the only *not* literals occurring in  $P_{F[Z]}$  are *not*  $w$  in the last rule and the literals *not*  $z$  in rule bodies; furthermore, atoms from  $Z$  occur only in rule bodies. For void  $Z$ , the program  $P_{F[Z]}$  amounts to the program  $P_F$  in [Eiter and Gottlob, 1995b, Proof of Theorem 3] (after converting the  $\forall\exists$ -QBF into the  $\exists\forall$ -QBF used there). A simple extension of the proof of Theorem 3 in [Eiter and Gottlob, 1995b] gives the following lemma.

**Lemma 7** *For each truth assignment  $\tau$  to  $Z$ ,  $\forall X \exists Y C[Z/\tau(Z)]$  evaluates to true iff the program  $P_{F[Z/\tau(Z)]}$  has no answer set.*

Let  $Z' = \{z' \mid z \in Z\}$ ,  $l$ , and  $l'$ , be fresh atoms, and let  $P$  be the logic program obtained from  $P_{F[Z]}$  by adding  $l$  to the body of each rule and the following further rules:

- (1)  $l \leftarrow \text{not } l'. \quad l' \leftarrow \text{not } l.$
- (2)  $z \leftarrow \text{not } z'. \quad z' \leftarrow \text{not } z,$  for each  $z \in Z$ ,
- (3)  $x \leftarrow \text{not } l. \quad x' \leftarrow \text{not } l. \quad y \leftarrow \text{not } l. \quad y' \leftarrow \text{not } l,$  for each  $x \in X, y \in Y$ , and
- (4) the rule  $w \leftarrow \text{not } l.$

Informally, the rules in (1) select one of  $l$  and  $l'$ , and the rules in (2) select a truth assignment  $\tau$  to  $Z$ . If  $l$  is selected, then the program  $P_{F[Z]}$  is activated and evaluated for the selected  $\tau$ , while the rules (3) and (4) are discarded. The evaluation will lead to some answer set  $S_\tau^l$ , iff the program  $P_{F[Z/\tau(Z)]}$  has some answer set; note that every such  $S$  contains  $l$ . On the other hand, if  $l'$  is selected, then only the rules in (2)-(4) are active. The truth assignment  $\tau$  selected in (2) will be complemented with  $X \cup X' \cup Y \cup Y' \cup \{w\}$  to a unique answer set  $S_\tau^{l'}$  of the program  $P$ .

Now  $S_\tau^l \subset_l S_\tau^{l'}$  holds for every  $S_\tau^l$ . Furthermore, answer sets  $S$  and  $S'$  corresponding to different truth assignments  $\tau$  and  $\tau'$ , respectively, are always incomparable w.r.t.  $\subseteq$  on  $Z \cup Z'$ , and thus also w.r.t.  $\subseteq_l$ .

Therefore, for every  $\tau$ ,  $S_\tau^{l'}$  is an  $l$ -answer set of  $P$  iff no answer set  $S_\tau^l$  exists, i.e.,  $P_{F[Z/\tau(Z)]}$  has no answer set, which by Lemma 7 equals that  $\forall X \exists Y C[Z/\tau(Z)]$  evaluates to true. Hence, for some  $\tau$ ,  $S_\tau^{l'}$  is an  $l$ -answer set of  $P$  iff  $F' = \exists Z \forall X \exists Y C$  evaluates to true. Since every  $l$ -answer set of  $P$  that contains  $l'$  is of the form  $S_\tau^{l'}$  for  $\tau$ , it follows from Proposition 6 that  $\text{forget}(P, l) \models_c l'$  iff  $F'$  evaluates to true. Since  $P$  is constructible in polynomial time from  $F'$ , the  $\Sigma_3^p$ -hardness is proved.  $\square$

**Theorem 13** *Given a consistent negative program  $P$  and literals  $l$  and  $l'$ , deciding whether  $\text{forget}(N, l) \models_c l'$  is  $\Sigma_2^p$ -complete. Proof.* By Theorem 11, a guess for some  $l$ -answer set  $X$  of  $P$  such that  $l' \in X$  can be verified with an NP oracle in polynomial time. Hence, deciding  $\text{forget}(P, l) \models_c l'$  is in  $\Sigma_2^p$ .

