

# A Resolution Calculus for Forgetting in CTL

Renyan Feng<sup>1,2</sup>, Erman Acar<sup>2</sup>, Stefan Schlobach<sup>2</sup>, Yisong Wang<sup>1</sup>

<sup>1</sup>Guizhou University, P. R. China

<sup>2</sup>Vrije Universiteit Amsterdam, Netherlands

fengrenyan@gmail.com, {k.s.schlobach, erman.acar}@vu.nl, yswang@gzu.edu.cn

## Abstract

Computation Tree Logic (CTL) is one of the central logical formalisms in computer science with a wide-range of applications; it is used mostly in formal verification in the context of representing and reasoning about high-level system information (or *specification*), but also in other domains e.g., planning. Orthogonal to this, forgetting is the field of study which concerns removing such a piece of information, deemed irrelevant or obsolete, from a knowledge base (e.g., a set of such specifications) while guaranteeing certain properties.

In this paper, we present a resolution-based approach to perform *forgetting* in CTL. In particular, we develop a resolution calculus which extends an earlier work (CTL with *index* i.e.,  $\text{SNF}_{\text{CTL}}^g$ ) with additional rules i.e., EF-implication which connects *next state* and *future state*. Since tailoring a resolution calculus for forgetting is a challenging task, our technical contribution is manifolded: (i) we provide with a bisimulation between CTL formulae and  $\text{SNF}_{\text{CTL}}^g$  clauses; (ii) we introduce techniques for eliminating undesired atoms resulting from such transformation. Moreover, we show the soundness of our approach and analyse its computational complexity.

## 1 Introduction

*Model Checking* is an automatic, model-based, property-verification approach. In the past decades, it not only be used in *verification* but also in *planning* (Giunchiglia and Traverso 1999; Akintunde 2017). Computation Tree Logic (CTL) (Clarke and Emerson 1981) is one of the central logical formalisms in computer science with a wide-range of applications; it is used mostly in formal verification in the context of representing and reasoning about high-level system information (or *specification*), but also in other domains e.g., planning. In those cases, as the system or planning domain has been updated-some of the elements previously considered are no longer considered-if we have had those specifications, we do not need to reproduce the specifications that different with the original one only in that the latter using lesser atoms. Besides, with the size of the system or planning domain growing, not only the number of available (proposition) increased considerably, but those specifications are often large in size and becoming more complex. This leads to the specification being difficult to maintain and modify, and costly to reuse for later processing, where only a specific part of an specification is of interest. All of

this working directly on the whole of the original specification and building a new sub-specification are unadvisable. Therefore, a strong demand for techniques and automated tools for obtaining the specific sub-specification.

*Forgetting*, as a technology distilling from a knowledge base only the part that is relevant to a subset of alphabet, just happens to solve the problem talked above. As a logical notion, *forgetting* was first formally defined in propositional and first order-logics by Lin and Reiter (Lin and Reiter 1994). 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 (Eiter and Kern-Isberner 2019), such as forgetting in logic programs under answer set/stable model semantics (Zhang and Foo 2006; Eiter and Wang 2008; Wong 2009; Wang et al. 2012; Wang, Wang, and Zhang 2013), forgetting in description logic (Wang et al. 2010; Lutz and Wolter 2011; Zhao and Schmidt 2017a) and knowledge forgetting in modal logic (Zhang and Zhou 2009; Su et al. 2009; Liu and Wen 2011; Fang, Liu, and Van Ditmarsch 2019). In application, forgetting has been used in planning (Lin 2003), conflict solving (Lang and Marquis 2010; Zhang, Foo, and Wang 2005), creating restricted views of ontologies (Zhao and Schmidt 2017a), strongest and weakest definitions (Lang and Marquis 2008), SNC (WSC) (Lin 2001) and so on.

Though forgetting has been extensively investigated from various aspects of different logical systems. The existing forgetting methods in propositional logic, answer set programming, description logic and modal logic are not directly applicable in CTL. In this paper, we give the definition of forgetting in CTL from the semantic forgetting point of view and explore a resolution-based method to compute it.

For one thing, the existing algorithm of computing forgetting in different logics talked above are not directly applicable in CTL. For instance, in propositional forgetting theory, forgetting atom  $q$  from  $\varphi$  is equivalent to a formula  $\varphi[q/\top] \vee \varphi[q/\perp]$ , where  $\varphi[q/X]$  is a formula obtained from  $\varphi$  by replacing each  $q$  with  $X$  ( $X \in \{\top, \perp\}$ ). This method cannot be extended to a CTL formula. Consider a CTL formula  $\psi = \text{AG}p \wedge \neg \text{AG}q \wedge \neg \text{AG}\neg q$ . If we want to forget atom  $q$  from  $\psi$  by using the above method, we would have  $\psi[q/\top] \vee \psi[q/\perp] \equiv \perp$ . This is obviously not correct since after forgetting  $q$  this specification should not become

inconsistent.

For another, as far as I know the existing methods to compute forgetting include the classical one talked above and resolution-based approaches in propositional logic (Lin and Reiter 1994; Wang 2015) and Ackermann-based approach (second-order elimination) in description logic (Zhao and Schmidt 2017b). However, the resolution and Ackermann-based methods need a specific normal form of the formula, it is hard to obtain such normal form in CTL. Although any CTL formula can be transformed into a set of  $\text{SNF}_{\text{CTL}}^g$  clauses, but it do introduce the *index* and new atoms. Both the two problems are we should solve.

In this paper we extend the Resolution Calculus in (Zhang, Hustadt, and Dixon 2014) by eliminating the atoms introduced in the transformation process and combining the CTL with  $\text{SNF}_{\text{CTL}}^g$  by using the *binary bisimulation relation* (one is the set of atoms and another one is the set of indexes). Such a bisimulation relation is an extension of the set-based bisimulation talked in “the theory paper” by taking *index* into account.

The paper is structured as follows: Section 2 introduces the notation and technical preliminaries. In section 3 we give a more precise definition of the problem. As key contributions, Section 4, introduces the resolution-based approach. Conclusion closes the paper.

## 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  $\mathcal{A}$ .

### 2.1 Model structure in CTL

In general, a transition system can be described by a *model structure* (or *Kripke structure*) (see (Baier and Katoen 2008) for details). A model structure is a triple  $\mathcal{M} = (S, R, L)$ , where

- $S$  is a finite nonempty set of states,
- $R \subseteq S \times S$  and, for each  $s \in S$ , there is  $s' \in S$  such that  $(s, s') \in R$ ,
- $L$  is a labeling function  $S \rightarrow 2^{\mathcal{A}}$ .

Given a model structure  $\mathcal{M} = (S, R, L)$ , a *path*  $\pi_{s_i}$  starting from  $s_i$  of  $\mathcal{M}$  is an infinite sequence of states  $\pi_{s_i} = (s_i, s_{i+1}, s_{i+2}, \dots)$ , where for each  $j$  ( $0 \leq i \leq j$ ),  $(s_j, s_{j+1}) \in R$ . By  $s' \in \pi_{s_i}$  we mean that  $s'$  is a state in the path  $\pi_{s_i}$ . A state  $s \in S$  is *initial* if for any state  $s' \in S$ , there is a path  $\pi_s$  s.t  $s' \in \pi_s$ . If  $s_0$  is an initial state of  $\mathcal{M}$ , then we denote this model structure  $\mathcal{M}$  as  $(S, R, L, s_0)$ .

For a given model structure  $\mathcal{M} = (S, R, L, s_0)$  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 (Browne, Clarke, and Grumberg 1988), 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  $\text{Tr}_n(s')$ .

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  $s = s_0$  is an initial state of  $\mathcal{M}$ , the K-structure is *initial*.

### 2.2 Syntax and semantics of CTL

In the following we briefly review the basic syntax and semantics of the CTL (Clarke, Emerson, and Sistla 1986). The *signature* of the language  $\mathcal{L}$  of CTL includes:

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

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

where  $p \in \mathcal{A}$ . The formulas  $\phi \wedge \psi$  and  $\phi \supset \psi$  are defined in a standard manner of propositional logic. The other form formulas of  $\mathcal{L}$  are abbreviated using the forms of (1).

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

- $(\mathcal{M}, s) \not\models \perp$  and  $(\mathcal{M}, s) \models \top$ ;
- $(\mathcal{M}, s) \models p$  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  $i \geq 1$ ;
- $(\mathcal{M}, s) \models E[\phi_1 \cup \phi_2]$  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  $1 \leq j < i$ .

Similar to the work in (Browne, Clarke, and Grumberg 1988; Bolotov 1999), only initial K-structures are considered to be candidate models in the following, unless otherwise noted. Formally, an initial K-structure  $\mathcal{K}$  is a *model* of a formula  $\phi$  whenever  $\mathcal{K} \models \phi$ . We denote  $\text{Mod}(\phi)$  the set of models of  $\phi$ . The formula  $\phi$  is *satisfiable* if  $\text{Mod}(\phi) \neq \emptyset$ . Given two formulas  $\phi_1$  and  $\phi_2$ ,  $\phi_1 \models \phi_2$  we mean  $\text{Mod}(\phi_1) \subseteq \text{Mod}(\phi_2)$ , and by  $\phi_1 \equiv \phi_2$ , we mean  $\phi_1 \models \phi_2$  and  $\phi_2 \models \phi_1$ . In this case  $\phi_1$  is *equivalent* to  $\phi_2$ . The set of atoms occurring in  $\phi_1$ , is denoted by  $\text{Var}(\phi_1)$ .  $\phi_1$  is *V-irrelevant*, written  $\text{IR}(\phi_1, V)$ , if there is a formula  $\psi$  with  $\text{Var}(\psi) \cap V = \emptyset$  such that  $\phi_1 \equiv \psi$ .

### 2.3 The normal form of CTL

It has proved that any CTL formula  $\varphi$  can be transformed into a set  $T_\varphi$  of  $\text{SNF}_{\text{CTL}}^g$  (Separated Normal Form with Global Clauses for CTL) clauses in polynomial time such that  $\varphi$  is satisfiable iff  $T_\varphi$  is satisfiable (Zhang, Hustadt, and Dixon 2008). 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  $\text{Ind}$ ; and (4) temporal operators:  $E_{\langle \text{ind} \rangle} X$ ,  $E_{\langle \text{ind} \rangle} F$ ,  $E_{\langle \text{ind} \rangle} G$ , and  $E_{\langle \text{ind} \rangle} U$ .

Before talk about the semantic 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.

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

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  $\text{ind}$  is an element of  $\text{Ind}$  ( $\text{Ind}$  is a countably infinite index set). By clause we mean the classical clause or the  $\text{SNF}_{\text{CTL}}^g$  clause unless explicitly stated. As all clauses are of the form  $\text{AG}(P \supset D)$ , we often simply write  $P \supset D$  instead.

Formulae of  $\text{SNF}_{\text{CTL}}^g$  over  $\mathcal{A}$  are interpreted in Ind-model structure  $\mathcal{M} = (S, R, L, [\cdot], s_0)$ , where  $S$ ,  $R$ ,  $L$  and  $s_0$  is the same as our model structure talked above and  $[\cdot] : \text{Ind} \rightarrow 2^{(S \times S)}$  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  $(s, s') \in [\text{ind}]$  and  $(s, s') \in R$ . An infinite path  $\pi_{s_i}^{(\text{ind})}$  is an infinite sequence of states  $s_i, s_{i+1}, s_{i+2}, \dots$  such that for every  $j \geq i$ ,  $(s_j, s_{j+1}) \in [\text{ind}]$ .

