

Contribution Title^{*}

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Abstract. This paper proved a method to computing the forgetting in CTL which has been submitted to IJCAI, from the resolution proposed by Zhang at all by extending the resolution rules.

Keywords: Forgetting · CTL · Model checking.

1 Introduction

As a logical notion, *forgetting* was first formally defined in propositional and first order logics by Lin and Reiter [13]. Over the last twenty years, researchers have developed forgetting notions and theories not only in classical logic but also in other non-classical logic systems [?], such as forgetting in logic programs under answer set/stable model semantics [23,6,20,18,17], forgetting in description logic [19,15,26] and knowledge forgetting in modal logic [25,16,14,8]. In application, forgetting has been used in planning [12], conflict solving [10,24], creating restricted views of ontologies [26], strongest and weakest definitions [9], SNC (WSC) [11] and so on.

Though forgetting has been extensively investigated from various aspects of different logical systems. However, the existing forgetting method in propositional logic, answer set programming, description logic and modal logic are not directly applicable in CTL. Similar with that in [25], we research forgetting in CTL from the semantic forgetting point of view. And it is shown that our definition of forgetting satisfies those four postulates of forgetting.

2 Preliminaries

We start with some technical and notational preliminaries. Throughout this paper, we fix a finite set \mathcal{A} of propositional variables (or atoms), and use V, V' for subsets of \mathcal{A} . In the following several parts, we will introduce the structure we use for CTL, syntactic and semantic of CTL and the normal form $\text{SNF}_{\text{CTL}}^g$ (Separated Normal Form with Global Clauses for CTL) of CTL [22].

^{*} Supported by organization x.

2.1 Model structure in CTL

In general, a transition system⁴ is described as a *model structure* (or *Kripke structure*)(in this article, we treat transition system and model structure as the same thing), and a model structure is a triple $\mathcal{M} = (S, R, L)$ [7], where

- S is a set of states,
- $R \subseteq S \times S$ is a total binary relation over S , i.e., for each state $s \in S$ there is a state $s' \in S$ such that $(s, s') \in R$, and
- L is an interpretation function $S \rightarrow 2^{\mathcal{A}}$ mapping every state to the set of atoms true at that state.

In this article, the same as [4], all of our results apply only to finite Kripke structures. Besides, we restrict ourselves to model structure $\mathcal{M} = (S, R, L, s_0)$ (similar with that in [22]) such that

- there exists a state s_0 , called the *initial state*, such that for every state $s \in S$ there is a path π_{s_0} s.t. $s \in \pi_{s_0}$.

We call a model structure \mathcal{M} on a set V of atoms if $L : S \rightarrow 2^V$, i.e., the labeling function L map every state to V (not the \mathcal{A}). A *path* π_{s_i} start from s_i of \mathcal{M} is a infinite sequence of states $\pi_{s_i} = (s_i, s_{i+1}, s_{i+2}, \dots)$, where for each j ($i \leq j$), $(s_j, s_{j+1}) \in R$. By $s' \in \pi_{s_i}$ we mean that s' is a state in the path π_{s_i} .

For a given model structure (S, R, L, s_0) and $s \in S$, the *computation tree* $\text{Tr}_n^{\mathcal{M}}(s)$ of \mathcal{M} (or simply $\text{Tr}_n(s)$), that has depth n and is rooted at s , is recursively defined as [4], for $n \geq 0$,

- $\text{Tr}_0(s)$ consists of a single node s with label s .
- $\text{Tr}_{n+1}(s)$ has as its root a node m with label s , and if $(s, s') \in R$ then the node m has a subtree $\text{Tr}_n(s')$ ⁵.

By s_n we mean the node at the n th level in tree $\text{Tr}_m(s)$ ($m \geq n$).

A *K-structure* (or *K-interpretation*) is a model structure $\mathcal{M} = (S, R, L, s_0)$ associating with a state $s \in S$, which is written as (\mathcal{M}, s) for convenience in the following. In the case s is an initial state of \mathcal{M} , the K-structure is *initial*.

2.2 Syntactic and semantic of CTL

In the following we briefly review the basic syntax and semantics of the *Computation Tree Logic* (CTL in short) [5]. The *signature* of \mathcal{L} includes:

- a finite set of Boolean variables, called *atoms* of \mathcal{L} : \mathcal{A} ;
- the classical connectives: \perp, \vee and \neg ;

⁴ According to [1], a *transition system* TS is a tuple $(S, \text{Act}, \rightarrow, I, \text{AP}, L)$ where (1) S is a set of states, (2) Act is a set of actions, (3) $\rightarrow \subseteq S \times \text{Act} \times S$ is a transition relation, (4) $I \subseteq S$ is a set of initial states, (5) AP is a set of atomic propositions, and (6) $L : S \rightarrow 2^{\text{AP}}$ is a labeling function.

⁵ Though some nodes of the tree may have the same label, they are different nodes in the tree.

- the path quantifiers: A and E;
- the temporal operators: X, F, G U and W, that means ‘neXt state’, ‘some Future state’, ‘all future states (Globally)’, ‘Until’ and ‘Unless’, respectively;
- parentheses: (and).

The (*existential normal form or ENF in short*) formulas of \mathcal{L} are inductively defined via a Backus Naur form:

$$\phi ::= \perp \mid p \mid \neg\phi \mid \phi \vee \phi \mid \text{EX}\phi \mid \text{EG}\phi \mid \text{E}[\phi \text{ U } \phi] \quad (1)$$

where $p \in \mathcal{A}$. The formulas $\phi \wedge \psi$ and $\phi \rightarrow \psi$ are defined in a standard manner of propositional logic. The other form formulas of \mathcal{L} are abbreviated using the forms of (1). Notice that, according to the above definition for formulas of CTL, each of the CTL *temporal connectives* has the form XY where $X \in \{A, E\}$ and $Y \in \{X, F, G, U, W\}$. The priorities for the CTL connectives are assumed to be (from the highest to the lowest):

$$\neg, \text{EX}, \text{EF}, \text{EG}, \text{AX}, \text{AF}, \text{AG} \prec \wedge \prec \vee \prec \text{EU}, \text{AU}, \text{EW}, \text{AW}, \rightarrow .$$

We are now in the position to define the semantics of \mathcal{L} . Let $\mathcal{M} = (S, R, L, s_0)$ be an model structure, $s \in S$ and ϕ a formula of \mathcal{L} . The *satisfiability* relationship between \mathcal{M} , s and ϕ , written $(\mathcal{M}, s) \models \phi$, is inductively defined on the structure of ϕ as follows:

- $(\mathcal{M}, s) \not\models \perp$;
- $(\mathcal{M}, s) \models p$ iff $p \in L(s)$;
- $(\mathcal{M}, s) \models \phi_1 \vee \phi_2$ iff $(\mathcal{M}, s) \models \phi_1$ or $(\mathcal{M}, s) \models \phi_2$;
- $(\mathcal{M}, s) \models \neg\phi$ iff $(\mathcal{M}, s) \not\models \phi$;
- $(\mathcal{M}, s) \models \text{EX}\phi$ iff $(\mathcal{M}, s_1) \models \phi$ for some $s_1 \in S$ and $(s, s_1) \in R$;
- $(\mathcal{M}, s) \models \text{EG}\phi$ iff \mathcal{M} has a path $(s_1 = s, s_2, \dots)$ such that $(\mathcal{M}, s_i) \models \phi$ for each $i \geq 1$;
- $(\mathcal{M}, s) \models \text{E}[\phi_1 \text{ U } \phi_2]$ iff \mathcal{M} has a path $(s_1 = s, s_2, \dots)$ such that, for some $i \geq 1$, $(\mathcal{M}, s_i) \models \phi_2$ and $(\mathcal{M}, s_j) \models \phi_1$ for each $j < i$.