As for  $\Sigma_2^p$ -hardness, take a QBF  $\exists X \forall Z E$ , where  $E = \bigvee_{i=1}^k D_i$  is a DNF on  $X \cup Z$  such that without loss of generality in each disjunct  $D_i$  some variable from  $Z$  occurs. Construct the same program as above in Theorem 11 for  $C = \neg E$  and where  $Y = X \cup Z$  and  $y_1$  is an arbitrary variable from  $Z$ , but (1) omit the clauses  $x_i \leftarrow \text{not } l$  and  $x'_i \leftarrow \text{not } l$ , and (2) add a clause  $l' \leftarrow \text{not } l$ , where  $l'$  is a fresh literal. Then for each set  $X' \subseteq X$ , the set

$$S_{X'} = X' \cup \{x'_i \mid x_i \in X \setminus X'\} \cup Z \cup \{z'_j \mid z_j \in Z\} \cup \{l'\}$$

is an answer set of  $N$ . The sets  $S_{X'}$  are also all answer sets of  $N$  that contain  $l'$  (and do not contain  $l$ ). Furthermore,  $S_{X'}$  is an  $l$ -answer set of  $N$  iff there exists no satisfying truth assignment for  $C (= \neg E)$  which corresponds on  $X$  to  $S_{X'}$  in the obvious way. In summary, this means that  $N$  has  $l$ -answer set in which  $l'$  is true, i.e.,  $\text{forget}(N, l) \models_c l'$ , iff the formula  $\exists X \forall Z E$  evaluates to true.  $\square$

**Theorem 14** *Given a consistent normal program  $N$  and literals  $l$  and  $l'$ , deciding whether  $\text{forget}(N, l) \models_c l'$  is  $\Sigma_2^p$ -complete. Proof.* It is sufficient to note that the program  $N$  constructed in the proof of Theorem 13 is normal.  $\square$

**Theorem 15** *Given a consistent logic program  $P$  and literals  $l$  and  $l'$ , deciding whether  $\text{forget}(P, l) \models_s l'$  is (i)  $\Pi_2^p$ -complete for arbitrary disjunctive logic programs  $P$ , and (ii) co-NP-complete for normal logic programs and for negative logic programs  $P$ . Proof.* By Theorem 7, to decide  $\text{forget}(P, l) \models_s l'$ , we need only to decide  $P \models_s l'$ . The latter is  $\Pi_2^p$ -complete for logic programs [Eiter and Gottlob, 1995b] and co-NP complete for normal/negative programs; for both cases, membership in co-NP follows since testing  $X \in \mathcal{AS}(P)$  is polynomial (cf. proof of Theorem 11 for negative programs), and co-NP hardness from the results in [Marek and Truszczyński, 1991] (incorporating the consistency requirement is easy).  $\square$

**Proposition 15** *Let  $P$  be a logic program and let  $F$  be a consistent set of literals. Suppose that (1) no literal in  $F$  occurs in a rule body in  $P$ , and (2) for each rule  $r$ , either no or every literal in  $\text{head}(r)$  is in  $F$ . Then  $\text{forget}(P, F) = P \setminus R(F)$ .*

We first provide two lemmas, which are straightforward corollaries of the well-known, more general Splitting Set Theorem in [Lifschitz and Turner, 1994] and elementary properties of answer set semantics. To avoid introducing the necessary notions for that result, we provide for self-containedness simple genuine proofs.

