2000

Fages' Theorem and Answer Set Programming

Yuliya Lierler  
*University of Nebraska at Omaha, ylierler@unomaha.edu*

Esta Erdem  
*Sabancı University*

Vladimir Lifschitz  
*University of Texas at Austin*

Follow this and additional works at: [https://digitalcommons.unomaha.edu/compsicfacproc](https://digitalcommons.unomaha.edu/compsicfacproc)

**Recommended Citation**

Lierler, Yuliya; Erdem, Esta; and Lifschitz, Vladimir, "Fages' Theorem and Answer Set Programming" (2000). *Computer Science Faculty Proceedings & Presentations*. 34.

[https://digitalcommons.unomaha.edu/compsicfacproc/34](https://digitalcommons.unomaha.edu/compsicfacproc/34)
Fages’ Theorem
and Answer Set Programming

Yuliya Babovich, Esra Erdem and Vladimir Lifschitz
Department of Computer Sciences
University of Texas at Austin
Austin, TX 78712, USA
Email: {yuliya,esra,vl}@cs.utexas.edu

Abstract
We generalize a theorem by François Fages that describes the relationship between the completion semantics and the answer set semantics for logic programs with negation as failure. The study of this relationship is important in connection with the emergence of answer set programming. Whenever the two semantics are equivalent, answer sets can be computed by a satisfiability solver, and the use of answer set solvers such as SMOCKS and DLV is unnecessary. A logic programming representation of the blocks world due to Ilkka Niemelä is discussed as an example.

Introduction
This note is about the relationship between the completion semantics (Clark 1978) and the answer set (“stable model”) semantics (Gelfond & Lifschitz 1991) for logic programs with negation as failure. The study of this relationship is important in connection with the emergence of answer set programming (Marek & Truszczynski 1999; Niemelä 1999; Lifschitz 1999). Whenever the two semantics are equivalent, answer sets can be computed by a satisfiability solver, and the use of “answer set solvers” such as SMOCKS1 and DLV2 is unnecessary.

Consider a finite propositional (or grounded) program $\Pi$ without classical negation, and a set $X$ of atoms. If $X$ is an answer set for $\Pi$ then $X$, viewed as a truth assignment, satisfies the completion of $\Pi$. The converse, generally, is not true. For instance, the completion of

$$p \leftarrow p$$

is $p \equiv p$. This formula has two models $\emptyset$, $\{p\}$; the first is an answer set for (1), but the second is not. François Fages [1994] defined a syntactic condition on logic programs that implies the equivalence between the two semantics—“positive-order-consistency,” also called “tightness” (Lifschitz 1996). What he requires is the existence of a function $\lambda$ from atoms to nonnegative integers (or, more generally, ordinals) such that,

$$\lambda(A_1), \ldots, \lambda(A_m) < \lambda(A_0).$$

It is clear, for instance, that program (1) is not tight. Fages proved that, for a tight program, every model of its completion is an answer set. Thus, for tight programs, the completion semantics and the answer set semantics are equivalent.

Our generalization of Fages’ theorem allows us to draw similar conclusions for some programs that are not tight. Here is one such program:

$$A_0 \leftarrow A_1, \ldots, A_m, \text{not } A_{m+1}, \ldots, \text{not } A_n$$

in $\Pi$,

$$\lambda(A_1), \ldots, \lambda(A_m) < \lambda(A_0).$$

It is not tight. Nevertheless, each of the two models $\{p\}$, $\{q\}$ of its completion

$$p \equiv \neg q \lor (p \land r),$$
$$q \equiv \neg p,$$
$$r \equiv \bot$$

is an answer set for (2).

The idea of this generalization is to make function $\lambda$ partial. Instead of tight programs, we will consider programs that are “tight on a set of literals.”

First we relate answer sets to a model-theoretic counterpart of completion introduced in (Apt, Blair, & Walker 1988), called supportedness. This allows us to make the theorem applicable to programs with both negation as failure and classical negation, and to programs with infinitely many rules.3 Then a corollary about completion is derived, and applied to a logic programming representation of the blocks world due to Ilkka Niemelä. We show how the satisfiability solver SATO (Zhang 1997) can be used to find answer sets for that representation, and compare the performance of SMOCKS and SATO on several benchmarks.