Similarly, an *Ind-structure* (or *Ind-interpretation*) is a Ind-model structure  $\mathcal{M} = (S, R, L, [\cdot], s_0)$  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 then defined as shown next as an extension of the semantics of CTL. Let  $\varphi$  and  $\psi$  be two  $\text{SNF}_{\text{CTL}}^g$  formulae and  $\mathcal{M} = (S, R, L, [\cdot], s_0)$  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  $s_i = s_0$ ;
- $(\mathcal{M}, s_i) \models E_{\langle \text{ind} \rangle} X\psi$  iff for the path  $\pi_{s_i}^{(\text{ind})}$ ,  $(\mathcal{M}, s_{i+1}) \models \psi$ ;
- $(\mathcal{M}, s_i) \models E_{\langle \text{ind} \rangle} G\psi$  iff for every  $s_j \in \pi_{s_i}^{(\text{ind})}$ ,  $(\mathcal{M}, s_j) \models \psi$ ;
- $(\mathcal{M}, s_i) \models E_{\langle \text{ind} \rangle} [\varphi U \psi]$  iff there exists  $s_j \in \pi_{s_i}^{(\text{ind})}$  such that  $(\mathcal{M}, s_j) \models \psi$  and for every  $s_k \in \pi_{s_i}^{(\text{ind})}$ , if  $i \leq k < j$ , then  $(\mathcal{M}, s_k) \models \varphi$ ;
- $(\mathcal{M}, s_i) \models E_{\langle \text{ind} \rangle} F\psi$  iff  $(\mathcal{M}, s_i) \models E_{\langle \text{ind} \rangle} [\top U \psi]$ .

The semantics of the remaining operators is analogous to that of CTL given previously but in the extended Ind-model structure  $\mathcal{M} = (S, R, L, [\cdot], s_0)$ . A  $\text{SNF}_{\text{CTL}}^g$  formula  $\varphi$  is satisfiable, iff for some Ind-model structure  $\mathcal{M} = (S, R, L, [\cdot], s_0)$ ,  $(\mathcal{M}, s_0) \models \varphi$ , and unsatisfiable otherwise. And if  $(\mathcal{M}, s_0) \models \varphi$  then  $(\mathcal{M}, s_0)$  is called an Ind-model of  $\varphi$ , and we say that  $(\mathcal{M}, s_0)$  satisfies  $\varphi$ . 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 CTL sub-section.

### 3 Problem Definition

In this section, we present the notion of forgetting in CTL. For convenience, in the following we denote  $\mathcal{M} = (S, R, L, s_0)$ ,  $\mathcal{M}' = (S', R', L', s'_0)$ ,  $\mathcal{M}_i = (S_i, R_i, L_i, s_i)$  (or  $\mathcal{M} = (S, R, L, [\cdot], s_0)$ ,  $\mathcal{M}' = (S', R', L', [\cdot], s'_0)$ ,  $\mathcal{M}_i = (S_i, R_i, L_i, [\cdot], s_i)$ ) and  $\mathcal{K}_i = (\mathcal{M}_i, s_i)$  with  $s_i \in S_i$  and  $i \in \mathbb{N}$ .

Let  $\mathcal{K}_i = (\mathcal{M}_i, s_i)$  with  $i \in \{1, 2\}$ ,

- $(\mathcal{K}_1, \mathcal{K}_2) \in \mathcal{B}_0$  if  $L_1(s_1) - V = L_2(s_2) - V$ ;
- for  $n \geq 0$ ,  $(\mathcal{K}_1, \mathcal{K}_2) \in \mathcal{B}_{n+1}$  if:
  - $(\mathcal{K}_1, \mathcal{K}_2) \in \mathcal{B}_0$ ,
  - for every  $(s_1, s'_1) \in R_1$ , there is a  $(s_2, s'_2) \in R_2$  such that  $(\mathcal{K}'_1, \mathcal{K}'_2) \in \mathcal{B}_n$ , and
  - for every  $(s_2, s'_2) \in R_2$ , there is a  $(s_1, s'_1) \in R_1$  such that  $(\mathcal{K}'_1, \mathcal{K}'_2) \in \mathcal{B}_n$ ,

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

Now, we define the notion of  $V$ -bisimulation between  $\mathcal{K}$ -structures:

**Definition 1** ( $V$ -bisimulation). *Let  $V \subseteq \mathcal{A}$ . Given two  $\mathcal{K}$ -structures (or Ind-structures)  $\mathcal{K}_1$  and  $\mathcal{K}_2$  are  $V$ -bisimilar, denoted  $\mathcal{K}_1 \leftrightarrow_V \mathcal{K}_2$  if and only if  $(\mathcal{K}_1, \mathcal{K}_2) \in \mathcal{B}_i$  for all  $i \geq 0$ . Moreover, two paths  $\pi_i = (s_{i,1}, s_{i,2}, \dots)$  of  $\mathcal{M}_i$  with  $i \in \{1, 2\}$  are  $V$ -bisimilar if  $\mathcal{K}_{1,j} \leftrightarrow_V \mathcal{K}_{2,j}$  for every  $j \geq 1$  where  $\mathcal{K}_{i,j} = (\mathcal{M}_i, s_{i,j})$ .*

It is apparent that  $\leftrightarrow_V$  is a binary relation. In the sequel, we abbreviate  $\mathcal{K}_1 \leftrightarrow_V \mathcal{K}_2$  by  $s_1 \leftrightarrow_V s_2$  whenever the underlying model structures of states  $s_1$  and  $s_2$  are clear from the context.

**Lemma 1.** *The relation  $\leftrightarrow_V$  is an equivalence relation.*

Besides, we have the following properties:

**Proposition 1.** *Let  $i \in \{1, 2\}$ ,  $V_1, V_2 \subseteq \mathcal{A}$ , 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)  $\mathcal{K}_1 \leftrightarrow_{V_1 \cup V_2} \mathcal{K}_3$ ;
- (ii) If  $V_1 \subseteq V_2$  then  $\mathcal{K}_1 \leftrightarrow_{V_2} \mathcal{K}_2$ .

Intuitively, if two  $\mathcal{K}$ -structures are  $V$ -bisimilar, then they satisfy the same formula  $\varphi$  that dose not contain any atoms in  $V$ , i.e.  $\text{IR}(\varphi, V)$ .

**Theorem 1.** *Let  $V \subseteq \mathcal{A}$ ,  $\mathcal{K}_i$  ( $i = 1, 2$ ) be two  $\mathcal{K}$ -structures such that  $\mathcal{K}_1 \leftrightarrow_V \mathcal{K}_2$  and  $\phi$  a formula with  $\text{IR}(\phi, V)$ . Then  $\mathcal{K}_1 \models \phi$  if and only if  $\mathcal{K}_2 \models \phi$ .*

*Proof.* (sketch) This can be proved by induction on the structures of  $\phi$ . For instance, let  $\phi = \psi_1 \vee \psi_2$ , the induction hypothesis is  $\mathcal{K}_1 \models \psi_i$  iff  $\mathcal{K}_2 \models \psi_i$  with  $i \in \{1, 2\}$ . Then we can see that  $\mathcal{K}_1 \models \phi$  iff  $\mathcal{K}_1 \models \psi_1$  or  $\mathcal{K}_1 \models \psi_2$  iff  $\mathcal{K}_2 \models \psi_1$  or  $\mathcal{K}_2 \models \psi_2$  by induction hypothesis.  $\square$

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  $\phi$  a CTL formula. A CTL formula  $\psi$  with  $\text{Var}(\psi) \cap V = \emptyset$  is a result of forgetting  $V$  from  $\phi$ , if*

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

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

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

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

**Definition 3** (binary bisimulation relation). *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  $s_1 \leftrightarrow_V s'_1$ .

We call this relation as *binary bisimulation relation*, also denoted as  $\leftrightarrow_{\langle V, I \rangle}$ . 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 initial 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 (ii) of Proposition 1.  $\square$

## 4 The Calculus

*Resolution* in CTL is a method to decide the satisfiability of a CTL formula. In this part, we will explore a resolution-based method to compute forgetting in CTL. We use the transformation rules Trans(1) to Trans(12) and resolution rules (SRES1), ..., (SRES8), RW1, RW2, (ERES1), (ERES2) in (Zhang, Hustadt, and Dixon 2009). Due to space restrictions, these rules are not enumerated here.

The key problems of this method include: (1) How to fill the gap between CTL and  $\text{SNF}_{\text{CTL}}^g$  since there is index for existential quantifier in  $\text{SNF}_{\text{CTL}}^g$ ; and (2) How to eliminate the irrelevant atoms, which we want to forget and introduced by the transformation rules, in the formula. We will resolve these two problems by  $\langle V, I \rangle$ -bisimulation and *eliminate* operator respectively. 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 introduced in the transformation and resolution processes below,  $\varphi$  the CTL formula,  $T_\varphi$  be the set of  $\text{SNF}_{\text{CTL}}^g$  clauses obtained from  $\varphi$  by using transformation rules on it and  $\mathcal{M} = (S, R, L, [-], s_0)$  unless explicitly state. Let  $T, T'$  be two sets of formulae,  $I$  a set of indexes introduced in the transformation 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 algorithm of computing the forgetting in CTL is as Algorithm 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 Transform process), and then compute all the possible resolutions on the specified set of atoms (the Resolution process). Third, eliminating, which include *Instantiate*, *Connect* and *Removing\_atoms* sub-processes, all the irrelevant atoms. Changing the result obtained before into a CTL formula at last, this include three sub-processes: *Removing\_index* (removing the index in the formula), *Replacing\_atoms* (replacing the atoms in  $V'$  with a formula) and  $T_{\text{CTL}}$  (removing the **start** in  $T$ ). To describe our algorithm clearly, we illustrate it with the following example.

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

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

### 4.1 The Transform process

The *Transform* process, denoted as  $\text{Transform}(\varphi)$ , 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) in (Zhang, Hustadt, and Dixon 2009)).

The *transformation* of any CTL formula  $\varphi$  into the set  $T_\varphi$  is a sequence  $T_0, T_1, \dots, T_n = T_\varphi$  of sets of formulae with

**Input:** A CTL formula  $\varphi$  and a set  $V$  of atoms  
**Output:**  $ERes(\varphi, V)$

```

1  $T_\varphi \leftarrow \emptyset$  // the initial set of  $SNF_{CTL}^g$  clauses of  $\varphi$ ;
2  $V' \leftarrow \emptyset$  // the set of atoms introduced in Transform and
  Resolution processes;
3  $T_\varphi, V' \leftarrow Transform(\varphi)$ ;
4  $Res \leftarrow Resolution(T_\varphi, V \cup V')$ ;
5  $Inst_{V'} \leftarrow Instantiate(Res, V')$ ;
6  $Com_{EF} \leftarrow Connect(Inst_{V'})$ ;
7  $RemA \leftarrow Removing\_atoms(Com_{EF}, Inst_{V'})$ ;
8  $NI \leftarrow Removing\_index(RemA)$ ;
9  $Rp \leftarrow Replacing\_atoms(NI)$ ;
10 return  $\bigwedge_{\psi \in Rp_{CTL}} \psi$ .
```

**Algorithm 1:** Computing forgetting - A resolution-based method

$T_0 = \{AG(\mathbf{start} \supset p), AG(p \supset \mathbf{simp}(\mathbf{nnf}(\varphi)))\}$  such that for every  $i$  ( $0 \leq i < n$ ),  $T_{i+1} = (T_i \setminus \{\psi\}) \cup R_i$  (Zhang, Hustadt, and Dixon 2009)), where  $p$  is a new atom not appearing in  $\varphi$ ,  $\psi$  is a formula in  $T_i$  not in  $SNF_{CTL}^g$  clause and  $R_i$  is the result set of applying a matching transformation rule to  $\psi$ . Note that throughout the transformation formulae are kept in negation normal form.