Similar to the work in [4,2], only initial K-structures are considered to be candidate models in the following, unless explicitly stated. Formally, an initial K-structure \mathcal{K} is a *model* of a formula ϕ whenever $\mathcal{K} \models \phi$. Let Π be a set of formulae, $\mathcal{K} \models \Pi$ if for each $\phi \in \Pi$ there is $\mathcal{K} \models \phi$. We denote $\text{Mod}(\phi)$ ($\text{Mod}(\Pi)$) the set of models of ϕ (Π). The formula ϕ (set Π of formulae) is *satisfiable* if $\text{Mod}(\phi) \neq \emptyset$ ($\text{Mod}(\Pi) \neq \emptyset$). Since both the underlying states in model structure and signatures are finite, $\text{Mod}(\phi)$ ($\text{Mod}(\Pi)$) is finite for any formula ϕ (set Π of formulae).

Let ϕ_1 and ϕ_2 be two formulas or set of formulas. By $\phi_1 \models \phi_2$ we denote $\text{Mod}(\phi_1) \subseteq \text{Mod}(\phi_2)$. By $\phi_1 \equiv \phi_2$ we mean $\phi_1 \models \phi_2$ and $\phi_2 \models \phi_1$. In this case ϕ_1 is *equivalent* to ϕ_2 .

Let ϕ be a formula or set of formulas. By $\text{Var}(\phi)$ we mean the set of atoms occurring in ϕ . Let $V \subseteq \mathcal{A}$. The formula ϕ is *V-irrelevant*, written $\text{IR}(\phi, V)$, if there is a formula ψ with $\text{Var}(\psi) \cap V = \emptyset$ such that $\phi \equiv \psi$.

2.3 The normal form of CTL

It has proved that any CTL formula φ can be transformed into a set T_φ of $\text{SNF}_{\text{CTL}}^g$ (Separated Normal Form with Global Clauses for CTL) clauses in polynomial time such that φ is satisfiable iff T_φ is satisfiable [21]. An important difference between CTL formulae and $\text{SNF}_{\text{CTL}}^g$ is that $\text{SNF}_{\text{CTL}}^g$ is an extension of the syntax of CTL to use indices. These indices can be used to preserve a particular path context. The language of $\text{SNF}_{\text{CTL}}^g$ clauses is defined over an extension of CTL. That is the language is based on: (1) the language of CTL; (2) a propositional constant **start**; (3) a countably infinite index set Ind ; and (4) temporal operators: $E_{\langle \text{ind} \rangle} X$, $E_{\langle \text{ind} \rangle} F$, $E_{\langle \text{ind} \rangle} G$, $E_{\langle \text{ind} \rangle} U$ and $E_{\langle \text{ind} \rangle} W$.

The priorities for the $\text{SNF}_{\text{CTL}}^g$ connectives are assumed to be (from the highest to the lowest):

$$\neg, (EX, E_{\langle \text{ind} \rangle} X), (EF, E_{\langle \text{ind} \rangle} F), (EG, E_{\langle \text{ind} \rangle} G), AX, AF, AG \\ \prec \wedge \prec \vee \prec (EU, E_{\langle \text{ind} \rangle} U), AU, (EW, E_{\langle \text{ind} \rangle} W), AW, \rightarrow .$$

Where the operators in the same brackets have the same priority.

Before talked about the sematic of this language, we introduce the $\text{SNF}_{\text{CTL}}^g$ clauses at first. The $\text{SNF}_{\text{CTL}}^g$ clauses consists of formulae of the following forms.

$$\begin{aligned} AG(\mathbf{start} \supset \bigvee_{j=1}^k m_j) & \quad (\text{initial clause}) \\ AG(\text{true} \supset \bigvee_{j=1}^k m_j) & \quad (\text{global clause}) \\ AG(\bigwedge_{i=1}^n l_i \supset AX \bigvee_{j=1}^k m_j) & \quad (\text{A - step clause}) \\ AG(\bigwedge_{i=1}^n l_i \supset E_{\langle \text{ind} \rangle} X \bigvee_{j=1}^k m_j) & \quad (\text{E - step clause}) \\ AG(\bigwedge_{i=1}^n l_i \supset AF l) & \quad (\text{A - sometime clause}) \\ AG(\bigwedge_{i=1}^n l_i \supset E_{\langle \text{ind} \rangle} F l) & \quad (\text{E - sometime clause}). \end{aligned}$$

where $k \geq 0$, $n > 0$, **start** is a propositional constant, l_i ($1 \leq i \leq n$), m_j ($1 \leq j \leq k$) and l are literals, that is atomic propositions or their negation and ind is an element of Ind (Ind is a countably infinite index set). By clause we mean the classical clause or the $\text{SNF}_{\text{CTL}}^g$ clause unless explicitly stated.

Formulae of $\text{SNF}_{\text{CTL}}^g$ over \mathcal{A} are interpreted in Ind -model structure $\mathcal{M} = (S, R, L, [-], s_0)$, where S , R , L and s_0 is the same as our model structure talked in 2.1 and $[-] : \text{Ind} \rightarrow 2^{(S \times S)}$ maps every index $\text{ind} \in \text{Ind}$ to a successor function $[\text{ind}]$ which is a functional relation on S and a subset of the binary accessibility relation R , such that for every

$s \in S$ there exists exactly a state $s' \in S$ such that $(s, s') \in [ind]$ and $(s, s') \in R$. An infinite path $\pi_{s_i}^{(ind)}$ is an infinite sequence of states $s_i, s_{i+1}, s_{i+2}, \dots$ such that for every $j \geq i$, $(s_j, s_{j+1}) \in [ind]$.

Similarly, an *Ind-structure* (or *Ind-interpretation*) is a Ind-model structure $\mathcal{M} = (S, R, L, [-], s_0)$ associating with a state $s \in S$, which is written as (\mathcal{M}, s) for convenience in the following. In the case s is an initial state of \mathcal{M} , the Ind-structure is *initial*.

The semantics of $\text{SNF}_{\text{CTL}}^g$ is an extension of the semantics of CTL defined in Section 2.2 except using the Ind-model structure $\mathcal{M} = (S, R, L, [-], s_0)$ replace model structure, $(\mathcal{M}, s_i) \models \mathbf{start}$ iff $s_i = s_0$ and for all $E_{(ind)}\Gamma$ are explained in the path $\pi_{s_i}^{(ind)}$, where $\Gamma \in \{X, G, U, W\}$. The semantics of $\text{SNF}_{\text{CTL}}^g$ is then defined as shown next as an extension of the semantics of CTL defined in Section 2.2. Let φ and ψ be two $\text{SNF}_{\text{CTL}}^g$ formulae and $\mathcal{M} = (S, R, L, [-], s_0)$ be an Ind-model structure, the relation “ \models ” between $\text{SNF}_{\text{CTL}}^g$ formulae and \mathcal{M} is defined recursively as follows:

- $(\mathcal{M}, s_i) \models \mathbf{start}$ iff $s_i = s_0$;
- $(\mathcal{M}, s_i) \models E_{(ind)}X\psi$ iff for the path $\pi_{s_i}^{(ind)}$, $(\mathcal{M}, s_{i+1}) \models \psi$;
- $(\mathcal{M}, s_i) \models E_{(ind)}G\psi$ iff for every $s_j \in \pi_{s_i}^{(ind)}$, $(\mathcal{M}, s_j) \models \psi$;
- $(\mathcal{M}, s_i) \models E_{(ind)}[\varphi U \psi]$ iff there exists $s_j \in \pi_{s_i}^{(ind)}$ such that $(\mathcal{M}, s_j) \models \psi$ and for every $s_k \in \pi_{s_i}^{(ind)}$, if $i \leq k < j$, then $(\mathcal{M}, s_k) \models \varphi$;
- $(\mathcal{M}, s_i) \models E_{(ind)}F\psi$ iff $(\mathcal{M}, s_i) \models E_{(ind)}[\top U \psi]$;
- $(\mathcal{M}, s_i) \models E_{(ind)}[\varphi W \psi]$ iff $(\mathcal{M}, s_i) \models E_{(ind)}G\varphi$ or $(\mathcal{M}, s_i) \models E_{(ind)}[\varphi U \psi]$.