1The familiar definition of completion (see Appendix) is applicable to finite programs only, unless we allow infinite disjunctions in completion formulas.
Generalized Fages’ Theorem
We define a rule to be an expression of the form
\[ \text{Hea}d \leftarrow L_1, \ldots, L_m, \neg L_{m+1}, \ldots, \neg L_n \quad (3) \]
\((n \geq m \geq 0)\) where each \(L_i\) is a literal (propositional atom possibly preceded by classical negation \(\neg\)), and \(\text{Hea}d\) is a literal or the symbol \(\bot\). A rule (3) is called a fact if \(n = 0\), and a constraint if \(\text{Hea}d = \bot\). A program is a set of rules. The familiar definitions of answer sets, closed sets and supported sets for a program, as well as the definition of the completion of a program, are reproduced in the appendix.

Instead of “level mappings” used by Fages, we consider here partial level mappings—partial functions from literals to ordinals. A program \(\Pi\) is tight on a set \(X\) of literals if there exists a partial level mapping \(\lambda\) with the domain \(X\) such that, for every rule (3) in \(\Pi\), if \(\text{Hea}d, L_1, \ldots, L_m \in X\) then
\[ \lambda(L_1), \ldots, \lambda(L_m) < \lambda(\text{Hea}d). \]
(For the constraints in \(\Pi\) this condition holds trivially, because the head of a constraint is not a literal and thus cannot belong to \(X\).)

**Theorem.** For any program \(\Pi\) and any consistent set \(X\) of literals such that \(\Pi\) is tight on \(X\), \(X\) is an answer set for \(\Pi\) iff \(X\) is closed under and supported by \(\Pi\).

The proof below is almost unchanged from the proof of Fages’ theorem given in (Lifschitz & Turner 1999, Section 7.4).

**Lemma.** For any program \(\Pi\) without negation as failure and any consistent set \(X\) of literals such that \(\Pi\) is tight on \(X\), if \(X\) is closed under and supported by \(\Pi\), then \(X\) is an answer set for \(\Pi\).

**Proof:** We need to show that \(X\) is minimal among the sets closed under \(\Pi\). Assume that it is not. Let \(Y\) be a proper subset of \(X\) that is closed under \(\Pi\), and let \(\lambda\) be a partial level mapping establishing that \(\Pi\) is tight on \(X\). Take a literal \(L \in X \setminus Y\) such that \(\lambda(L)\) is minimal. Since \(X\) is supported by \(\Pi\), there is a rule
\[ L \leftarrow L_1, \ldots, L_m \]
in \(\Pi\) such that \(L_1, \ldots, L_m \in X\). By the choice of \(\lambda\),
\[ \lambda(L_1), \ldots, \lambda(L_m) < \lambda(L). \]
By the choice of \(L\), we can conclude that
\[ L_1, \ldots, L_m \in Y. \]
Consequently \(Y\) is not closed under \(\Pi\), contrary to the choice of \(Y\).

**Proof of the Theorem:** Left-to-right, the proof is straightforward. Right-to-left: assume that \(X\) is closed under and supported by \(\Pi\). Then \(X\) is closed under and supported by \(\Pi^X\). Since \(\Pi\) is tight on \(X\), so is \(\Pi^X\). Hence, by the lemma, \(X\) is an answer set for \(\Pi^X\), and consequently an answer set for \(\Pi\).

In the special case when \(\Pi\) is a finite program without classical negation, a set of atoms satisfies the completion of \(\Pi\) iff it is closed under and supported by \(\Pi\). We conclude:

**Corollary 1.** For any finite program \(\Pi\) without classical negation and any set \(X\) of atoms such that \(\Pi\) is tight on \(X\), \(X\) is an answer set for \(\Pi\) iff \(X\) satisfies the completion of \(\Pi\).