**Proposition 3.** Let  $\varphi$  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  $T_i$ . As an example, we prove  $\varphi \equiv_{\langle \{p\}, \emptyset \rangle} T_0$ .

For one thing,  $\forall (\mathcal{M}_1, s_1) \in Mod(\varphi)$ , i.e.  $(\mathcal{M}_1, s_1) \models \varphi$ . We can construct an initial Ind-model structure  $\mathcal{M}_2$  is identical to  $\mathcal{M}_1$  except  $L_2(s_2) = L_1(s_1) \cup \{p\}$ . It is apparent that  $(\mathcal{M}_2, s_2) \models T_0$  and  $(\mathcal{M}_1, s_1) \leftrightarrow_{\langle \{p\}, \emptyset \rangle} (\mathcal{M}_2, s_2)$ .

For another,  $\forall (\mathcal{M}_1, s_1) \in Mod(T_0)$ , it is apparent that  $(\mathcal{M}_1, s_1) \models \varphi$  by the semantic of **start**.  $\square$

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

**Example 2.** By the Transform process, the result  $T_\varphi$  of the Example 1 can be listed as follows:

- |  |                                 |  |
|--|---------------------------------|--|
| 1. $\mathbf{start} \supset z$          | 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. $z \supset AFx$                     | 8. $y \supset AX(x \vee y)$     |  |

Besides, the set of new atoms introduced in this process is  $V' = \{x, y, x\}$ .

## 4.2 The Resolution process

The *Resolution* process is to compute all the possible resolutions of  $T_\varphi$  on  $V \cup V'$ , denoted as  $Resolution(T_\varphi, V \cup V')$ . A derivation on a set  $V \cup V'$  of atoms and  $T_\varphi$  is a sequence  $T_0 = T_\varphi, T_1, T_2, \dots, T_n = Res$  of sets of  $SNF_{CTL}^g$  clauses such that  $T_{i+1} = T_i \cup R_i$  for all  $0 \leq i < n$ , 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

**Input:** A CTL formula  $\varphi$   
**Output:** A set  $T_\varphi$  of  $SNF_{CTL}^g$  clauses and a set  $V'$  of atoms

```

1  $T_\varphi \leftarrow \emptyset$  // the initial set of  $SNF_{CTL}^g$  clauses of  $\varphi$ ;
2  $OldT \leftarrow \{\mathbf{start} \supset z, z \supset \mathbf{simp}(\mathbf{nnf}(\varphi))\}$ ;
3  $V' \leftarrow \{z\}$ ;
4 while  $OldT \neq T_\varphi$  do
5    $OldT \leftarrow T_\varphi$ ;
6    $R \leftarrow \emptyset$ ;
7    $X \leftarrow \emptyset$ ;
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     $V' \leftarrow V' \cup X$ ;
12     $T_\varphi \leftarrow OldT \setminus \{\psi\} \cup R$ ;
13  end
14 end
```

**Algorithm 2:**  $Transform(\varphi)$

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  $res(C, C')$  is a set of  $SNF_{CTL}^g$  clauses, if there is a resolution rule using  $C$  and  $C'$  as the premises on some given atom. And the pseudocode of *Resolution* process is as Algorithm 3.

**Proposition 4.** Let  $\varphi$  be a CTL formula, then  $T_\varphi \equiv_{\langle V \cup V', \emptyset \rangle} Res$ .

*Proof.* (sketch) This can be proved from  $T_i$  to  $T_{i+1}$  ( $0 \leq i < n$ ) by using one resolution rule on  $T_i$ . For instance, if we can use resolution rule (SRES1) on  $\psi \subseteq T_i$  and obtain the result  $R$ , then we can prove  $T_i \equiv T_{i+1}$  with  $T_{i+1} = T_i \cup R$  as follows.

On one hand, it is apparent that  $\psi \models R$  and then  $T_i \models T_{i+1}$ . On the other hand,  $T_i \subseteq T_{i+1}$  and then  $T_{i+1} \models T_i$ .  $\square$

Proposition 3 and Proposition 4 mean that  $\varphi \equiv_{\langle V \cup V', I \rangle} Res$ , this resolve part of the problem (1).

**Example 3.** The resolutions of  $T_\varphi$  obtained from Example 2 on  $V \cup V'$  are listed as follows:

- |  |               |
|--|---------------|
| (1) $\mathbf{start} \supset r$                 | (1, 2, SRES5) |
| (2) $\mathbf{start} \supset x \vee y$          | (1, 4, SRES5) |
| (3) $\top \supset \neg z \vee y \vee f \vee m$ | (3, 4, SRES8) |
| (4) $y \supset AX(f \vee m \vee y)$            | (3, 8, SRES6) |
| (5) $\top \supset \neg z \vee x \vee p$        | (4, 5, SRES8) |

**Input:** A set  $T_\varphi$  of  $\text{SNF}_{\text{CTL}}^g$  clauses and a set  $V \cup V'$  of atoms

**Output:** A set  $\text{Res}$  of  $\text{SNF}_{\text{CTL}}^g$  clauses

```

1  $S \leftarrow \{C \mid C \in T_\varphi \text{ and } \text{Var}(C) \cap (V \cup V') = \emptyset\};$ 
2  $\Pi \leftarrow T \setminus S;$ 
3 for ( $p \in V \cup V'$ ) do
4    $\Pi' \leftarrow \{C \in \Pi \mid p \in \text{Var}(C)\};$ 
5    $\Sigma \leftarrow \Pi \setminus \Pi';$ 
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) do
8        $\Sigma \leftarrow \Sigma \cup \text{res}(C, C');$ 
9        $\Pi' \leftarrow \Pi' \cup \{C'' \in \text{res}(C, C') \mid p \in \text{Var}(C'')\};$ 
10    end
11  end
12   $\Pi \leftarrow \Sigma;$ 
13 end
14  $\text{Res} \leftarrow \Pi \cup S;$ 

```

**Algorithm 3:**  $\text{Resolution}(T, V \cup V')$

- |   |                           |
|---|---------------------------|
| (6) $\top \supset \neg z \vee x \vee q$         | (4, 6, $\text{SRES8}$ )   |
| (7) $y \supset \text{AX}(x \vee p)$             | (5, 7, $\text{SRES6}$ )   |
| (8) $y \supset \text{AX}(x \vee q)$             | (5, 8, $\text{SRES6}$ )   |
| (9) <b>start</b> $\supset f \vee m \vee y$      | (3, (2), $\text{SRES5}$ ) |
| (10) <b>start</b> $\supset x \vee p$            | (5, (2), $\text{SRES5}$ ) |
| (11) <b>start</b> $\supset x \vee q$            | (6, (2), $\text{SRES5}$ ) |
| (12) $\top \supset p \vee \neg z \vee f \vee m$ | (5, (3), $\text{SRES8}$ ) |
| (13) $\top \supset q \vee \neg z \vee f \vee m$ | (6, (3), $\text{SRES8}$ ) |
| (14) $y \supset \text{AX}(p \vee f \vee m)$     | (5, (4), $\text{SRES6}$ ) |
| (15) $y \supset \text{AX}(q \vee f \vee m)$     | (6, (4), $\text{SRES6}$ ) |
| (16) <b>start</b> $\supset f \vee m \vee p$     | (5, (9), $\text{SRES5}$ ) |
| (17) <b>start</b> $\supset f \vee m \vee q$     | (6, (9), $\text{SRES5}$ ) |

### 4.3 The Elimination process

For resolving problem (2), we should pay attention to the following properties that obtained from the transformation and resolution rules at first:

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

This *Elimination* process include three sub-processes: *Instantiate*, *Connect* and *Removing atoms*. We will describe those sub-processes carefully blow.

**The Instantiation process** An *instantiate formula*  $\psi$  of set  $V''$  of atoms is a formula such that  $\text{Var}(\psi) \cap V'' = \emptyset$ . Given a formula of the form  $p \supset \psi$  with  $p$  is an atom not in  $V'' \cup \text{Var}(\psi)$ , if  $\psi$  is an instantiate formula of set  $V''$  then we call  $p$  is instantiated by  $\psi$ . A key point to compute forgetting is eliminateing those irrelevant atoms, for this purpose we define the follow instantiation process.

**Definition 4** (instantiation). Let  $V'' = V'$  and  $\Gamma = \text{Res}$ , then the process of instantiation is as follows:

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

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

Intuitively, this process iteratively removes the atoms in  $V'$  that can be represented by the formula of  $\text{Var}(\varphi) \setminus (V'' \cup V)$ . We denote this process as  $\text{Instantiate}(\Gamma, V')$ . After this process we obtain a set of atoms that do not has been instantiated by any instantiate formula of  $V \cup V''$  in this process.

**Example 4.** By using the instantiation process on result of Example 3, we obtain that  $x$  is instantiated by  $f \vee m$  at first since there is  $\top \supset \neg x \vee f \vee m \in T_\varphi$  with  $x \in V'$  and  $\text{Var}(f \vee m) \cap (V \cup V') = \emptyset$ , then  $V'' = \{y, z\}$ .

Similarly, due to  $\top \supset \neg y \vee q \in T_\varphi$  and  $y \supset \text{AX}(q \vee f \vee m) \in T_\varphi$ , then  $y$  can be instantiated by  $q \wedge \text{AX}(q \vee f \vee m)$ . And  $z$  can be instantiated by  $r$ . Therefore  $V'' = \emptyset$  That is  $\text{Instantiate}(\text{Res}, V') = V$ , which means all the introduced atoms are instantiated.

By instantiation operator, we guarantee those atoms in  $V \cup V''$  are really irrelevant, i.e. should be forgot.

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

- (EF1)  $\{P \supset \text{Afl}, P \supset \text{E}_{\langle \text{ind} \rangle} \text{X}(l_1 \vee C_4)\} \cup A \rightarrow \alpha$
- (EF2)  $\{P \supset \text{Afl}, P \supset \text{AX}(l_1 \vee C_4)\} \cup A \rightarrow \beta$
- (EF3)  $\{P \supset \text{E}_{\langle \text{ind} \rangle} \text{Fl}, P \supset \text{E}_{\langle \text{ind} \rangle} \text{X}(l_1 \vee C_4)\} \cup A \rightarrow \gamma$
- (EF4)  $\{P \supset \text{E}_{\langle \text{ind} \rangle} \text{Fl}, P \supset \text{AX}(l_1 \vee C_4)\} \cup A \rightarrow \gamma$ .

By  $\text{Connect}(\text{Instantiate}(\text{Res}, V'))$  we mean using (EF1) to (EF4) on  $\text{Res}$  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 rule (EF1), 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  $P \supset \text{AX}(\neg l \vee C_2 \vee C_4)$  with  $P \supset \text{AX}(\neg l \vee C_2 \vee C_4) \vee \gamma$  for other rules when  $l, C_2, C_3$  and  $C_4$  are instantiate formulae of  $\text{Sub}(\text{Res}, V')$  and  $\text{Var}(l_1) \in$

$V \cup V'$ . The reason why we specify  $l$ ,  $C_2$ ,  $C_3$  and  $C_4$  are instantiate formulae of  $\text{Sub}(\text{Res}, V')$  in this process will be explained later.