The semantics of the remaining operators is analogous to that given previously but in the extended Ind-model structure $\mathcal{M} = (S, R, L, [-], s_0)$. A $\text{SNF}_{\text{CTL}}^g$ formula φ is satisfiable, iff for some Ind-model structure $\mathcal{M} = (S, R, L, [-], s_0)$, $(\mathcal{M}, s_0) \models \varphi$, and unsatisfiable otherwise. And if $(\mathcal{M}, s_0) \models \varphi$ then (\mathcal{M}, s_0) is called a Ind-model of φ , and we say that (\mathcal{M}, s_0) satisfies φ . By $T \wedge \varphi$ we mean $\bigwedge_{\psi \in T} \psi \wedge \varphi$, where T is a set of formulae. Other terminologies are similar with those in section 2.2.

3 Problem Definition

In order to define our problem, *i.e.* forgetting in CTL, we review our definition of V -bisimulation (read ?? for more details).

Definition 1. Let $V \subseteq \mathcal{A}$ and $\mathcal{K}_i = (\mathcal{M}_i, s_i)$ ($i = 1, 2$) be \mathcal{K} -structures (Ind-structures). Then $(\mathcal{K}_1, \mathcal{K}_2) \in \mathcal{B}$ if and only if

- (i) $L_1(s_1) - V = L_2(s_2) - V$,
- (ii) for every $(s_1, s'_1) \in R_1$, there is $(s_2, s'_2) \in R_2$ such that $(\mathcal{K}'_1, \mathcal{K}'_2) \in \mathcal{B}$, and
- (iii) for every $(s_2, s'_2) \in R_2$, there is $(s_1, s'_1) \in R_1$

where $\mathcal{K}'_i = (\mathcal{M}_i, s'_i)$ with $i \in \{1, 2\}$.

Proposition 1. Let $i \in \{1, 2\}$, $V_1, V_2 \subseteq \mathcal{A}$, s'_i s be two states and π'_i s be two paths, and $\mathcal{K}_i = (\mathcal{M}_i, s_i)$ ($i = 1, 2, 3$) be \mathcal{K} -structures (Ind-structures) such that $\mathcal{K}_1 \leftrightarrow_{V_1} \mathcal{K}_2$ and $\mathcal{K}_2 \leftrightarrow_{V_2} \mathcal{K}_3$. Then:

- (i) $s'_1 \leftrightarrow_{V_i} s'_2$ ($i = 1, 2$) implies $s'_1 \leftrightarrow_{V_1 \cup V_2} s'_2$;
- (ii) $\pi'_1 \leftrightarrow_{V_i} \pi'_2$ ($i = 1, 2$) implies $\pi'_1 \leftrightarrow_{V_1 \cup V_2} \pi'_2$;
- (iii) for each path π_{s_1} of \mathcal{M}_1 there is a path π_{s_2} of \mathcal{M}_2 such that $\pi_{s_1} \leftrightarrow_{V_1} \pi_{s_2}$, and vice versa;
- (iv) $\mathcal{K}_1 \leftrightarrow_{V_1 \cup V_2} \mathcal{K}_3$;
- (v) If $V_1 \subseteq V_2$ then $\mathcal{K}_1 \leftrightarrow_{V_2} \mathcal{K}_2$.

Now we give the formal definition of forgetting in CTL from the semantic forgetting point view.

Definition 2 (Forgetting). Let $V \subseteq \mathcal{A}$ and ϕ a CTL formula. A CTL formula ψ with $\text{Var}(\psi) \cap V = \emptyset$ is a result of forgetting V from ϕ , if

$$\text{Mod}(\psi) = \{\mathcal{K} \text{ is initial} \mid \exists \mathcal{K}' \in \text{Mod}(\phi) \ \& \ \mathcal{K}' \leftrightarrow_V \mathcal{K}\}. \quad (2)$$

Where \mathcal{K} and \mathcal{K}' are K-structures.

Note that if both ψ and ψ' are results of forgetting V from ϕ then $\text{Mod}(\psi) = \text{Mod}(\psi')$, i.e., ψ and ψ' have the same models. In the sense of equivalence the forgetting result is unique (up to equivalence).

Similar with the V -bisimulation between K-structures, we define the $\langle V, I \rangle$ -bisimulation between Ind-structures as follows:

Definition 3. ($\langle V, I \rangle$ -bisimulation) Let $\mathcal{M}_i = (S_i, R_i, L_i, [-]_i, s_0^i)$ with $i \in \{1, 2\}$ be two Ind-structures, V be a set of atoms and $I \subseteq \text{Ind}$. The $\langle V, I \rangle$ -bisimulation $\beta_{\langle V, I \rangle}$ between initial Ind-structures is a set that satisfy $((\mathcal{M}_1, s_0^1), (\mathcal{M}_2, s_0^2)) \in \beta_{\langle V, I \rangle}$ if and only if $(\mathcal{M}_1, s_0^1) \leftrightarrow_V (\mathcal{M}_2, s_0^2)$ and $\forall j \notin I$ there is

- (i) $\forall (s, s_1) \in [j]_1$ there is $(s', s'_1) \in [j]_2$ such that $s \leftrightarrow_V s'$ and $s_1 \leftrightarrow_V s'_1$, and
- (ii) $\forall (s', s'_1) \in [j]_2$ there is $(s, s_1) \in [j]_1$ such that $s \leftrightarrow_V s'$ and $s_1 \leftrightarrow_V s'_1$.

Apparently, this definition is similar with our concept V -bisimulation except that this $\langle V, I \rangle$ -bisimulation has introduced the index.

Proposition 2. Let $i \in \{1, 2\}$, $V_1, V_2 \subseteq \mathcal{A}$, $I_1, I_2 \subseteq \text{Ind}$ and $\mathcal{K}_i = (\mathcal{M}_i, s_0^i)$ ($i = 1, 2, 3$) be Ind-structures such that $\mathcal{K}_1 \leftrightarrow_{\langle V_1, I_1 \rangle} \mathcal{K}_2$ and $\mathcal{K}_2 \leftrightarrow_{\langle V_2, I_2 \rangle} \mathcal{K}_3$. Then:

- (i) $\mathcal{K}_1 \leftrightarrow_{\langle V_1 \cup V_2, I_1 \cup I_2 \rangle} \mathcal{K}_3$;
- (ii) If $V_1 \subseteq V_2$ and $I_1 \subseteq I_2$ then $\mathcal{K}_1 \leftrightarrow_{\langle V_2, I_2 \rangle} \mathcal{K}_2$.

Proof. (i) By Proposition 1 we have $\mathcal{K}_1 \leftrightarrow_{V_1 \cup V_2} \mathcal{K}_3$. For (i) of Definition 3 we can prove it as follows: $\forall (s, s_1) \in [j]_1$ there is a $(s', s'_1) \in [j]_2$ such that $s \leftrightarrow_{V_1} s'$ and $s_1 \leftrightarrow_{V_1} s'_1$ and there is a $(s'', s''_1) \in [j]_3$ such that $s' \leftrightarrow_{V_2} s''$ and $s'_1 \leftrightarrow_{V_2} s''_1$, and then we have $\forall (s, s_1) \in [j]_1$ there is a $(s'', s''_1) \in [j]_3$ such that $s \leftrightarrow_{V_1 \cup V_2} s''$ and $s_1 \leftrightarrow_{V_1 \cup V_2} s''_1$. The (ii) of Definition 3 can be proved similarly.