For instance, program (2) is tight on the model \(\{p\}\) of its completion: take \(\lambda(p) = 0\). By Corollary 1, it follows that \(\{p\}\) is an answer set for (2). In a similar way, the theorem shows that \(\{q\}\) is an answer set also.

By \(\text{pos}(\Pi)\) we denote the set of all literals that occur without negation as failure at least once in the body of a rule of \(\Pi\).

**Corollary 2.** For any program \(\Pi\) and any consistent set \(X\) of literals disjoint from \(\text{pos}(\Pi)\), \(X\) is an answer set for \(\Pi\) iff \(X\) is closed under and supported by \(\Pi\).

**Corollary 3.** For any finite program \(\Pi\) without classical negation and any set \(X\) of atoms disjoint from \(\text{pos}(\Pi)\), \(X\) is an answer set for \(\Pi\) iff \(X\) satisfies the completion of \(\Pi\).

To derive Corollary 2 from the theorem, and Corollary 3 from Corollary 1, take \(\lambda(L) = 0\) for every \(L \in X\).

Consider, for instance, the program
\[ p \leftarrow \neg q, \quad q \leftarrow \neg p, \quad r \leftarrow r, \quad p \leftarrow r. \quad (4) \]

The completion of (4) is
\[ p \equiv \neg q \lor r, \quad q \equiv \neg p, \quad r \equiv r. \]

The models of these formulas are \(\{p\}\), \(\{q\}\) and \(\{p,r\}\). The only literal occurring in the bodies of the rules of (4) without negation as failure is \(r\). In accordance with Corollary 3, the models of the completion that do not contain \(r\)—sets \(\{p\}\) and \(\{q\}\)—are answer sets for (4).

Planning in the Blocks World
As a more interesting example, consider a logic programming encoding of the blocks world due to Ilkka Niemelä. The main part of the encoding consists of the following schematic rules:

\[ \text{goal} :- \text{time}(T), \text{goal}(T). \]
\[ :- \neg \text{goal}. \]

\[ \text{goal}(T2) :- \text{nextstate}(T2,T1), \text{goal}(T1). \]

\[ \text{move}(X,Y,T):- \]
\[ \text{time}(T), \text{block}(X), \text{object}(Y), X \neq Y, \text{on}_{\text{something}}(X,T), \text{available}(Y,T), \text{not\ covered}(X,T), \text{not\ covered}(Y,T), \text{not\ blocked\ move}(X,Y,T). \]
on(X,Y,T2) :-
    block(X), object(Y), nextstate(T2,T1),
    moveop(X,Y,T1).

on_something(X,T) :-
    block(X), object(Z), time(T), on(X,Z,T).

available(table,T) :- time(T).

available(X,T) :-
    block(X), time(T), on_something(X,T).

covered(X,T) :-
    block(Z), block(X), time(T), on(Z,X,T).

on(X,Y,T2) :-
    nextstate(T2,T1), block(X), object(Y),
    on(X,Y,T1), not moving(X,T1).

moving(X,T) :- time(T), block(X), object(Y),
    moveop(X,Y,T).

blocked_move(X,Y,T) :-
    block(X), object(Y), time(T), goal(T).

blocked_move(X,Y,T) :-
    time(T), block(X), object(Y),
    not moveop(X,Y,T).

blocked_move(X,Y,T) :-
    block(X), object(Y), object(Z), time(T),
    moveop(X,Z,T), Y != Z.

blocked_move(X,Y,T) :-
    block(X), object(Y), time(T), moving(Y,T).

blocked_move(X,Y,T) :-
    block(X), block(Y), block(Z), time(T),
    moveop(Z,Y,T), X != Z.

:- block(X), time(T), moveop(X,table,T),
    on(X,table,T).

:- nextstate(T2,T1), block(X), object(Y),
    moveop(X,Y,T1), moveop(X,table,T2).

nextstate(Y,X) :- time(X), time(Y),
    Y = X + 1.

object(table).
object(X) :- block(X).

block(a). block(b). block(c).

(iii) a set of facts encoding the initial state, such as
on(a,b,0). on(b,table,0).