**Proposition 5.** Let  $\Gamma = \text{Res}$ , we have  $\Gamma \equiv_{\langle V', \emptyset \rangle} \text{Connect}(\text{Instantiate}(\Gamma, V'))$ .

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

We prove the (EF1), 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.  $\varphi = P \supset ((\neg C_3 \wedge \neg C_2) \supset (\text{E}_{\langle \text{ind} \rangle} X(C_3 \wedge \neg(C_2 \vee C_4) \supset \text{AXAF}(C_3 \vee C_2))))$ . It is apparent that  $T_{i+1} \models T_i$  and  $T_i \models P \supset \text{E}_{\langle \text{ind} \rangle} X(\neg l \vee C_2 \vee C_4)$ . 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 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  $(\mathcal{M}, s) \models P \supset \text{E}_{\langle \text{ind} \rangle} X(l_1 \vee C_4)$ ). If  $(\mathcal{M}, s_1) \not\models \text{AXAF}(C_3 \vee C_2)$  then we have  $(\mathcal{M}, s_1) \models l$  due to  $(\mathcal{M}, s) \models \text{AG}(l \supset C_3 \vee C_2)$  and  $(\mathcal{M}, s) \models \text{AF}l$ . And then  $(\mathcal{M}, s_1) \models \neg l_1$  by  $(\mathcal{M}, s) \models \text{AG}(l \supset \neg l_1 \vee C_2)$ . It is contract with our supposing. Then  $(\mathcal{M}, s_1) \models \text{AXAF}(C_3 \vee C_2)$ .  $\square$

**The Removing\_atoms process** For eliminating those irrelevant atoms, we define the following *Removing\_atoms* operator.

**Definition 5** (Removing\_atoms). Let  $T$  be a set of formulae,  $C \in T$  and  $V$  a set of atoms, then the *Removing\_atoms* operator is defined as:

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

Which means that if the formula  $C$  containing at least one of atoms in  $V$  then let  $\text{Removing\_atoms}(C, V)$  be true, else be  $C$  itself. For convenience, for any set  $T$  of formula we have  $\text{Removing\_atoms}(T, V) = \{\text{Removing\_atoms}(r, V) \mid r \in T\}$ .

**Proposition 6.** Let  $V'' = V \cup V'$ ,  $\Gamma = \text{Instantiate}(\text{Res}, V')$  and  $\Gamma_1 = \text{Removing\_atoms}(\text{Connect}(\Gamma), \Gamma)$ , then  $\Gamma_1 \equiv_{\langle V'', \emptyset \rangle} \text{Res}$  and  $\Gamma_1 \equiv_{\langle V'', I \rangle} \varphi$ .

*Proof.* (stretch) Take note the fact that for each clause  $C = T \supset H$  in  $\text{Connect}(\Gamma)$ , 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  $\text{Connect}(\Gamma) \models \Gamma_1$ , we will show  $\forall (\mathcal{M}, s_0) \in \text{Mod}(\Gamma_1)$  there is a  $(\mathcal{M}', s_0)$  such that  $(\mathcal{M}', s_0) \models \text{Connect}(\Gamma)$  and  $(\mathcal{M}, s_0) \leftrightarrow_{\langle \Gamma, \emptyset \rangle} (\mathcal{M}', s_0)$ . Let  $C = T \supset H$  in  $\text{Connect}(\Gamma)$  with  $\Gamma \cap \text{Var}(C) \neq \emptyset$ ,  $\forall (\mathcal{M}, s_0) \in \text{Mod}(\Gamma_1)$  we construct  $(\mathcal{M}', s_0)$  s.t.  $(\mathcal{M}, s_0) \leftrightarrow_{\Gamma} (\mathcal{M}', s_0)$  and  $(\mathcal{M}', s_0) \models C$  by adding or deleting some atoms in  $\Gamma$  to the  $L'(s')$  with  $s' \in S'$ .  $\square$

**Example 5.** After removing the clauses that include atoms in  $V = \{p\}$ , the following clauses have been left:

$\text{start} \supset z$	$\top \supset \neg z \vee r$
$\top \supset \neg x \vee f \vee m$	$\top \supset \neg z \vee x \vee y$
$\top \supset \neg y \vee p$	$\top \supset \neg y \vee q$
$z \supset \text{AF}x$	$y \supset \text{AX}(x \vee y)$
$\text{start} \supset r$	$\text{start} \supset x \vee y$
$\top \supset \neg z \vee y \vee f \vee m$	$y \supset \text{AX}(f \vee m \vee y)$
$\top \supset \neg z \vee x \vee q$	$y \supset \text{AX}(x \vee q)$
$\text{start} \supset f \vee m \vee y$	$\text{start} \supset x \vee q$
$\top \supset q \vee \neg z \vee f \vee m$	$y \supset \text{AX}(q \vee f \vee m)$
$\text{start} \supset f \vee m \vee q$	

In this case, if we do not specify  $l$ ,  $C_2$ ,  $C_3$  and  $C_4$  are instantiate formulae of  $\text{Sub}(\text{Res}, V')$ , it is easy to check that all results including  $P \supset \text{E}_{\langle \text{ind} \rangle} X(\neg l \vee C_2 \vee C_4)$  and  $P \supset \text{AX}(\neg l \vee C_2 \vee C_4)$  obtained from the *Connect* process will be deleted in the *Removing\_atoms* process.

#### 4.4 Remove the Index and start

The *Removing\_index*(*RemA*) process is to change the set *RemA* obtained above into a set of formulas without the index by using the equations in Proposition 7.

**Proposition 7.** Let  $P$ ,  $P_i$  and  $\varphi_i$  be CTL formulae, then

- (i)  $\bigwedge_{i=1}^n (P \supset \text{E}_{\langle \text{ind} \rangle} X \varphi_i) \equiv_{\langle \emptyset, \{\text{ind}\} \rangle} P \supset \text{EX} \bigwedge_{i=1}^n \varphi_i$ ,
- (ii)  $\bigwedge_{i=1}^n (P_i \supset \text{E}_{\langle \text{ind} \rangle} X \varphi_i) \equiv_{\langle \emptyset, \{\text{ind}\} \rangle} \bigwedge_{e \in 2^{\{0, \dots, n\}} \setminus \{\emptyset\}} (\bigwedge_{i \in e} P_i \supset \text{EX}(\bigwedge_{i \in e} \varphi_i))$ ,
- (iii)  $\bigwedge_{i=1}^n (P \supset \text{E}_{\langle \text{ind} \rangle} F \varphi_i) \equiv_{\langle \emptyset, \{\text{ind}\} \rangle} P \supset \bigvee \text{EF}(\varphi_{j_1} \wedge \text{EF}(\varphi_{j_2} \wedge \text{EF}(\dots \wedge \text{EF} \varphi_{j_n})))$ , where  $(j_1, \dots, j_n)$  are sequences of all elements in  $\{0, \dots, n\}$ ,
- (iv)  $P \supset (C \vee \text{E}_{\langle \text{ind} \rangle} X \varphi_1) \wedge P \supset \text{E}_{\langle \text{ind} \rangle} X \varphi_2 \equiv_{\langle \emptyset, \{\text{ind}\} \rangle} P \supset ((C \wedge \text{EX} \varphi_2) \vee \text{EX}(\varphi_1 \wedge \varphi_2))$ ,
- (v)  $P \supset (C \vee \text{E}_{\langle \text{ind} \rangle} X \varphi_1) \vee P \supset \text{E}_{\langle \text{ind} \rangle} X \varphi_2 \equiv_{\langle \emptyset, \{\text{ind}\} \rangle} P \supset (C \vee \text{EX}(\varphi_1 \vee \varphi_2))$ .

*Proof.* (i)  $\forall (\mathcal{M}, s_0) \in \text{Mod}(\bigwedge_{i=1}^n (P \supset \text{E}_{\langle \text{ind} \rangle} X \varphi_i))$  there is  $(s_0, s_1) \in [\text{ind}]$  such that  $(\mathcal{M}, s_1) \models \varphi_1, \dots, (\mathcal{M}, s_1) \models \varphi_n$ , then there is  $(s_0, s'_1) \in R$  s.t.  $(\mathcal{M}, s'_1) \models \bigwedge_{i=1}^n \varphi_i$ , i.e.  $(\mathcal{M}, s_0) \models P \supset \text{EX} \bigwedge_{i=1}^n \varphi_i$ .

For each  $(\mathcal{M}, s_0) \in \text{Mod}(P \supset \text{EX} \bigwedge_{i=1}^n \varphi_i)$ , we suppose there is  $(s_0, s_1) \in R$  s.t.  $(\mathcal{M}, s_1) \models \bigwedge_{i=1}^n \varphi_i$ . It is easy to construct an initial Ind-model  $(\mathcal{M}', s_0)$  such that  $(\mathcal{M}', s_0)$  is identical to  $(\mathcal{M}, s_0)$  except the  $(s_0, s_1) \in [\text{ind}]$ , i.e.  $(\mathcal{M}, s_0) \leftrightarrow_{\langle \emptyset, \{\text{ind}\} \rangle} (\mathcal{M}', s_0)$ .

(ii) Intuitively, from the left to the right: for any model  $(\mathcal{M}, s_0)$  of the left side of the equation if there is  $(\mathcal{M}, s_0) \models \bigwedge_{i=1}^m P_{j_i}$  with  $j_i \in \{1, \dots, n\}$  and  $1 \leq m \leq n$ , then there is some next state  $s_1$  of  $s_0$  with  $(s_0, s_1) \in [\text{ind}]$  such that  $(\mathcal{M}, s_1) \models \bigwedge_{i=1}^m \varphi_{j_i}$ . By the definition of  $[\text{ind}]$ , we have  $(s_0, s_1) \in R$  and then  $(\mathcal{M}, s_0) \models \bigwedge_{i=1}^m P_{j_i} \supset \text{EX}(\bigwedge_{i=1}^m \varphi_{j_i})$ . The other side can be similarly proved as (i).

(iii) From the right to the left: for any model  $(\mathcal{M}, s_0)$  of the right side of the equation if there is  $(\mathcal{M}, s_0) \models P$  then

there is a path  $\pi_{s_0}$  such that  $\varphi_i \in \pi_{s_0}$  ( $1 \leq i \leq n$ ). Then we can construct an initial Ind-model  $(\mathcal{M}', s_0)$  such that  $(\mathcal{M}', s_0)$  is identical to  $(\mathcal{M}, s_0)$  except for each  $(s_i, s_{i+1})$  of  $\pi_{s_0}$  there is  $(s_i, s_{i+1}) \in [ind]$ . It is easy to check  $(\mathcal{M}', s_0) \models \bigwedge_{i=1}^n (P \supset E_{\langle ind \rangle} \varphi_i)$  and  $(\mathcal{M}, s_0) \leftrightarrow_{\langle \emptyset, \{ind\} \rangle} (\mathcal{M}', s_0)$ . The other side can be similarly proved as (ii).

Other results can be proved similarly.  $\square$

**Proposition 8. (NI-BRemain)** *Let  $I$  be the set of indexes appearing in RemA, we have  $RemA \equiv_{\langle \emptyset, I \rangle} Removing\_index(RemA)$ .*

In our Example 5 we do not need this process since there is no index in the set of formulae. Let  $T$  be a set of  $SNF_{CTL}^g$  clauses, then we define the following operator:

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

Then  $T \equiv T_{CTL}$  by  $\varphi \equiv AG(\mathbf{start} \supset \varphi)$  (Bolotov 2000).