(ii) This can be proved from (i).

4 The Calculus

Resolution in CTL is a method to decide the satisfiability of a CTL formula. In this paper, we will explore a resolution-based method to compute forgetting in CTL. In this part we use the transformation rules Trans(1) to Trans(12) and resolution rules (SRES1), ..., (SRES8), RW1, RW2, (ERES1), (ERES2) in [22].

The key problems of this method include (1) How to fill the gap between CTL and $\text{SNF}_{\text{CTL}}^g$; and (2) How to eliminate the irrelevant atoms in the formula. We will resolve these two problems by $\langle V, I \rangle$ -bisimulation and *substitution* operator. For convenient, we use $V \subseteq \mathcal{A}$ denote the set we want to forget, $V' \subseteq \mathcal{A}$ with $V \cap V' = \emptyset$ the set of atoms (I be the set of index) introduced in the transformation process, φ the CTL formula, T_φ be the set of $\text{SNF}_{\text{CTL}}^g$ clause obtained from φ by using transformation rules and $\mathcal{M} = (S, R, L, [-], s_0)$ unless explicitly stated. Let T, T' be two set of formulae, I a set of indexes and $V'' \subseteq \mathcal{A}$, by $T \equiv_{\langle V'', I \rangle} T'$ we mean that $\forall (\mathcal{M}, s_0) \in \text{Mod}(T)$ there is a (\mathcal{M}', s'_0) such that $(\mathcal{M}, s_0) \leftrightarrow_{\langle V'', I \rangle} (\mathcal{M}', s'_0)$ and $(\mathcal{M}', s'_0) \models T'$ and vice versa.

The following proposition is important for our algorithm since

Proposition 3. *Let P, P_i and φ_i be CTL formulas, then*

- (i) $P \supset \text{E}_{\langle \text{ind} \rangle} \text{X}\varphi_1 \wedge \dots \wedge P \supset \text{E}_{\langle \text{ind} \rangle} \text{X}\varphi_n \equiv_{\langle \emptyset, \{\text{ind}\} \rangle} P \supset \text{EX} \bigwedge_{i \in \{0, \dots, n\}} \varphi_i$
- (ii) $P_1 \supset \text{E}_{\langle \text{ind} \rangle} \text{X}\varphi_1 \wedge \dots \wedge P_n \supset \text{E}_{\langle \text{ind} \rangle} \text{X}\varphi_n \in T \equiv_{\langle \emptyset, \{\text{ind}\} \rangle} \bigwedge_{e \in 2^{\{0, \dots, n\}} \setminus \{\emptyset\}} (\bigwedge_{i \in e} P_i \supset \text{EX}(\bigwedge_{i \in e} \varphi_i))$,
- (iii) $P \supset \text{E}_{\langle \text{ind} \rangle} \text{F}\varphi_1 \wedge \dots \wedge P \supset \text{E}_{\langle \text{ind} \rangle} \text{F}\varphi_n \in T \equiv_{\langle \emptyset, \{\text{ind}\} \rangle} P \supset \bigvee \text{EF}(\varphi_{j_1} \wedge \text{EF}(\varphi_{j_2} \wedge \text{EF}(\dots \wedge \text{EF}\varphi_{j_n})))$, where (j_1, \dots, j_n) are sequences of all elements in $\{0, \dots, n\}$,
- (iv) $P \supset (C \vee \text{E}_{\langle \text{ind} \rangle} \text{X}\varphi_1) \wedge P \supset \text{E}_{\langle \text{ind} \rangle} \text{X}\varphi_2 \equiv_{\langle \emptyset, \{\text{ind}\} \rangle} P \supset ((C \wedge \text{EX}\varphi_2) \vee \text{EX}(\varphi_1 \wedge \varphi_2))$

Proof. It is easy to check.

The algorithm of computing the forgetting in CTL is as Algorithm 1 and the block diagram is as Figure 1. The main idea of this algorithm is to change the CTL formula into a set of $\text{SNF}_{\text{CTL}}^g$ clauses at first (the Tran process), and then compute all the possible resolutions on the specified set of atoms (the Res process), eliminate all the irrelevant atoms which dose not be removed by the resolution. We will describe this process in detail below (the Sub, EF, Elm and R processes). To describe our algorithm cleanly, we illustrate it with the following example.

Example 1. Let $\varphi = A((p \wedge q)U(f \vee m)) \wedge r$ and $V = \{p\}$.

In the following context we will show how to compute the $\text{F}_{\text{CTL}}(\varphi, V)$ step by step using our algorithm.

4.1 The Tran process

The Tran process is to transform the CTL formula into a set of $\text{SNF}_{\text{CTL}}^g$ clauses by using the rules Trans(1) to Trans(12). The transformation of an arbitrary CTL formula into the set T_φ is a sequence $T_0, T_1, \dots, T_n = T_\varphi$ of formulae with $T_0 = \{\text{AG}(\text{start} \supset$

Algorithm 1: Computing forgetting - A resolution-based method**Input:** A CTL formula φ and a set V of atoms**Output:** $F_{\text{CTL}}(\varphi, V)$

- 1 $T = \emptyset$ // the initial set of $\text{SNF}_{\text{CTL}}^g$ clauses of φ ;
- 2 $T' = \emptyset$ // the set of $\text{SNF}_{\text{CTL}}^g$ clauses without index;
- 3 $V' = \emptyset$ // the set of atoms introduced in the process of transforming φ into $\text{SNF}_{\text{CTL}}^g$ clauses;
- 4 $T, V' \leftarrow \text{Tran}(\varphi, V)$;
- 5 $\text{Res} \leftarrow \text{Res}(T, V')$;
- 6 $\Gamma \leftarrow \text{Sub}(\text{Res}, V')$;
- 7 $\Omega \leftarrow \text{Elm}(\text{EF}(\Gamma), \Gamma)$;
- 8 $\Gamma_1 \leftarrow \text{NI}(\Omega)$;
- 9 **return** $\bigwedge_{\psi \in R(\Gamma_1)_{\text{CTL}}} \psi$.

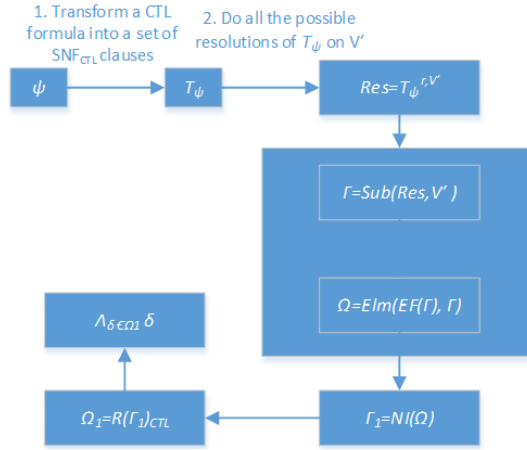


Fig. 1: The block diagram of the algorithm

$p), \text{AG}(p \supset \mathbf{simp}(\mathbf{nnf}(\varphi)))\}$ such that for every i ($0 \leq i < n$), $T_{i+1} = (T_i \setminus \{\psi\}) \cup R_i$ [22], where ψ is a formula in T_i not in $\text{SNF}_{\text{CTL}}^g$ clause and R_i is the result set of applying a matching transformation rule to ψ . Note that throughout the transformation formulae are kept in negation normal form (NNF).

Proposition 4. Let φ be a CTL formula, then $\varphi \equiv_{\langle V', I \rangle} T_\varphi$.

Proof. (sketch) This can be proved from T_i to T_{i+1} ($0 \leq i < n$) by using one transformation rule on T_i .

This means that φ has the same models with T_φ excepting that the atoms in V' and the relations $[i]$ with $i \in I$.