(iv) a rule that encodes the goal, such as

goal(T) :- time(T), on(a,b,T), on(b,c,T).

The union is given as input to the “intelligent grounding” program LPARSE, and the result of grounding is passed on to SMODELS (Niemelä 1999, Section 7). The answer sets for the program correspond to valid plans.

Concurrently executed actions are allowed in this formalization as long as their effects are not in conflict, so that they can be arbitrarily interleaved.

The schematic rules above contain the variables T, T1, T2, X, Y, Z that range over the object constants occurring in the program, that is, over the nonnegative integers that occur in the definition of time/1, the names of blocks a, b, . . . that occur in the definition of block/1, and the object constant table.

The expressions in the bodies of the schematic rules that contain = and != restrict the constants that are substituted for the variables in the process of grounding. For instance, we understand the schematic rule

nextstate(Y,X) :- time(X), time(Y),
    Y = X + 1.

as an abbreviation for the set of all ground instances of

nextstate(Y,X) :- time(X), time(Y).

in which X and Y are instantiated by a pair of consecutive integers. The schematic rule

blocked_move(X,Y,T) :-
    block(X), object(Y), object(Z), time(T),
    moveop(X,Z,T), Y != Z.

stands for the set of all ground instances of

blocked_move(X,Y,T) :-
    block(X), object(Y), object(Z), time(T),
    moveop(X,Z,T), Y != Z.

in which Y and Z are instantiated by different object constants.

According to this understanding of variables and “built-in predicates,” Niemelä’s schematic program, including rules (i)-(iv), is an abbreviation for a finite program BW in the sense defined above.

In the proposition below we assume that schematic rule (iv) has the form

goal(T) :- time(T), ...

where the dots stand for a list of schematic atoms with the predicate symbol on and the last argument T.

**Proposition.** Program BW is tight on each of the models of its completion.
Lemma. For any atom of the form nextstate(Y,X) that belongs to a model of the completion of program BW, Y = X + 1.

Proof: The completion of BW contains the formula
\[ \text{nextstate}(Y, X) \equiv \text{false} \]
for all Y, X such that Y ≠ X + 1.

Proof of the Proposition. Let X be an answer set for BW. By \( T_{\text{max}} \) we denote the largest argument of time/1 in its definition (i). Consider the partial level mapping \( \lambda \) with domain \( X \) defined as follows:

\[
\begin{align*}
\lambda(\text{time}(T)) &= 0, \\
\lambda(\text{block}(X)) &= 0, \\
\lambda(\text{object}(X)) &= 1, \\
\lambda(\text{nextstate}(Y, X)) &= 1, \\
\lambda(\text{covered}(X, T)) &= 4 \cdot T + 3, \\
\lambda(\text{on something}(X, T)) &= 4 \cdot T + 3, \\
\lambda(\text{available}(X, T)) &= 4 \cdot T + 4, \\
\lambda(\text{move op}(X, Y, T)) &= 4 \cdot T + 4, \\
\lambda(\text{on}(X, Y, T)) &= 4 \cdot T + 5, \\
\lambda(\text{goal}(T)) &= 4 \cdot T, \\
\lambda(\text{moving}(X, T)) &= 4 \cdot T + 6, \\
\lambda(\text{goal}(T)) &= 4 \cdot T + 7, \\
\lambda(\text{blocked move}(X, Y, T)) &= 4 \cdot T + 7, \\
\lambda(\text{goal}) &= 4 \cdot T_{\text{max}} + 4.
\end{align*}
\]

This level mapping satisfies the inequality from the definition of a tight program for every rule of BW: the lemma above allows us to verify this assertion for the rules containing nextstate in the body.

According to Corollary 1, we can conclude that the answer sets for program BW can be equivalently characterized as the models of the completion of BW.

Answer Set Programming with CCALC and SATO

The equivalence of the completion semantics to the answer set semantics for program BW shows that it is not necessary to use an answer set solver, such as SMODELS, to compute answer sets for BW: a satisfiability solver can be used instead. Planning by giving the completion of BW as input to a satisfiability solver is a form of answer set programming and, at the same time, a special case of satisfiability planning (Kautz & Selman 1992).