The last step of our algorithm is to eliminate all the atoms in  $V'$  which has been introduced in the Transform process. Let  $\Gamma = Instantiate(Res, V')$  and  $\Gamma_1 = Removing\_atoms(Connect(\Gamma))$ , then  $Replacing\_atoms(Removing\_index(\Gamma_1))$  is obtained from  $Removing\_index(\Gamma_1)$  by doing the following three steps for each  $p \in (V' \setminus \Gamma)$ :

- replacing each  $p \supset \varphi_1 \vee \dots \vee p \supset \varphi_n$  with  $p \supset \bigvee_{i=1}^n \varphi_i$ ;
- replacing  $p \supset \varphi_1 \wedge \dots \wedge p \supset \varphi_m$  with  $\varphi_j$  are instantiate formulae of  $\Gamma$  ( $j \in \{1, \dots, m\}$ ) with  $p \supset \psi$ , where  $\psi = \bigwedge_{j=1}^m \varphi_j$  and  $p$  do not appear in  $\varphi_j$ .
- For any formula  $C \in \Gamma_1$ , replacing every  $p$  in  $C$  with  $\psi$ .

Recall that any atom in  $V'$  introduced in the Transform process is a name of the sub-formula of  $\varphi$  (Bolotov 2000). Apparently, this process is just a process of replacing each atom with an equivalent formula. Then we have:

**Proposition 9.** *Let  $\Gamma_1 = Instantiate(Res, V')$ ,  $\Gamma_2 = Removing\_atoms(Connect(\Gamma_1, \Gamma_1))$  and  $\Gamma_3 = Replacing\_atoms(Removing\_index(\Gamma_2))$ , then  $\Gamma_2 \equiv_{\langle V' \setminus \Gamma_1, I \rangle} \Gamma_3$  and  $\varphi \equiv_{\langle V \cup V', \emptyset \rangle} (\Gamma_3)_{CTL}$ .*

**Example 6.** By using the Replacing\_atoms process on result of Example 5 directly since there is not index in those clauses, we obtain that  $x$  is replaced by  $f \vee m$  at first, then  $y$  is replaced by  $q \wedge AX(q \vee f \vee m)$  and  $z$  is replaced by  $r \wedge (f \vee m \vee q) \wedge (f \vee m \vee (q \wedge AX(f \vee m \vee q))) \wedge AF(f \vee m)$ .

#### 4.5 An example for Connect process

In order to show the necessity of the Connect process, we give the following example at first.

**Example 7.** Let  $\psi = AF(p \wedge q) \wedge EX\neg p$  and  $V = \{p\}$ . By the processes Transform and Resolution, we can obtain  $V' = \{f, z\}$  and the following set Res of  $SNF_{CTL}^g$  clauses.

$$\begin{array}{lll} \mathbf{start} \supset z & z \supset AFf & z \supset E_{\langle ind \rangle} X\neg p \\ \top \supset \neg f \vee p & \top \supset \neg f \vee q & z \supset E_{\langle ind \rangle} X\neg f \end{array}$$

According to our Algorithm 1, we have  $Instantiate(Res, V') = V$  since  $f$  can be instantiated by  $q$  and  $z$  can be instantiated by  $AFf$ .

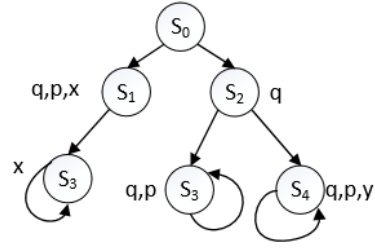


Figure 1: A model  $(\mathcal{M}, s_0)$  of  $\varphi$

On the one hand, in the Connect process, by using (EF1) rule on the Res we have  $\alpha = z \supset (\neg q \supset (E_{\langle ind \rangle} X(q \supset AXAFq)))$  and replace  $z \supset E_{\langle ind \rangle} X\neg f \in Res$  with  $z \supset E_{\langle ind \rangle} X\neg f \vee \alpha$  since  $l, C_2, C_3$  and  $C_4$  are instantiate formulae. Apparently,  $z \supset E_{\langle ind \rangle} X\neg f \vee \alpha \equiv z \supset q \vee E_{\langle ind \rangle} X(\neg f \vee \neg q \vee AXAFq)$ .

After the Removing\_atoms process, we have the following set RemA of formulae:

$$\begin{array}{ll} \mathbf{start} \supset z & z \supset AFf \\ \top \supset \neg f \vee q & z \supset q \vee E_{\langle ind \rangle} X(\neg f \vee \neg q \vee AXAFq) \end{array}$$

Removing the indexes appearing in the RemA, we obtain the following set NI:

$$\begin{array}{ll} \mathbf{start} \supset z & z \supset AFf \\ \top \supset \neg f \vee q & z \supset q \vee EX(\neg f \vee \neg q \vee AXAFq) \end{array}$$

Replacing the atoms in  $V'$  that have been instantiated, i.e.  $f$  is replaced with  $q$  and  $z$  is replaced with  $AFq \wedge (q \vee EX(\neg q \vee AXAFq))$ , we have

$$Rp = \{\mathbf{start} \supset AFq \wedge (q \vee EX(\neg q \vee AXAFq))\}.$$

As all the formulas  $\mathcal{F}$  in the  $T_\varphi$  are the form  $AG\mathcal{F}$ , hence we have:

$$Rp_{CTL} = \{AFq \wedge (q \vee EX(\neg q \vee AXAFq))\}.$$

i.e.  $ERes(\varphi, V) = AFq \wedge (q \vee EX(\neg q \vee AXAFq))$ . In this case, we can easily check that  $ERes(\varphi, V) \equiv_{\langle V, \emptyset \rangle} \varphi$ .

On the other hand, if we do not using the Connect process, we can easily obtain the result of ERes, i.e.  $ERes(\varphi, V) = AFq \wedge EX(\neg q)$ . It is apparent that  $ERes(\varphi, V) \not\equiv_{\langle V, \emptyset \rangle} \varphi$ . This can be proved by model  $(\mathcal{M}, s_0)$  as in Figure 1 since  $(\mathcal{M}, s_0) \models \varphi$  and  $(\mathcal{M}, s_0) \not\models ERes(\varphi, V)$ .

This example shows that why we introduce the EF-implication rules. Intuitively, the result of replacing the atoms that have been instantiated in  $V'$  with an instantiate formula is more stronger than our method, because by the Removing\_atoms process, we have removing some clauses, such as  $C = \top \supset \neg f \vee p$ , that contain  $f$ . The original one is  $f \supset p \wedge q$ , but after removing  $C$  we only obtain that  $f \supset q$ . In this example, there is a clause  $z \supset EX\neg f \in Res$ , after replacing  $f$  with  $q$ , we obtain  $z \supset EX\neg q$ . However, if we do not removing  $C$  (i.e.  $f \supset p \wedge q$ ), then we have  $z \supset EX(\neg q \vee \neg p)$ , this is weaker than  $z \supset EX\neg q$ . In fact, for any model  $(\mathcal{M}, s_0)$  of  $\varphi$  there is not necessary  $q \notin L(s)$



for some next state  $s$  of  $s_0$  and if there is  $q \in L(s)$  for all next states  $s$ , then there must be a next state  $s$  of  $s_0$  with  $p \notin L(s)$  s.t. for all next state  $s'$  of  $s$  there is  $(\mathcal{M}, s') \models \text{AF}q$  (see Fig. 1). This is what the meaning of the *Connect* process.

#### 4.6 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.

The correction means that the result  $\text{ERes}(\varphi, V)$  obtained from our Algorithm is  $\text{F}_{\text{CTL}}(\varphi, V)$ , i.e. input  $\varphi$  and  $V$  to Algorithm 1 output the result of forgetting  $V$  from  $\varphi$ .

**Theorem 2** (Resolution-based CTL-forgetting). *Let  $V'' = V \cup V'$  and  $\Gamma_1 = \text{ERes}(\varphi, V)$ , then*

- (i)  $\text{F}_{\text{CTL}}(\varphi, V'') \equiv \Gamma_1$ ;
- (ii)  $\text{F}_{\text{CTL}}(\varphi, V) \equiv \Gamma_1$ .

*Proof.* (i)  $(\Rightarrow) \forall (\mathcal{M}, s_0) \in \text{Mod}(\text{F}_{\text{CTL}}(\varphi, V''))$  there is  $(\mathcal{M}', s'_0) \in \text{Mod}(\varphi)$  s.t.  $(\mathcal{M}, s_0) \leftrightarrow_{V''} (\mathcal{M}', s'_0)$  by the definition of  $\text{F}_{\text{CTL}}$ , and then  $\exists (\mathcal{M}_1, s_1) \in \text{Mod}(\Gamma_1)$  s.t.  $(\mathcal{M}_1, s_1) \leftrightarrow_{V''} (\mathcal{M}', s'_0)$  by Proposition 9. Hence,  $(\mathcal{M}, s_0) \leftrightarrow_{V''} (\mathcal{M}_1, s_1)$  due to  $\leftrightarrow$  is an equivalence relation. Therefore,  $(\mathcal{M}, s_0) \models \Gamma_1$  due to  $\text{IR}(\Gamma_1, V'')$  and Theorem 1.

$(\Leftarrow) \forall (\mathcal{M}_1, s_1) \in \text{Mod}(\Gamma_1)$  there is  $(\mathcal{M}', s'_0) \in \text{Mod}(\varphi)$  s.t.  $(\mathcal{M}_1, s_1) \leftrightarrow_{V''} (\mathcal{M}', s'_0)$  by Proposition 9. Hence,  $(\mathcal{M}_1, s_1) \models \text{F}_{\text{CTL}}(\varphi, V'')$  due to  $\text{IR}(\text{F}_{\text{CTL}}(\varphi, V''), V'')$  and  $\varphi \models \text{F}_{\text{CTL}}(\varphi, V'')$ .

(ii) It is obtained from (i) since  $\text{IR}(\varphi, V')$ .  $\square$

Then we can obtain the result of forgetting of Example 4:

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

**Proposition 10.** *Let  $\varphi$  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 indices introduced during transformation.*

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

### 5 Related work

#### 5.1 Resolution-based satisfiability of CTL

Deciding the satisfiability with resolution calculus in Propositional Linear Temporal Logic (PLTL) was firstly introduced in (Fisher 1991) and further discussed in (Fisher 1997; Fisher, Dixon, and Peim 2001). The main idea is that transforming any PLTL formula into the normal form, called Separated Normal Form (SNF) by introducing a new connective **start** that holds only at the beginning of time.

After that the Resolution-based satisfiability in CTL was proposed by Bolotov in (Bolotov 2000) at first and then be refined by Zhang in (Zhang, Hustadt, and Dixon 2009; Zhang, Hustadt, and Dixon 2014). In those papers, the main idea is also to transform any CTL formula into the normal form  $\text{SNF}_{\text{CTL}}^g$ . But the CTL is a kind of branch time temporal logic, they introduced the “index” besides **start** for that purpose.

All in all, a complete set of transformation and resolution rules had been proposed for both PLTL and CTL. And it shows that the transformation is satisfiability preserving and also for the result obtained from using the resolution rules on the normal form.

#### 5.2 Using Resolution Computing forgetting

Resolution, a kind of methods of Second-order quantifier elimination, has been used to compute the forgetting or uniform interpretation in propositional logic (Wang 2015) and Modal logic (Herzig and Mengin 2008). In those case, the formula is required to be a paradigm with a particular form-“CNF” (the definition of CNF in Modal logic can be found in (Herzig and Mengin 2008)).