Algorithm 2: $Tran(\varphi)$

Input: A CTL formula φ
Output: A set T of SNF_{CTL}^g clauses and a set V' of atoms

```

1  $T = \emptyset$  // the initial set of  $SNF_{CTL}^g$  clauses of  $\varphi$ ;
2  $OldT = \{\mathbf{start} \supset z, z \supset \varphi\}$ ;
3  $V' = \{z\}$ ;
4 while  $OldT \neq T$  do
5    $OldT = T$ ;
6    $R = \emptyset$ ;
7    $X = \emptyset$ ;
8   if Chose a formula  $\psi \in OldT$  that dose not a  $SNF_{CTL}^g$  clause then
9     Using a match rule  $Rl$  to transform  $\psi$  into a set  $R$  of  $SNF_{CTL}^g$  clauses;
10     $X$  is the set of atoms introduced by using  $Rl$ ;
11     $V' = V' \cup X$ ;
12     $T = OldT \setminus \{\psi\} \cup R$ ;
13  end
14 end

```

4.2 The Resolution process

The Resolution (Res in short) process is to compute all the possible resolutions of T_φ on V' . A *derivation* on a set $V \cup V'$ of atoms and T_φ is a sequence $T_0, T_1, T_2, \dots, T_n = T_\varphi^{r, V \cup V'}$ of sets of SNF_{CTL}^g clauses such that $T_0 = T_\varphi$ and $T_{i+1} = T_i \cup R_i$ where R_i is a set of clauses obtained as the conclusion of the application of a resolution rule to premises in T_i . Note that all the T_i ($0 \leq i \leq n$) are set of SNF_{CTL}^g clauses. Besides, if there is a T_i containing $\mathbf{start} \supset \perp$ or $\top \supset \perp$, then we have $F_{CTL}(\varphi, V) = \perp$. Given two clauses C and C' , we call C and C' are resolvable, the result denote as $res(C, C')$, if there is a resolution rule using C and C' as the premises on some given atom. Then the pseudocode of algorithm Res is as Algorithm 3.

Proposition 5. Let φ be a CTL formula, then $T_\varphi \equiv_{\langle V \cup V', \emptyset \rangle} T_\varphi^{r, V \cup V'}$.

Proof. (sketch) This can be proved from T_i to T_{i+1} ($0 \leq i < n$) by using one resolution rule on T_i .

Proposition 4 and Proposition 5 mean that $\varphi \equiv_{\langle V \cup V', I \rangle} T_\varphi^{r, V \cup V'}$, this resolve the problem (1).

4.3 The Elimination process

For resolving problem (2), we should pay attention to the fact that by the transformation and resolution rules, we have the following several important properties:

- **(GNA)** for all atom p in $Var(\varphi)$, p do not positively appear in the left hand of the SNF_{CTL}^g clause;
- **(PI)** for each atom $p \in V'$, if p appearing in the left hand of a SNF_{CTL}^g clause, then p appear positively.

Algorithm 3: $Res(T, V')$

Input: A set T of SNF_{CTL}^g clauses and a set V' of atoms
Output: A set Res of SNF_{CTL}^g clauses

```

1  $S = \{C \mid C \in T \text{ and } Var(C) \cap V = \emptyset\};$ 
2  $\Pi = T \setminus S;$ 
3 for ( $p \in V \cup V'$ ) do
4    $\Pi' = \{C \in \Pi \mid p \in Var(C)\};$ 
5    $\Sigma = \Pi \setminus \Pi';$ 
6   for ( $C \in \Pi'$  s.t.  $p$  appearing in  $C$  positively) do
7     for ( $C' \in \Pi'$  s.t.  $p$  appearing in  $C'$  negatively and  $C, C'$  are resolvable)
8       do
9          $\Sigma = \Sigma \cup \{res(C, C')\};$ 
10         $\Pi' = \Pi' \cup \{C'' = res(C, C') \mid p \in Var(C'')\};$ 
11      end
12    end
13   $\Pi = \Sigma;$ 
14 end
15  $Res = \Pi \cup S;$ 

```

The Instantiation process An *instantiate formula* ψ of set V'' of atoms is a formula such that $Var(\psi) \cap V'' = \emptyset$. A key point to compute forgetting is eliminate those irrelevant atoms, for this purpose, we define the follow substitution to find out those atoms that do irrelevant.

Definition 4. [instantiation] Let $V'' = V'$ and $\Gamma = T_{\varphi}^{r, V \cup V'}$, then the process of instantiation is as follows:

- (i) for each global clause $C = \top \supset D \vee \neg p \in \Gamma$, if there is one and on one atom $p \in V'' \cap Var(C)$ and $Var(D) \cap V = \emptyset$ then let $C = p \supset D$ and $V'' := V'' \setminus \{p\}$;
- (ii) find out all the possible instantiate formulae $\varphi_1, \dots, \varphi_m$ of $V \cup V''$ in the $p \supset \varphi_i \in \Gamma$ ($1 \leq i \leq m$);
- (iii) if there is $p \supset \varphi_i$ for some $i \in \{1, \dots, m\}$, then let $V'' := V'' \setminus \{p\}$, which means p is a instantiate formula;
- (iv) for $\bigwedge_{j=1}^m p_j \supset \varphi_i \in \Gamma$ ($i \in \{1, \dots, m\}$), if there is $\alpha \supset p_1, \dots, \alpha \supset p_m \in \Gamma$ then let $\Gamma_1 := \Gamma \cup \{\alpha \supset \varphi\}$. if $\Gamma_1 \neq \Gamma$ then let $\Gamma := \Gamma_1$ go to step (i), else return $V \cup V''$.

Where p, p_i ($1 \leq i \leq m$) are atoms and α is a conjunction of literals or **start**.

We denote this process as $Sub(\Gamma, V')$, which can be described as the following Algorithm 4.

Example 2. Let $\varphi = A((p \wedge q) \cup (f \vee m)) \wedge r$ and $V = \{p\}$. Then we can compute $Sub(T_{\varphi}^{r, V \cup V'}, V)$ as follows:

Algorithm 4: Computing $\text{Sub}(\Gamma, V')$

Input: A set Γ of $\text{SNF}_{\text{CTL}}^g$ clauses φ and $V, V' \subseteq \mathcal{A}$
Output: A set of atoms

```

1 Let  $V'' := V'$ ;
2 Let  $V_1 = \emptyset$ ;
3 Let  $\Gamma_1 := \emptyset$ ;
4 Let  $\Gamma_2 := \Gamma$ ;
5 while ( $\Gamma_1 \neq \Gamma_2$  or  $V_1 \neq V''$ ) do
6    $\Gamma_1 := \Gamma_2$ ;
7    $V_1 := V''$ ;
8   for ( $C \in \Gamma_2$ ) do
9     if ( $C$  is a global clause) then
10      Let  $C := D \vee \neg p$ ;
11      if ( $p \in V'' \cap \text{Var}(C)$  and  $\text{Var}(D) \cap V == \emptyset$ ) then
12         $C := p \supset D$ ;
13         $V'' := V'' \setminus \{p\}$ ;
14      end
15    end
16  end
17  for ( $C \in \Gamma_2$ ) do
18    if ( $C == p \supset \varphi$  and  $p \in V''$  and  $\text{Var}(\varphi) \cap V \cup V'' == \emptyset$ ) then
19       $V'' := V'' \setminus \{p\}$ ;
20    end
21  end
22  for ( $C \in \Gamma_2$ ) do
23    if ( $C == \bigwedge_{j=1}^m p_j \supset \varphi$  and  $\text{Var}(\varphi) \cap V \cup V'' == \emptyset$ ) then
24      if (there is  $\alpha \supset p_1, \dots, \alpha \supset p_m \in \Gamma_2$ ) then
25         $\Gamma_2 := \Gamma_2 \cup \{\alpha \supset \varphi\}$ ;
26      end
27    end
28  end
29 end
30 return  $V \cup V''$ .