The Causal Calculator, or CCALC\(^4\), is a system that is capable, among other things, of grounding and completing a schematic logic program, classifying the completion, and calling a satisfiability solver (for instance, SATO) to find a model. We have conducted a series of experiments aimed at comparing the run times of SATO, when its input is generated from BW by CCALC, with the run times of SMODELS, when its input is generated from BW by LPARSE.

Because the built-in arithmetic of CCALC is somewhat different from that of LPARSE, we had to modify BW slightly. Our CCALC input file uses variables of sorts object, block and time instead of the unary predicates with these names. The rules of BW that contain these predicates in their bodies are modified accordingly. For instance, rule
\[
\text{on something}(X, T) :- \\
\text{block}(X), \text{object}(Z), \text{time}(T), \text{on}(X, Z, T).
\]

This turns into
\[
\text{on something}(B_1, T) :- \text{on}(B_1, 0_2, T).
\]

The macro expansion facility of CCALC expands
\[
\text{nextstate}(T_2, T_1)
\]
into the expression
\[
T_2 \text{ is } T_1 + 1
\]
that contains Prolog's built-in is.

Figure 1 shows the run times of SMODELS (Version 2.25) and SATO (Version 3.1.2) in seconds, measured using the Unix time command, on the benchmarks from (Niemelä 1999, Section 9, Table 3). For each problem, one of the two entries corresponds to the largest number of steps for which the problem is not solvable, and the other to the smallest number of steps for which a solution exists. The experiments were performed on an UltraSPARC with 124 MB main memory and a 167 MHz CPU.

The numbers in Figure 1 are “search times”—the grounding and completion times are not included. The computation involved in grounding and completion does not depend on the initial state or the goal of the planning problem and, in this sense, can be viewed as “preprocessing.” LPARSE performs grounding more efficiently than CCALC, partly because the former is written in C++ and the latter in Prolog. The last benchmark in Figure 1 was grounded by LPARSE (Version 0.99.49) in 16 seconds; CCALC (Version 1.23) spent 30 seconds in grounding and about the same amount of time forming the completion. But the size of the grounded program is approximately the same in both cases: LPARSE generated 191621 rules containing 13422 atoms, and CCALC generated 200661 rules containing 13410 atoms.

<table>
<thead>
<tr>
<th>Problem</th>
<th>Blocks</th>
<th>Steps</th>
<th>Run time of SMODELS</th>
<th>Run time of SATO</th>
</tr>
</thead>
<tbody>
<tr>
<td>large.c</td>
<td>15</td>
<td>7</td>
<td>9.86</td>
<td>1.82</td>
</tr>
<tr>
<td></td>
<td></td>
<td></td>
<td>8</td>
<td>31.25</td>
</tr>
<tr>
<td>large.d</td>
<td>17</td>
<td>8</td>
<td>18.25</td>
<td>2.96</td>
</tr>
<tr>
<td></td>
<td></td>
<td></td>
<td>9</td>
<td>62.48</td>
</tr>
<tr>
<td>large.e</td>
<td>19</td>
<td>9</td>
<td>27.31</td>
<td>5.40</td>
</tr>
<tr>
<td></td>
<td></td>
<td></td>
<td>10</td>
<td>101.4</td>
</tr>
</tbody>
</table>

Figure 1: Planning with BW: SATO vs. SMODELS

---

\(^4\)http://www.cs.utexas.edu/users/tag/ccalc/.
Discussion

Fages’ theorem, and its generalization proved in this note, allow us to compute answer sets for some programs by completing them and then calling a satisfiability solver. We showed that this method can be applied, for instance, to the representation of the blocks world proposed in (Niemelä 1999). This example shows that satisfiability solvers may serve as useful computational tools in answer set programming.