As have said above that the normal form used to resolution is an extension of CTL by introducing the **start** and “index”. In this article, we propose the  $\langle V, I \rangle$ -bisimulation to solve the “index” problem. Besides, in order to eliminate those atoms introduced in the transformation, we proposed the four EF-implication rules.

### 6 Conclusion and Future Work

This article proposed a resolution-based algorithm to compute the forgetting in CTL. Our method extend the resolution calculus in (Zhang, Hustadt, and Dixon 2014) by adding processes including removing those irrelevant atoms and transforming the result into CTL formula. For this purpose, a kind of binary bisimulation relation, called  $\langle V, I \rangle$ -bisimulation, has been defined to bridge the gap between CTL and  $\text{SNF}_{\text{CTL}}^g$ . Besides, for connect the next state and future state we propose four EF-implication rules, this weaken the result of the original resolution and our Replacing-atoms processes to obtain the correct result. More important, our algorithm is correct, i.e. return the result of forgetting some set of atoms. Besides, examples show how to compute forgetting using our algorithm.

#### A Supplementary Material: Proof Appendix

**Lemma 2.** *Let  $\mathcal{B}_0, \mathcal{B}_1, \dots$  be the ones in the definition of section 3.1. Then, for each  $i \geq 0$ ,*

- (i)  $\mathcal{B}_{i+1} \subseteq \mathcal{B}_i$ ;
- (ii) there is a (smallest)  $k \geq 0$  such that  $\mathcal{B}_{k+1} = \mathcal{B}_k$ ;
- (iii)  $\mathcal{B}_i$  is reflexive, symmetric and transitive.

*Proof.* (i) Base: it is clear for  $i = 0$  by the above definition.

Step: suppose it holds for  $i = n$ , i.e.,  $\mathcal{B}_{n+1} \subseteq \mathcal{B}_n$ .

$(s, s') \in \mathcal{B}_{n+2}$

$\Rightarrow$  (a)  $(s, s') \in \mathcal{B}_0$ , (b) for every  $(s, s_1) \in R$ , there is  $(s', s'_1) \in R'$  such that  $(s_1, s'_1) \in \mathcal{B}_{n+1}$ , and (c) for every  $(s', s'_1) \in R'$ , there is  $(s, s_1) \in R$  such that  $(s_1, s'_1) \in \mathcal{B}_{n+1}$

$\Rightarrow$  (a)  $(s, s') \in \mathcal{B}_0$ , (b) for every  $(s, s_1) \in R$ , there is  $(s', s'_1) \in R'$  such that  $(s_1, s'_1) \in \mathcal{B}_n$  by inductive assumption, and (c) for every  $(s', s'_1) \in R'$ , there is  $(s, s_1) \in R$  such that  $(s_1, s'_1) \in \mathcal{B}_n$  by inductive assumption  
 $\Rightarrow (s, s') \in \mathcal{B}_{n+1}$ .

(ii) and (iii) are evident from (i) and the definition of  $\mathcal{B}_i$ .  $\square$

**Lemma 1** The relation  $\leftrightarrow_V$  is an equivalence relation.

*Proof.* It is clear from Lemma 2 (ii) such that there is a  $k \geq 0$  where  $\mathcal{B}_k = \mathcal{B}_{k+1}$  which is  $\leftrightarrow_V$ , and it is reflexive, symmetric and transitive by (iii).  $\square$

**Proposition 1** Let  $i \in \{1, 2\}$ ,  $V_1, V_2 \subseteq \mathcal{A}$ , 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)  $\mathcal{K}_1 \leftrightarrow_{V_1 \cup V_2} \mathcal{K}_3$ ;
- (ii) If  $V_1 \subseteq V_2$  then  $\mathcal{K}_1 \leftrightarrow_{V_2} \mathcal{K}_2$ .

*Proof.* (i) Let  $\mathcal{M}_i = (S_i, R_i, L_i, s_i)$  ( $i = 1, 2, 3$ ),  $s_1 \leftrightarrow_{V_1} s_2$  via a binary relation  $\mathcal{B}$ , and  $s_2 \leftrightarrow_{V_2} s_3$  via a binary relation  $\mathcal{B}''$ . Let  $\mathcal{B}' = \{(w_1, w_3) | (w_1, w_2) \in \mathcal{B} \text{ and } (w_2, w_3) \in \mathcal{B}_2\}$ . It's apparent that  $(s_1, s_3) \in \mathcal{B}'$ . We prove  $\mathcal{B}'$  is a  $V_1 \cup V_2$ -bisimulation containing  $(s_1, s_3)$  from the (a), (b) and (c) of the previous step (iii) of  $X$ -bisimulation (where  $X$  is a set of atoms). For all  $(w_1, w_3) \in \mathcal{B}'$ :

- (a) there is  $w_2 \in S_2$  such that  $(w_1, w_2) \in \mathcal{B}$  and  $(w_2, w_3) \in \mathcal{B}''$ , and  $\forall q \notin V_1, q \in L_1(w_1)$  iff  $q \in L_2(w_2)$  by  $w_1 \leftrightarrow_{V_1} w_2$  and  $\forall q' \notin V_2, q' \in L_2(w_2)$  iff  $q' \in L_3(w_3)$  by  $w_2 \leftrightarrow_{V_2} w_3$ . Then we have  $\forall r \notin V_1 \cup V_2, r \in L_1(w_1)$  iff  $r \in L_3(w_3)$ .
- (b) if  $(w_1, u_1) \in \mathcal{R}_1$ , then  $\exists u_2 \in S_2$  such that  $(w_2, u_2) \in \mathcal{R}_2$  and  $(u_1, u_2) \in \mathcal{B}$  (due to  $(w_1, w_2) \in \mathcal{B}$  and  $(w_2, w_3) \in \mathcal{B}''$  by the definition of  $\mathcal{B}'$ ); and then  $\exists u_3 \in S_3$  such that  $(w_3, u_3) \in \mathcal{R}_3$  and  $(u_2, u_3) \in \mathcal{B}''$ , hence  $(u_1, u_3) \in \mathcal{B}'$  by the definition of  $\mathcal{B}'$ .
- (c) if  $(w_3, u_3) \in \mathcal{R}_3$ , then  $\exists u_2 \in S_2$  such that  $(w_2, u_2) \in \mathcal{R}_2$  and  $(u_2, u_3) \in \mathcal{B}_2$ ; and then  $\exists u_1 \in S_1$  such that  $(w_1, u_1) \in \mathcal{R}_1$  and  $(u_1, u_2) \in \mathcal{B}$ , hence  $(u_1, u_3) \in \mathcal{B}'$  by the definition of  $\mathcal{B}'$ .

(ii) Let  $\mathcal{K}_{i,j} = (\mathcal{M}_i, s_{i,j})$  and  $(s_{i,k}, s_{i,k+1}) \in R_i$  mean that  $s_{i,k+1}$  is the  $(k+2)$ -th node in the path  $(s_i, s_{i,1}, s_{i,2}, \dots, s_{i,k+1}, \dots)$  ( $i = 1, 2$ ). We will show that  $(\mathcal{K}_1, \mathcal{K}_2) \in \mathcal{B}_n^{V_2}$  for all  $n \geq 0$  inductively.

Base:  $L_1(s_1) - V_1 = L_2(s_2) - V_1$   
 $\Rightarrow \forall q \in \mathcal{A} - V_1$  there is  $q \in L_1(s_1)$  iff  $q \in L_2(s_2)$   
 $\Rightarrow \forall q \in \mathcal{A} - V_2$  there is  $q \in L_1(s_1)$  iff  $q \in L_2(s_2)$  due to  $V_1 \subseteq V_2$   
 $\Rightarrow L_1(s_1) - V_2 = L_2(s_2) - V_2$ , i.e.,  $(\mathcal{K}_1, \mathcal{K}_2) \in \mathcal{B}_0^{V_2}$ .

Step: Supposing that  $(\mathcal{K}_1, \mathcal{K}_2) \in \mathcal{B}_i^{V_2}$  for all  $0 \leq i \leq k$  ( $k > 0$ ), we will show  $(\mathcal{K}_1, \mathcal{K}_2) \in \mathcal{B}_{k+1}^{V_2}$ .

- (a) It is apparent that  $L_1(s_1) - V_2 = L_2(s_2) - V_2$  by base.

- (b)  $\forall (s_1, s_{1,1}) \in R_1$ , we will show that there is a  $(s_2, s_{2,1}) \in R_2$  s.t.  $(\mathcal{K}_{1,1}, \mathcal{K}_{2,1}) \in \mathcal{B}_k^{V_2}$ .  $(\mathcal{K}_{1,1}, \mathcal{K}_{2,1}) \in \mathcal{B}_{k-1}^{V_2}$  by inductive assumption, we need only to prove the following points:

- (a)  $\forall (s_{1,k}, s_{1,k+1}) \in R_1$  there is a  $(s_{2,k}, s_{2,k+1}) \in R_2$  s.t.  $(\mathcal{K}_{1,k+1}, \mathcal{K}_{2,k+1}) \in \mathcal{B}_0^{V_2}$  due to  $(\mathcal{K}_{1,1}, \mathcal{K}_{2,1}) \in \mathcal{B}_k^{V_1}$ . It is easy to see that  $L_1(s_{1,k+1}) - V_1 = L_1(s_{2,k+1}) - V_1$ , then there is  $L_1(s_{1,k+1}) - V_2 = L_1(s_{2,k+1}) - V_2$ . Therefore,  $(\mathcal{K}_{1,k+1}, \mathcal{K}_{2,k+1}) \in \mathcal{B}_0^{V_2}$ .
- (b)  $\forall (s_{2,k}, s_{2,k+1}) \in R_1$  there is a  $(s_{1,k}, s_{1,k+1}) \in R_1$  s.t.  $(\mathcal{K}_{1,k+1}, \mathcal{K}_{2,k+1}) \in \mathcal{B}_0^{V_2}$  due to  $(\mathcal{K}_{1,1}, \mathcal{K}_{2,1}) \in \mathcal{B}_k^{V_1}$ . This can be proved as (a).

- (c)  $\forall (s_2, s_{2,1}) \in R_1$ , we will show that there is a  $(s_1, s_{1,1}) \in R_2$  s.t.  $(\mathcal{K}_{1,1}, \mathcal{K}_{2,1}) \in \mathcal{B}_k^{V_2}$ . This can be proved as (ii).  $\square$

**Theorem 1** Let  $V \subseteq \mathcal{A}$ ,  $\mathcal{K}_i$  ( $i = 1, 2$ ) be two  $\mathcal{K}$ -structures such that  $\mathcal{K}_1 \leftrightarrow_V \mathcal{K}_2$  and  $\phi$  a formula with  $\text{IR}(\phi, V)$ . Then  $\mathcal{K}_1 \models \phi$  if and only if  $\mathcal{K}_2 \models \phi$ .

*Proof.* This theorem can be proved by inducting on the formula  $\phi$  and supposing  $\text{Var}(\phi) \cap V = \emptyset$ . Let  $\mathcal{K}_1 = (\mathcal{M}, s)$  and  $\mathcal{K}_2 = (\mathcal{M}', s')$ .