```

At first, we transform φ into a set of $\text{SNF}_{\text{CTL}}^g$ with $V' = \{x, y, z\}$, which is listed as:

- | | | |
|--|--------------------------------------|--|
| 1. $\text{start} \supset z$ | 2. $\top \supset \neg z \vee r$ | 3. $\top \supset \neg x \vee f \vee m$ |
| 4. $\top \supset \neg z \vee x \vee y$ | 5. $\top \supset \neg y \vee p$ | 6. $\top \supset \neg y \vee q$ |
| 7. $z \supset \text{AF}x$ | 8. $y \supset \text{AX}(x \vee y)$. | |

In the second, we compute all the possible resolutions on $V \cup V'$ and is listed as:

- | | | |
|---|---|---|
| (1) start $\supset r$ | (2) start $\supset x \vee y$ | (3) $\top \supset \neg z \vee y \vee f \vee m$ |
| (4) $y \supset \text{AX}(f \vee m \vee y)$ | (5) $\top \neg z \vee x \vee p$ | (6) $\top \neg z \vee x \vee q$ |
| (7) $y \supset \text{AX}(x \vee p)$ | (8) $y \supset \text{AX}(x \vee q)$ | (9) start $\supset f \vee m \vee y$ |
| (10) start $\supset x \vee p$ | (11) start $\supset x \vee q$ | (12) $\top \supset p \vee \neg z \vee f \vee m$ |
| (13) $\top \supset q \vee \neg z \vee f \vee m$ | (14) $y \supset \text{AX}(p \vee f \vee m)$ | (15) $y \supset \text{AX}(q \vee f \vee m)$ |
| (16) start $\supset f \vee m \vee p$ | (17) start $\supset f \vee m \vee q$. | |

By the process of instantiation we obtain that y is instantiated by $q \wedge \text{AX}(p \vee f \vee m)$, x is instantiated by $f \vee m$ and z is instantiated by r . That is $\text{Sub}(T_\varphi^{r, V \cup V'}, V') = V$, which means all the introduced atoms are instantiated.

By Sub operator, we guarantee those atoms in $V \cup V''$ are really irrelevant atoms.

The EF process Let P be a conjunction of literals, l, l_1 be literals, in which $\text{Var}(C_1) \cap V \cup V' = \emptyset$, and C_i ($i \in \{2, 3, 4\}$) be classical clauses. Let $\alpha = P \supset ((\neg C_3 \wedge \neg C_2) \supset (\text{E}_{\langle \text{ind} \rangle} \text{X}(C_3 \wedge \neg(C_2 \vee C_4) \supset \text{AXAF}(C_3 \vee C_2))))$ and $\beta = P \supset ((\neg C_3 \wedge \neg C_2) \supset (\text{AX}(C_3 \wedge \neg(C_2 \vee C_4) \supset \text{AXAF}(C_3 \vee C_2))))$, we add following new rules, we call it EF imply.

- (EF1) $\{P \supset \text{Afl}, P \supset \text{E}_{\langle \text{ind} \rangle} \text{X}(l_1 \vee C_4), l \supset \neg l_1 \vee C_2, l \supset C_3 \vee C_2\} \rightarrow \alpha$
- (EF2) $\{P \supset \text{Afl}, P \supset \text{AX}(l_1 \vee C_4), l \supset \neg l_1 \vee C_2, l \supset C_3 \vee C_2\} \rightarrow \beta$
- (EF3) $\{P \supset \text{E}_{\langle \text{ind} \rangle} \text{Fl}, P \supset \text{E}_{\langle \text{ind} \rangle} \text{X}(l_1 \vee C_4), l \supset \neg l_1 \vee C_2, l \supset C_3 \vee C_2\} \rightarrow \alpha$
- (EF4) $\{P \supset \text{E}_{\langle \text{ind} \rangle} \text{Fl}, P \supset \text{AX}(l_1 \vee C_4), l \supset \neg l_1 \vee C_2, l \supset C_3 \vee C_2\} \rightarrow \alpha$

By $\text{EF}(\text{Sub}(T_\varphi^{r, V \cup V'}, V'))$ we mean using (EF1) to (EF4) on $T_\varphi^{r, V \cup V'}$ and replacing $P \supset \text{AX}(\neg l \vee C_2 \vee C_4)$ with $P \supset \text{AX}(\neg l \vee C_2 \vee C_4) \vee \beta$ for rule (EF2) and replacing $P \supset \text{E}_{\langle \text{ind} \rangle} \text{X}(\neg l \vee C_2 \vee C_4)$ with $P \supset \text{E}_{\langle \text{ind} \rangle} \text{X}(\neg l \vee C_2 \vee C_4) \vee \alpha$ for other rules when l, C_2, C_3 and C_4 are instantiate formulae of $\text{Sub}(T_\varphi^{r, V \cup V'}, V')$ and $\text{Var}(l_1) \in V \cup V'$. This process can be described as Algorithm 5.

Proposition 6. Let $\Gamma = T_\varphi^{r, V \cup V'}$, we have $\Gamma \equiv_{\langle V', \emptyset \rangle} \text{EF}(\text{Sub}(\Gamma, V'))$.

Proof. It is obvious from the (EF1) to (EF4).

We prove the (EF1), for other rules can be proved similarly. Let $T_{i+1} = T_i \cup \{\varphi\}$, where $\{\varphi\}$ is obtained from T_i by using rule (EF1) on T_i , i.e. $\varphi = P \supset ((\neg C_3 \wedge \neg C_2) \supset (\text{E}_{\langle \text{ind} \rangle} \text{X}(C_3 \wedge \neg(C_2 \vee C_4) \supset \text{AXAF}(C_3 \vee C_2))))$. It is apparent that $T_{i+1} \models T_i$ and $T_i \models P \supset \text{E}_{\langle \text{ind} \rangle} \text{X}(\neg l \vee C_2 \vee C_4)$. We will show that $\forall (\mathcal{M}, s_0) \in \text{Mod}(T_i)$ there is an initial Ind-structure (\mathcal{M}', s'_0) such that $(\mathcal{M}', s'_0) \models T_{i+1}$ and $(\mathcal{M}', s'_0) \leftrightarrow_{\langle V', \emptyset \rangle} (\mathcal{M}, s_0)$

$\forall (\mathcal{M}, s) \models T_i$ we suppose $(\mathcal{M}, s) \models P \wedge \neg C_3 \wedge \neg C_2$ and $(\mathcal{M}, s_1) \models C_3 \wedge \neg C_2 \wedge \neg C_4$ with $(s, s_1) \in [\text{ind}]$ (due to other case can be proved easily). Then we

Algorithm 5: Computing $EF(\Gamma, V)$

Input: A set Γ of SNF_{CTL}^g clauses, a set of A-step clauses and a set of E-step clauses

Output: A set of formulae

```

1 for ( $C \in A$ ) do
2   Let  $C == P \supset AFl$ ;
3   if ( $P \supset E_{\langle ind \rangle} X(l_1 \vee C_4), l \supset \neg l_1 \vee C_2, l \supset C_3 \vee C_2 \in \Gamma$  and  $l, C_2, C_3, C_4$ 
   are instantiate formulae) then
4     Replacing  $P \supset E_{\langle ind \rangle} X(\neg l \vee C_2 \vee C_4)$  with
        $P \supset E_{\langle ind \rangle} X(\neg l \vee C_2 \vee C_4) \vee \alpha$ ;
5   end
6   if ( $P \supset AX(l_1 \vee C_4), l \supset \neg l_1 \vee C_2, l \supset C_3 \vee C_2 \in \Gamma$  and  $l, C_2, C_3, C_4$  are
   instantiate formulae) then
7     Replacing  $P \supset AX(\neg l \vee C_2 \vee C_4)$  with  $P \supset AX(\neg l \vee C_2 \vee C_4) \vee \beta$ ;
8   end
9 end
10 for ( $C \in E$ ) do
11   Let  $C == P \supset E_{\langle ind \rangle} Fl$ ;
12   if ( $P \supset E_{\langle ind \rangle} X(l_1 \vee C_4), l \supset \neg l_1 \vee C_2, l \supset C_3 \vee C_2 \in \Gamma$  or  $P \supset$ 
    $AX(l_1 \vee C_4), l \supset \neg l_1 \vee C_2, l \supset C_3 \vee C_2 \in \Gamma$  and  $l, C_2, C_3, C_4$  are
   instantiate formulae) then
13     Replacing  $P \supset E_{\langle ind \rangle} X(\neg l \vee C_2 \vee C_4)$  with
        $P \supset E_{\langle ind \rangle} X(\neg l \vee C_2 \vee C_4) \vee \alpha$ ;
14   end
15 end
16 return  $\Gamma$ .