There are cases, on the other hand, when the completion method is not applicable. Consider computing Hamiltonian cycles in a directed graph (Marek & Truszczyński 1999). We combine the rules

\[ \text{in}(U,V) :- \text{edge}(U,V), \text{not out}(U,V). \]
\[ \text{out}(U,V) :- \text{edge}(U,V), \text{not in}(U,V). \]
\[ :- \text{in}(U,V), \text{in}(U,W), V \neq W. \]
\[ :- \text{in}(U,W), \text{in}(V,W), U \neq V. \]
\[ \text{reachable}(V) :- \text{in}(v_0,V). \]
\[ \text{reachable}(V) :- \text{reachable}(U), \text{in}(U,V). \]

with a set of facts defining the vertices and edges of the graph; \( v_0 \) is assumed to be one of the vertices. The answer sets for the resulting program correspond to the Hamiltonian cycles. Generally, the completion of the program has models different from its answer sets. Take, for instance, the graph consisting of two disjoint loops:

\[ \text{vertex}(v_0). \text{vertex}(v_1). \]
\[ \text{edge}(v_0,v_0). \text{edge}(v_1,v_1). \]

This graph has no Hamiltonian cycles, and, accordingly, the corresponding program has no answer sets. But the set

\[ \text{vertex}(v_0), \text{vertex}(v_1), \text{edge}(v_0,v_0), \]
\[ \text{edge}(v_1,v_1), \text{in}(v_0,v_0), \text{in}(v_1,v_1), \]
\[ \text{reachable}(v_0), \text{reachable}(v_1) \]

is a model of the program’s completion.

Acknowledgements

We are grateful to Victor Marek, Emilio Remolina, Mirek Truszczynski and Hudson Turner, and to the anonymous referees, for comments and criticisms. This work was partially supported by National Science Foundation under grant IIS-9732744.

References


Appendix: Definitions

The notion of an answer set is defined first for programs whose rules do not contain negation as failure. Let \( \Pi \) be such a program, and let \( X \) be a consistent set of literals. We say that \( X \) is closed under \( \Pi \) if, for every rule

\[ \text{Head} \leftarrow \text{Body} \]

in \( \Pi \), \( \text{Head} \in X \) whenever \( \text{Body} \subseteq X \). (For a constraint, this condition means that the body is not contained in \( X \).) We say that \( X \) is an answer set for \( \Pi \) if \( X \) is minimal among the sets closed under \( \Pi \) w.r.t. set inclusion. It is clear that a program without negation as failure can have at most one answer set.

To extend this definition to arbitrary programs, take any program \( \Pi \), and let \( X \) be a consistent set of literals. The reduce \( \Pi^X \) of \( \Pi \) relative to \( X \) is the set of rules

\[ \text{Head} \leftarrow L_1, \ldots, L_m \]

for all rules (3) in \( \Pi \) such that \( L_{m+1}, \ldots, L_n \not\in X \). Thus \( \Pi^X \) is a program without negation as failure. We say that \( X \) is an answer set for \( \Pi \) if \( X \) is an answer set for \( \Pi^X \).

A set \( X \) of literals is closed under \( \Pi \) if, for every rule (3) in \( \Pi \), \( \text{Head} \in X \) whenever \( L_1, \ldots, L_m \in X \) and \( L_{m+1}, \ldots, L_n \not\in X \). We say that \( X \) is supported by \( \Pi \) if, for every \( L \in X \), there is a rule (3) in \( \Pi \) such that \( \text{Head} = L \), \( L_1, \ldots, L_m \in X \) and \( L_{m+1}, \ldots, L_n \not\in X \).
Let $\Pi$ be a finite program without classical negation. If $H$ is an atom or the symbol $\bot$, by $\text{Comp}(\Pi, H)$ we denote the formula

$$H \equiv \bigvee(A_1 \land \cdots \land A_m \land \neg A_{m+1} \land \cdots \land \neg A_n)$$

where the disjunction extends over all rules

$$H \leftarrow A_1, \ldots, A_m, \text{not } A_{m+1}, \ldots, \text{not } A_n$$

in $\Pi$ with the head $H$. The completion of $\Pi$ is set of formulas $\text{Comp}(\Pi, H)$ for all $H$. 