**Lemma 8** *Let  $P$ ,  $F$  and  $Q = P \setminus R(F)$  be given as in Proposition 15. If  $S'$  is an answer set of  $Q$ , then there exists a subset  $Y'$  of  $F$  such that  $S' \cup Y'$  is an answer set of  $P$ .*



Since we have a very special case here, Lemma 8 also allows a simple proof.

*Proof.* Suppose that  $S'$  is an answer set of  $Q$ . Denote  $D' = \{head(r) \mid r \in R(F), S' \models body(r)\}$ . Then every literal in  $D'$  must be in  $F$ . Let  $Y'$  be a minimal model of  $D'$  and  $S = S' \cup Y'$ . By the assumption of  $F$ ,  $S' \cap Y' = \emptyset$  and thus  $S$  is consistent. We show that  $S$  is an answer set of  $P$ .

First, since  $P^S = P^{S'} = Q^{S'} \cup (R(F))^{S'}$ , we have  $S \models P^S$ . Next, suppose  $X \subseteq S$  and  $X \models P^S$ . Take  $X' = X \setminus F$ ; then  $X' \subseteq S'$ . We can see that  $X' \models Q^{S'}$  and thus  $X' = S'$ . Notice that  $Y' = S \cap F$  is a minimal model of  $D'$ , and that  $X \cap F \models D'$  implies  $X \cap F = S \cap F$ . Thus,  $X = S$ .  $\square$

**Lemma 9** *Let  $P$ ,  $F$  and  $Q = P \setminus R(F)$  be given as in Proposition 15. If  $S$  is an answer set of  $P$ , then  $S \setminus F$  is an answer set of  $Q$ .*

*Proof.* By  $P^S = P^{S'}$  and  $Q \subseteq P$ , it is easy to see that  $S' \models Q^{S'}$  where  $S' = S \setminus F$ .

Suppose that  $Y' \subseteq S'$  and  $Y' \models Q^{S'}$ . Then  $Y \models P^{S'} = P^S$  where  $Y = Y' \cup (S \cap F)$ . Thus  $Y = S$ . Since  $Y' = Y \cap \bar{F}$  and  $S' = S \cap \bar{F}$ , we have  $Y' = S'$ . Here  $\bar{F} = Lit_P \setminus F$ .

Therefore,  $S \setminus F$  is an answer set of  $Q$ .  $\square$

**Proof of Proposition 15** Denote  $Q = P \setminus R(F)$ . Suppose that  $S'$  is an answer set of  $Q$ . By Lemma 8, there exists a subset  $Y'$  of  $F$  such that  $S = S' \cup Y'$  is an answer set of  $P$ .

To show that  $S'$  is an answer set of  $forget(P, F)$ , it suffices to prove that  $S$  is an  $F$ -answer set of  $P$ . In fact, if  $Z$  is an answer set of  $P$  such that  $Z \subseteq_F S$ , then  $Z' \subseteq S'$  where  $Z' = Z \setminus F$ . By Lemma 9,  $Z'$  is an answer set of  $Q$ . Thus  $S' = Z'$ , which means  $Z \sim_F S$ . That is,  $S$  is an  $F$ -answer set of  $P$ .

On the other hand, if  $S'$  is an answer set of  $forget(P, F)$ , then there exists an  $F$ -answer set  $S$  of  $P$  such that  $S' = S \setminus F$ . By Lemma 9,  $S'$  is also an answer set of  $Q$ .  $\blacksquare$

We are now ready to prove Theorem 16.

**Theorem 16** *Let  $(P, <)$  be an inheritance program and let  $S$  be a set of literals. Then  $S$  is an inheritance answer set of  $(P, <)$  iff  $S$  is an answer set of  $forget(P', F)$  where  $P'$  is obtained as above.*

*Proof.*  $S$  is an inheritance answer set of  $(P, <)$  iff  $S$  is a minimal model of  $(P, <)^S$  iff  $S$  is an answer set of  $P \setminus R(F)$  iff  $S$  is an answer set of  $forget(P', F)$ , by Proposition 15.  $\square$

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