```

have $(\mathcal{M}, s) \not\models l$ (by $(\mathcal{M}, s) \models l \supset C_3 \vee C_2$) and $(\mathcal{M}, s_1) \models l_1$ (by $(\mathcal{M}, s) \models P \supset E_{\langle ind \rangle} X(l_1 \vee C_4)$). If $(\mathcal{M}, s_1) \not\models AXAF(C_3 \vee C_2)$ then we have $(\mathcal{M}, s_1) \models l$ due to $(\mathcal{M}, s) \models AG(l \supset C_3 \vee C_2)$ and $(\mathcal{M}, s) \models AFl$. And then $(\mathcal{M}, s_1) \models \neg l_1$ by $(\mathcal{M}, s) \models AG(l \supset \neg l_1 \vee C_2)$. It is contract with our supposing. Then $(\mathcal{M}, s_1) \models AXAF(C_3 \vee C_2)$.

The Elimination process For eliminate those irrelevant atoms, we can do the following elimination operator.

Definition 5. (Elimination) Let T be a set of formulae, $C \in T$ and V a set of atoms, then the elimination operator, denoted as Elm , is defined as:

$$Elm(C, V) = \begin{cases} \top, & \text{if } \text{Var}(C) \cap V \neq \emptyset \\ C, & \text{else.} \end{cases}$$

For convenience, we let $Elm(T, V) = \{Elm(r, V) | r \in T\}$.

Proposition 7. *Let $V'' = V \cup V'$, $\Gamma = \text{Sub}(T_{\varphi}^{r,V''}, V')$ and $\Gamma_1 = \text{Elm}(\text{EF}(\Gamma), \Gamma)$, then $\Gamma_1 \equiv_{\langle V \cup V', \emptyset \rangle} T_{\varphi}^{r,V''}$ and $\Gamma_1 \equiv_{\langle V \cup V', I \rangle} \varphi$.*

Proof. Note the fact that for each clause $C = T \supset H$ in $\text{EF}(\Gamma)$, if $\Gamma \cap \text{Var}(C) \neq \emptyset$ then there must be an atom $p \in \Gamma \cap \text{Var}(H)$. It is apparent that $\text{EF}(\Gamma) \models \Gamma_1$, we will show $\forall (\mathcal{M}, s_0) \in \text{Mod}(\Gamma_1)$ there is a (\mathcal{M}', s_0) such that $(\mathcal{M}', s_0) \models \text{EF}(\Gamma)$ and $(\mathcal{M}, s_0) \leftrightarrow_{\langle \Gamma, \emptyset \rangle} (\mathcal{M}', s_0)$. Let $C = T \supset H$ in $\text{EF}(\Gamma)$ with $\Gamma \cap \text{Var}(C) \neq \emptyset$, $\forall (\mathcal{M}, s_0) \in \text{Mod}(\Gamma_1)$ we construct (\mathcal{M}', s_0) as (\mathcal{M}, s_0) except for each $s \in S$, if $(\mathcal{M}, s) \models T$ then $L'(s) = L(s)$, else:

- (i) if $(\mathcal{M}, s) \models H$, then $L'(s) = L(s)$;
- (ii) else if $(\mathcal{M}, s) \models T$ with $p \in \text{Var}(H) \cap V$, then if p appearing in H negatively, then if C is a global (or an initial) clause then let $L'(s) = L(s) \setminus \{p\}$ else let $L'(s_1) = L(s_1) \setminus \{p\}$ for (each (if C is an A-step or A-sometime clause)) $(s, s_1) \in R$, else if C is a global (or an initial) clause then let $L'(s) = L(s) \cup \{p\}$ else let $L'(s_1) = L(s_1) \cup \{p\}$ for (each (if C is an A-step or A-sometime clause)) $(s, s_1) \in R$.
- (iii) for other clause $C = Q \supset H$ with $p \in \text{Var}(H) \cap \Gamma$, we can do it as (ii).

It is apparent that $(\mathcal{M}, s_0) \leftrightarrow_{\langle \Gamma, \emptyset \rangle} (\mathcal{M}', s_0)$, we will show that $(\mathcal{M}', s_0) \models \text{EF}(\Gamma)$ from the following two points:

- (1) For (ii) talked-above, we show it from the form of $\text{SNF}_{\text{CTL}}^g$ clauses. Supposing C_1 and C_2 are instantiate formula of Γ :
 - (a) If C is a global clause, i.e. $C = \top \supset p \vee C_1$ with C_1 is a disjunction of literals (we suppose p appearing in C positively). If there is a $C' = \top \supset \neg p \vee C_2 \in \text{EF}(\Gamma)$, then there is $\top \supset C_1 \vee C_2 \in \text{EF}(\Gamma)$ by the resolution $((\mathcal{M}, s) \models C_2$ due to we have suppose $(\mathcal{M}, s) \models C$). It is apparent that $(\mathcal{M}', s_0) \models C \wedge C'$.
 - (b) If $C = T \supset E_{\langle \text{ind} \rangle} X(p \vee C_1)$. If there is a $C' = T' \supset E_{\langle \text{ind} \rangle} X(\neg p \vee C_2) \in \text{EF}(\Gamma)$, then there is $T \wedge T' \supset E_{\langle \text{ind} \rangle} X(C_1 \vee C_2) \in \text{EF}(\Gamma)$ by the resolution $((\mathcal{M}, s) \models E_{\langle \text{ind} \rangle} X C_2$ due to we have suppose $(\mathcal{M}, s) \models C$). It is apparent that $(\mathcal{M}', s_0) \models C \wedge C'$.
 - (c) Other cases can be proved similarly.
- (2) (iii) can be proved as (ii) due to the fact we point at the beginning.

Therefore, we have $\Gamma_1 \equiv_{\langle V \cup V', \emptyset \rangle} T_{\varphi}^{r,V''}$ by Proposition 2 and Proposition 6.

And then $\Gamma_1 \equiv_{\langle V \cup V', I \rangle} \varphi$ follows.

4.4 Remove the Index and start

The $\text{NI}(\Gamma)$ process is to change the set Γ of $\text{SNF}_{\text{CTL}}^g$ into a set of formulas without the index by using the equations in Proposition 3.

Lemma 1. (NI-BRemain) *Let T be a set of $\text{SNF}_{\text{CTL}}^g$ clauses and T' be the nonInd- $\text{SNF}_{\text{CTL}}^g$ of T . If T is satisfiable, then we have $T \equiv_{\langle \emptyset, I \rangle} \text{NI}(T)$, where I is the set of indexes in T .*

Proof. It is easy checking that from the definition of NI.

Similarly, let T be a set of $\text{SNF}_{\text{CTL}}^g$ clauses, then we define the following operator:

$$T_{\text{CTL}} = \{C | C' \in T \text{ and } C = D \text{ if } C' \text{ is the form } \text{AG}(\text{start} \supset D), \text{ else } C = C'\}.$$

Then $T \equiv T_{\text{CTL}}$ by $\varphi \equiv \text{AG}(\text{start} \supset \varphi)$ [3].