**Case**  $\phi = p$  where  $p \in \mathcal{A} - V$ :

$(\mathcal{M}, s) \models \phi$  iff  $p \in L(s)$  (by the definition of satisfiability)  
 $\Leftrightarrow p \in L'(s')$  ( $s \leftrightarrow_V s'$ )

$\Leftrightarrow (\mathcal{M}', s') \models \phi$

**Case**  $\phi = \neg\psi$ :

$(\mathcal{M}, s) \models \phi$  iff  $(\mathcal{M}, s) \not\models \psi$

$\Leftrightarrow (\mathcal{M}', s') \not\models \psi$

(induction hypothesis)

$\Leftrightarrow (\mathcal{M}', s') \models \phi$

**Case**  $\phi = \psi_1 \vee \psi_2$ :

$(\mathcal{M}, s) \models \phi$

$\Leftrightarrow (\mathcal{M}, s) \models \psi_1$  or  $(\mathcal{M}, s) \models \psi_2$

$\Leftrightarrow (\mathcal{M}', s') \models \psi_1$  or  $(\mathcal{M}', s') \models \psi_2$  (induction hypothesis)

$\Leftrightarrow (\mathcal{M}', s') \models \phi$

**Case**  $\phi = \text{EX}\psi$ :

$\mathcal{M}, s \models \phi$

$\Leftrightarrow$  There is a path  $\pi = (s, s_1, \dots)$  such that  $\mathcal{M}, s_1 \models \psi$

$\Leftrightarrow$  There is a path  $\pi' = (s', s'_1, \dots)$  such that  $\pi \leftrightarrow_V \pi'$  ( $s \leftrightarrow_V s'$ , Proposition 1)

$\Leftrightarrow s_1 \leftrightarrow_V s'_1$  ( $\pi \leftrightarrow_V \pi'$ )

$\Leftrightarrow (\mathcal{M}', s'_1) \models \psi$

(induction hypothesis)

$\Leftrightarrow (\mathcal{M}', s') \models \phi$

**Case**  $\phi = \text{EG}\psi$ :

$\mathcal{M}, s \models \phi$

$\Leftrightarrow$  There is a path  $\pi = (s = s_0, s_1, \dots)$  such that for each  $i \geq 0$  there is  $(\mathcal{M}, s_i) \models \psi$

$\Leftrightarrow$  There is a path  $\pi' = (s' = s'_0, s'_1, \dots)$  such that  $\pi \leftrightarrow_V \pi'$  ( $s \leftrightarrow_V s'$ , Proposition 1)

$\Leftrightarrow s_i \leftrightarrow_V s'_i$  for each  $i \geq 0$  ( $\pi \leftrightarrow_V \pi'$ )

$\Leftrightarrow (\mathcal{M}', s'_i) \models \psi$  for each  $i \geq 0$  (induction hypothesis)

$\Leftrightarrow (\mathcal{M}', s') \models \phi$

**Case**  $\phi = \text{E}[\psi_1 \cup \psi_2]$ :

$\mathcal{M}, s \models \phi$

$\Leftrightarrow$  There is a path  $\pi = (s = s_0, s_1, \dots)$  such that there is  $i \geq 0$  such that  $(\mathcal{M}, s_i) \models \psi_2$ , and for all  $0 \leq j < i$ ,

$(\mathcal{M}, s_j) \models \psi_1$   
 $\Leftrightarrow$  There is a path  $\pi' = (s = s'_0, s'_1, \dots)$  such that  $\pi \leftrightarrow_V \pi'$   
 $(s \leftrightarrow_V s', \text{Proposition 1})$   
 $\Leftrightarrow (\mathcal{M}', s'_i) \models \psi_2$ , and for all  $0 \leq j < i$   $(\mathcal{M}', s'_j) \models \psi_1$   
(induction hypothesis)  
 $\Leftrightarrow (\mathcal{M}', s') \models \phi$   $\square$

**Proposition 3** Let  $\varphi$  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  $T_i$ . We will prove this proposition from the following several aspects:

(1)  $\varphi \equiv_{\langle \{p\}, \emptyset \rangle} T_0$ .

$(\Rightarrow) \forall (\mathcal{M}_1, s_1) \in \text{Mod}(\varphi)$ , i.e.  $(\mathcal{M}_1, s_1) \models \varphi$ . We can construct an Ind-model structure  $\mathcal{M}_2$  is identical to  $\mathcal{M}_1$  except  $L_2(s_2) = L_1(s_1) \cup \{p\}$ . It is apparent that  $(\mathcal{M}_2, s_2) \models T_0$  and  $(\mathcal{M}_1, s_1) \leftrightarrow_{\langle \{p\}, \emptyset \rangle} (\mathcal{M}_2, s_2)$ .

$(\Leftarrow) \forall (\mathcal{M}_1, s_1) \in \text{Mod}(T_0)$ , it is apparent that  $(\mathcal{M}_1, s_1) \models \varphi$  by the semantic of **start**.

By  $\psi \rightarrow_t R_i$  we mean using transformation rules  $t$  on formula  $\psi$  (the formulae  $\psi$  as the premises of rule  $t$ ) and obtaining the set  $R_i$  of transformation results. Let  $X$  be a set of formulas we will show  $T_i \equiv_{\langle V', I \rangle} T_{i+1}$  by using the transformation rule  $t$ . Where  $T_i = X \cup \{\psi\}$ ,  $T_{i+1} = X \cup R_i$ ,  $V'$  is the set of atoms introduced by  $t$  and  $I$  is the set of indexes introduced by  $t$ . (We will prove this result in  $t \in \{\text{Trans}(1), \text{Trans}(4), \text{Trans}(6)\}$ , other cases can be proved similarly.)

(2) For  $t = \text{Trans}(1)$ :

$(\Rightarrow) \forall (\mathcal{M}_1, s_1) \in \text{Mod}(T_i)$  i.e.  $(\mathcal{M}_1, s_1) \models X \wedge \text{AG}(q \supset \text{EX}\varphi)$   
 $\Rightarrow (\mathcal{M}_1, s_1) \models X$  and for every  $\pi$  starting from  $s_1$  and every state  $s_1^j \in \pi$ ,  $(\mathcal{M}, s_1^j) \models \neg q$  or there exists a path  $\pi'$  starting from  $s_1^j$  such that there exists a state  $s_1^{j+1}$  such that  $(s_1^j, s_1^{j+1}) \in R_1$  and  $(\mathcal{M}, s_1^{j+1}) \models \varphi$

We can construct an Ind-model structure  $\mathcal{M}_2$  is identical to  $\mathcal{M}_1$  except  $[\text{ind}]_2 = \bigcup_{s \in S} R_s \cup R_y$ , where  $R_{s_1^j} = \{(s_1^j, s_1^{j+1}), (s_1^{j+1}, s_1^{j+2}), \dots\}$  and  $R_y = \{(s_x, s_y) | \forall s_x \in S \text{ if } \forall (s'_1, s'_2) \in \bigcup_{s \in S} R_s, s'_1 \neq s_x \text{ then find a unique } s_y \in S \text{ such that } (s_x, s_y) \in R\}$ . It is apparent that  $(\mathcal{M}_1, s_1) \leftrightarrow_{\langle \emptyset, \{\text{ind}\} \rangle} (\mathcal{M}_2, s_2)$  (let  $s_2 = s_1$ ).

$\Rightarrow$  for every path starting from  $s_1$  and every state  $s_1^j$  in this path,  $(\mathcal{M}_2, s_1^j) \models \neg q$  or  $(\mathcal{M}_2, s_1^j) \models \text{EX}\varphi_{\langle \text{ind} \rangle}$  (by the semantic of EX)

$\Rightarrow (\mathcal{M}_2, s_1) \models \text{AG}(q \supset \text{E}_{\langle \text{ind} \rangle} X\varphi)$

$\Rightarrow (\mathcal{M}_2, s_1) \models X \wedge \text{AG}(q \supset \text{E}_{\langle \text{ind} \rangle} X\varphi)$

$(\Leftarrow) \forall (\mathcal{M}_1, s_1) \in \text{Mod}(T_{i+1})$  i.e.  $(\mathcal{M}_1, s_1) \models X \wedge \text{AG}(q \supset \text{E}_{\langle \text{ind} \rangle} X\varphi)$

$\Rightarrow (\mathcal{M}_1, s_1) \models X$  and  $(\mathcal{M}_1, s_1) \models \text{AG}(q \supset \text{E}_{\langle \text{ind} \rangle} X\varphi)$

$\Rightarrow$  for every path starting from  $s_1$  and every state  $s_1^j$  in this path,  $(\mathcal{M}_1, s_1^j) \models \neg q$  or there exists a state  $s'$  such that  $(s_1^j, s') \in [\text{ind}]_1$  and  $(\mathcal{M}_1, s') \models \varphi$  (by the semantic of  $\text{E}_{\langle \text{ind} \rangle} X$ )

$\Rightarrow$  for every path starting from  $s_1$  and every state  $s_1^j$  in this path,  $(\mathcal{M}_1, s_1^j) \models \neg q$  or  $(\mathcal{M}_1, s_1^j) \models \text{EX}\varphi$  (by the semantic

of EX)

$\Rightarrow (\mathcal{M}_1, s_1) \models \text{AG}(q \supset \text{EX}\varphi)$

$\Rightarrow (\mathcal{M}_1, s_1) \models X \wedge \text{AG}(q \supset \text{EX}\varphi)$

It is apparent that  $(\mathcal{M}_1, s_1) \leftrightarrow_{\langle \emptyset, \{\text{ind}\} \rangle} (\mathcal{M}_1, s_1)$ .

(3) For  $t = \text{Trans}(4)$ :

$(\Rightarrow) \forall (\mathcal{M}_1, s_1) \in \text{Mod}(T_i)$ , i.e.  $(\mathcal{M}_1, s_1) \models X \wedge \text{AG}(q \supset \varphi_1 \vee \varphi_2)$

$\Rightarrow (\mathcal{M}_1, s_1) \models X$  and  $\forall s'_1 \in S$ ,  $(\mathcal{M}_1, s'_1) \models q \supset \varphi_1 \vee \varphi_2$

$\Rightarrow (\mathcal{M}_1, s'_1) \models \neg q$  or  $(\mathcal{M}_1, s'_1) \models \varphi_1 \vee \varphi_2$

The we can construct an Ind-model structure  $\mathcal{M}_2$  as follows.  $\mathcal{M}_2$  is the same with  $\mathcal{M}_1$  when  $(\mathcal{M}_1, s'_1) \models \neg q$ . When  $(\mathcal{M}_1, s'_1) \models q$ ,  $\mathcal{M}_2$  is identical to  $\mathcal{M}_1$  except if  $(\mathcal{M}_1, s'_1) \models \varphi_1$  then  $L_2(s'_1) = L_1(s'_1)$  else  $L_2(s'_1) = L_1(s'_1) \cup \{p\}$ . It is apparent that  $(\mathcal{M}_2, s'_1) \models (q \supset \varphi_1 \vee p) \wedge (p \supset \varphi_2)$ , then  $(\mathcal{M}_2, s_1) \models T_{i+1}$  and  $(\mathcal{M}_1, s_1) \leftrightarrow_{\langle \{p\}, \emptyset \rangle} (\mathcal{M}_2, s_2)$ .

$(\Leftarrow) \forall (\mathcal{M}_1, s_1) \in \text{Mod}(T_{i+1})$ , i.e.  $(\mathcal{M}_1, s_1) \models X \wedge \text{AG}(q \supset \varphi_1 \vee p) \wedge \text{AG}(p \supset \varphi_2)$ . It is apparent that  $(\mathcal{M}_1, s_1) \models T_i$ .