The last step of our algorithm is to eliminate all the atoms in V' which has been introduced in the process Tran. Let $V'' = V \cup V'$, $\Gamma = \text{Sub}(T_{\varphi}^{r, V''}, V')$ and $\Gamma_1 = \text{Elm}(\text{EF}(\Gamma))$, then $R(\text{NI}(\Gamma_1))$ is obtained from $\text{NI}(\Gamma_1)$ by doing the following two steps for each $p \in (V' \setminus \Gamma) \cup V^F$:

- replacing each $p \supset \varphi_1 \vee \dots \vee p \supset \varphi_n$ with $p \supset \bigvee_{i \in \{1, \dots, n\}} \varphi_i$;
- replacing $p \supset \varphi_1 \wedge \dots \wedge p \supset \varphi_m$, φ_j are instantiate formulae of Γ ($j \in \{1, \dots, m\}$), then let $\psi = \bigwedge_{i=1}^{j_n} \varphi_{j_i}$, where p do not appear in φ_{j_i} , with $p \leftrightarrow \psi$.
- For other formula $C \in \Omega_1$, replacing every p in C with ψ .

Where $\text{NI}(S)$ means do $\text{NI}(e)$ for each $e \in S$ with S is a set of sets of $\text{SNF}_{\text{CTL}}^g$ clause. Apparently, this process is just a process of replacing each atom with an equivalent formula. Then we have:

Proposition 8. Let $\Gamma = T_{\varphi}^{r, V \cup V'}$, $\Gamma_1 = \text{Sub}(\Gamma, V')$ and $\Gamma_2 = \text{Elm}(\text{EF}(\Gamma_1), \Gamma_1)$, then $\Gamma_2 \equiv_{\langle V' \setminus \Gamma_1, \emptyset \rangle} R(\text{NI}(\Gamma_2))$. and $\varphi \equiv_{\langle V \cup V', I \rangle} R(\text{NI}(\Gamma_2))_{\text{CTL}}$.

Proof. For each p talked above is a name of the formula ψ , i.e. $p \leftrightarrow \psi$. Then $\Gamma_2 \equiv_{\langle (V' \setminus \Gamma_1) \emptyset \rangle} R(\Gamma_2)$, and then $\Gamma_2 \equiv_{\langle V \cup V', \emptyset \rangle} R(\Gamma_2)$ by (V) of Proposition 1.

Therefore, $\varphi \equiv_{\langle V \cup V', I \rangle} \text{NI}(R(\Gamma_2))_{\text{CTL}}$ by Proposition 7 and the definitions of NI and T_{CTL} .

4.5 The Correction and Complexity of the Algorithm

In the case that formula dose not include index, we use model structure $\mathcal{M} = (S, R, L, s_0)$ to interpret formula instead of Ind-model structure. Therefore it is apparent that $\forall (\mathcal{M}, s_0) \in \text{Mod}(\varphi)$ there is a $(\mathcal{M}', s'_0) \in \text{Mod}(\Gamma_1)$ such that $(\mathcal{M}, s_0) \leftrightarrow_{V \cup V'} (\mathcal{M}', s'_0)$ and vice versa.

Theorem 1. Let $V'' = V \cup V'$, $\Gamma = \text{Sub}(T_{\varphi}^{r, V''}, V')$ and $\Gamma_1 = R(\text{NI}(\text{Elm}(\text{EF}(\Gamma), \Gamma)))_{\text{CTL}}$, then

$$\text{F}_{\text{CTL}}(\varphi, V' \cup V) \equiv \Gamma_1.$$

Proof. $(\Rightarrow) \forall (\mathcal{M}, s_0) \in \text{Mod}(\text{F}_{\text{CTL}}(\varphi, V' \cup V))$
 $\Rightarrow \exists (\mathcal{M}', s'_0) \in \text{Mod}(\varphi)$ s.t. $(\mathcal{M}, s_0) \leftrightarrow_{V' \cup V} (\mathcal{M}', s'_0)$
 $\Rightarrow \exists (\mathcal{M}_1, s_1) \in \text{Mod}(\text{Elm}(\text{Sub}((T_{\varphi}^{r, V \cup V'})', V)_{\text{CTL}}, V \cup V'))$ s.t. $(\mathcal{M}_1, s_1) \leftrightarrow_{V' \cup V} (\mathcal{M}', s'_0)$
 $\Rightarrow (\mathcal{M}, s_0) \leftrightarrow_{V' \cup V} (\mathcal{M}_1, s_1)$
 $\Rightarrow (\mathcal{M}, s_0) \models \text{Elm}(\text{Sub}((T_{\varphi}^{r, V \cup V'})', V)_{\text{CTL}}, V \cup V') (\text{IR}(\text{Elm}(\text{Sub}((T_{\varphi}^{r, V \cup V'})', V)_{\text{CTL}}, V \cup V'), V' \cup V))$
 $(\Leftarrow) \forall (\mathcal{M}_1, s_1) \in \text{Mod}(\text{Elm}(\text{Sub}((T_{\varphi}^{r, V \cup V'})', V)_{\text{CTL}}, V \cup V'))$
 $\Rightarrow \exists (\mathcal{M}', s'_0) \in \text{Mod}(\varphi)$ s.t. $(\mathcal{M}_1, s_1) \leftrightarrow_{V' \cup V} (\mathcal{M}', s'_0)$
 $\Rightarrow (\mathcal{M}_1, s_1) \models \text{F}_{\text{CTL}}(\varphi, V' \cup V)$ $(\text{IR}(\text{F}_{\text{CTL}}(\varphi, V' \cup V), V \cup V'))$ and
 $\varphi \models \text{F}_{\text{CTL}}(\varphi, V' \cup V)$

Then we have the following result:

Theorem 2. (Resolution-based CTL-forgetting) Let $V'' = V \cup V'$, $\Gamma = \text{Sub}(T_\varphi^{r,V''}, V')$ and $\Gamma_1 = R(\text{NI}(\text{Elm}(\text{EF}(\Gamma), \Gamma)))_{\text{CTL}}$, then

$$F_{\text{CTL}}(\varphi, V) \equiv \bigwedge_{\psi \in \Gamma_1} \psi.$$

We can obtain that $F_{\text{CTL}}(\varphi, V) \equiv F_{\text{CTL}}(\varphi, V' \cup V)$ by Theorem 1, Proposition ?? and Proposition ?. Therefore, the Theorem 2 is proved.

Then we can obtain the result of forgetting of Example 2:

$$\begin{aligned} F_{\text{CTL}}(\varphi, \{p\}) &\equiv r \wedge (f \vee m \vee q) \wedge \\ & (f \vee m \vee (q \wedge \text{AX}(f \vee m \vee q))) \wedge \text{AG}((q \wedge \text{AX}(f \vee m \vee q)) \\ & \supset \text{AX}(f \vee m \vee (q \wedge \text{AX}(f \vee m \vee q)))). \end{aligned}$$

Proposition 9. Let φ be a CTL formula and $V \subseteq \mathcal{A}$. The time and space complexity of Algorithm 1 are $O((m+1)2^{4(n+n')})$. Where $|\text{Var}(\varphi)| = n$, $|V'| = n'$ (V' is set of atoms introduced in transformation) and m is the number of the set Ind of indices introduced during transformation.

Proof. It follows from that the lines 19-31 of the algorithm, which is to compute all the possible resolution. The possible number of $\text{SNF}_{\text{CTL}}^g$ clauses under the give V , V' and Ind is $(m+1)2^{4(n+n')} + (m * (n+n') + n + n' + 1)2^{2(n+n')+1}$.

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