(4) For  $t = \text{Trans}(6)$ :

We prove for  $\text{E}_{\langle \text{ind} \rangle} X$ , while for the AX can be proved similarly.

$(\Rightarrow) \forall (\mathcal{M}_1, s_1) \in \text{Mod}(T_i)$ , i.e.  $(\mathcal{M}_1, s_1) \models X \wedge \text{AG}(q \supset \text{E}_{\langle \text{ind} \rangle} X\varphi)$

$\Rightarrow (\mathcal{M}_1, s_1) \models X$  and  $\forall s'_1 \in S$ ,  $(\mathcal{M}_1, s'_1) \models q \supset \text{E}_{\langle \text{ind} \rangle} X\varphi$

$\Rightarrow (\mathcal{M}_1, s'_1) \models \neg q$  or there exists a state  $s'$  such that  $(s'_1, s') \in [\text{ind}]$  and  $(\mathcal{M}_1, s') \models \varphi$

We can construct an Ind-model structure  $\mathcal{M}_2$  as follows.  $\mathcal{M}_2$  is the same with  $\mathcal{M}_1$  when  $(\mathcal{M}_1, s'_1) \models \neg q$ . When  $(\mathcal{M}_1, s'_1) \models q$ ,  $\mathcal{M}_2$  is identical to  $\mathcal{M}_1$  except for  $s'$  there is  $L_2(s') = L_1(s') \cup \{p\}$ . It is apparent that  $(\mathcal{M}_2, s_1) \models \text{AG}(q \supset \text{E}_{\langle \text{ind} \rangle} Xp) \wedge \text{AG}(p \supset \varphi)$ ,  $(\mathcal{M}_2, s_2) \models T_{i+1}$  and  $(\mathcal{M}_1, s_1) \leftrightarrow_{\langle \{p\}, \emptyset \rangle} (\mathcal{M}_2, s_2)$  ( $s_2 = s_1$ ).

$(\Leftarrow) \forall (\mathcal{M}_1, s_1) \in \text{Mod}(T_{i+1})$ , i.e.  $(\mathcal{M}_1, s_1) \models X \wedge \text{AG}(q \supset \text{E}_{\langle \text{ind} \rangle} Xp) \wedge \text{AG}(p \supset \varphi)$ . It is apparent that  $(\mathcal{M}_1, s_1) \models T_i$ .  $\square$

**Proposition 4** Let  $\varphi$  be a CTL formula, then  $T_\varphi \equiv_{\langle V \cup V', \emptyset \rangle} \text{Res}$ .

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

By  $\psi \rightarrow_r R_i$  we mean using resolution rules  $r$  on set  $\psi$  (the formulae in  $\psi$  as the premises of rule  $r$ ) and obtaining the set  $R_i$  of resolution results. we will show  $T_i \equiv_{\langle V, I \rangle} T_{i+1}$  by using the resolution rule  $r$ . Where  $T_i = X \cup \psi$ ,  $T_{i+1} = X \cup R_i$ ,  $X$  be a set of  $\text{SNF}_{\text{CTL}}^q$  clauses,  $p$  be the proposition corresponding with literal  $l$  used to do resolution in  $r$ .

(1) If  $\psi \rightarrow_r R_i$  by an application of  $r \in \{(\text{SRES1}), \dots, (\text{SRES8}), \text{RW1}, \text{RW2}\}$ , then  $T_i \equiv_{\langle \{p\}, \emptyset \rangle} T_{i+1}$ .

On one hand, it is apparent that  $\psi \models R_i$  and then  $T_i \models T_{i+1}$ . On the other hand,  $T_i \subseteq T_{i+1}$  and then  $T_{i+1} \models T_i$ .

(2) If  $\psi \rightarrow_r R_i$  by an application of  $r = (\text{ERES1})$ , then  $T_i \equiv_{\langle \{l, w_{\neg l}^{\neg l}\}, \emptyset \rangle} T_{i+1}$ .

It has been proved that  $\psi \models R_i$  in (Bolotov 2000), then there is  $T_{i+1} = T_i \cup \Lambda_{\neg l}^{\neg l}$  and then  $\forall (\mathcal{M}_1, s_1) \in \text{Mod}(T_i = X \cup \psi)$  there is a  $(\mathcal{M}_2, s_2) \in \text{Mod}(T_{i+1} = T_i \cup \Lambda_{\neg l}^{\neg l})$  s.t.

$(\mathcal{M}_1, s_1) \leftrightarrow_{\langle\{p, w_{-l}\}, \emptyset\rangle} (\mathcal{M}_2, s_2)$  and vice versa by Proposition 3.

For rule (ERES2) we have the same result.  $\square$

**Proposition 6** Let  $V'' = V \cup V'$ ,  $\Gamma = \text{Instantiate}(\text{Res}, V')$  and  $\Gamma_1 = \text{Removing\_atoms}(\text{Connect}(\Gamma), \Gamma)$ , then  $\Gamma_1 \equiv_{\langle V'', \emptyset \rangle} \text{Res}$  and  $\Gamma_1 \equiv_{\langle V'', I \rangle} \varphi$ .

*Proof.* Take note the fact that for each clause  $C = T \supset H$  in  $\text{Connect}(\Gamma)$ , 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  $\text{Connect}(\Gamma) \models \Gamma_1$ , we will show  $\forall (\mathcal{M}, s_0) \in \text{Mod}(\Gamma_1)$  there is a  $(\mathcal{M}', s_0)$  such that  $(\mathcal{M}', s_0) \models \text{Connect}(\Gamma)$  and  $(\mathcal{M}, s_0) \leftrightarrow_{\langle \Gamma, \emptyset \rangle} (\mathcal{M}', s_0)$ . Let  $C = T \supset H$  in  $\text{Connect}(\Gamma)$  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  $(\mathcal{M}, s) \not\models T$  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  $p \in \text{Var}(H) \cap \Gamma$ , 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  $L'(s^*) = L(s^*) \setminus \{p\}$  for (each (if  $C$  is an A-step or A-sometime clause))  $s^* \in \pi_s$ , else if  $C$  is a global (or an initial) clause then let  $L'(s) = L(s) \cup \{p\}$  else let  $L'(s^*) = L(s^*) \cup \{p\}$  for (each (if  $C$  is a A-step or A-sometime clause))  $s^* \in \pi_s$ . Where  $s^*$  is a next or future state of  $s$  (it depends on the type of the clause: if the clause is a  $X$ -step ( $X \in \{A, E\}$ ) clause then  $s^*$  is the next state, else if the clause is a  $X$ -sometime clause then  $s^*$  is a future state).

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

- (1) For (i), it is apparent  $(\mathcal{M}', s_0) \models C$ ;
- (2) 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  $\Gamma$ :
  - (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  $C' = \top \supset \neg p \vee C_2 \in \text{Connect}(\Gamma)$ , then there is  $\top \supset C_1 \vee C_2 \in \text{Connect}(\Gamma)$  by the resolution  $((\mathcal{M}, s) \models C_2$  due to we have suppose  $(\mathcal{M}, s) \not\models C$ ). 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  $C' = T' \supset E_{\langle \text{ind} \rangle} X(\neg p \vee C_2) \in \text{Connect}(\Gamma)$ , then there is  $T \wedge T' \supset E_{\langle \text{ind} \rangle} X(C_1 \vee C_2) \in \text{Connect}(\Gamma)$  by the resolution  $((\mathcal{M}, s) \models E_{\langle \text{ind} \rangle} X C_2$  due to we have suppose  $(\mathcal{M}, s) \not\models C$ ). It is apparent that  $(\mathcal{M}', s_0) \models C \wedge C'$ .
  - (c) Other cases can be proved similarly.

Therefore, we have  $\Gamma_1 \equiv_{\langle V'', \emptyset \rangle} \text{Res}$  by Proposition 2 and Proposition 5.

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

## References

- Akintunde, M. E. 2017. Planning for ctl\*-specified temporally extended goals via model checking.
- Baier, C., and Katoen, J. 2008. *Principles of Model Checking*. The MIT Press.
- Bolotov, A. 1999. A clausal resolution method for ctl branching-time temporal logic. *Journal of Experimental & Theoretical Artificial Intelligence* 11(1):77–93.
- Bolotov, A. 2000. *Clausal resolution for branching-time temporal logic*. Ph.D. Dissertation, Manchester Metropolitan University.
- Browne, M. C.; Clarke, E. M.; and Grumberg, O. 1988. Characterizing finite kripke structures in propositional temporal logic. *Theor. Comput. Sci.* 59:115–131.
- Clarke, E. M., and Emerson, E. A. 1981. Design and synthesis of synchronization skeletons using branching time temporal logic. In *Workshop on Logic of Programs*, 52–71. Springer.
- Clarke, E. M.; Emerson, E. A.; and Sistla, A. P. 1986. Automatic verification of finite-state concurrent systems using temporal logic specifications. *ACM Trans. Program. Lang. Syst.* 8(2):244–263.
- Eiter, T., and Kern-Isberner, G. 2019. A brief survey on forgetting from a knowledge representation and reasoning perspective. *KI-Künstliche Intelligenz* 33(1):9–33.
- Eiter, T., and Wang, K. 2008. *Semantic forgetting in answer set programming*. Elsevier Science Publishers Ltd.
- Fang, L.; Liu, Y.; and Van Ditmarsch, H. 2019. Forgetting in multi-agent modal logics. *Artificial Intelligence* 266:51–80.
- Fisher, M.; Dixon, C.; and Peim, M. 2001. Clausal temporal resolution. *ACM Transactions on Computational Logic (TOCL)* 2(1):12–56.
- Fisher, M. 1991. A resolution method for temporal logic. In *Ijcai*, volume 91, 99–104. Citeseer.
- Fisher, M. 1997. A normal form for temporal logics and its applications in theorem-proving and execution. *Journal of Logic and Computation* 7(4):429–456.
- Giunchiglia, F., and Traverso, P. 1999. Planning as model checking. In *European Conference on Planning*, 1–20. Springer.
- Herzig, A., and Mengin, J. 2008. Uniform interpolation by resolution in modal logic. In *European Workshop on Logics in Artificial Intelligence*, 219–231. Springer.
- Lang, J., and Marquis, P. 2008. On propositional definability. *Artificial Intelligence* 172(8):991–1017.
- Lang, J., and Marquis, P. 2010. *Reasoning under inconsistency: a forgetting-based approach*. Elsevier Science Publishers Ltd.
- Lin, F., and Reiter, R. 1994. Forget it. In *Working Notes of AAAI Fall Symposium on Relevance*, 154–159.
- Lin, F. 2001. On strongest necessary and weakest sufficient conditions. *Artif. Intell.* 128(1-2):143–159.
- Lin, F. 2003. Compiling causal theories to successor state axioms and strips-like systems. *Journal of Artificial Intelligence Research* 19:279–314.

- Liu, Y., and Wen, X. 2011. On the progression of knowledge in the situation calculus. In *IJCAI 2011, Proceedings of the 22nd International Joint Conference on Artificial Intelligence*, 976–982. Barcelona, Catalonia, Spain: IJCAI/AAAI.
- Lutz, C., and Wolter, F. 2011. Foundations for uniform interpolation and forgetting in expressive description logics. In *IJCAI 2011, Proceedings of the 22nd International Joint Conference on Artificial Intelligence*, 989–995. Barcelona, Catalonia, Spain: IJCAI/AAAI.
- Su, K.; Sattar, A.; Lv, G.; and Zhang, Y. 2009. Variable forgetting in reasoning about knowledge. *Journal of Artificial Intelligence Research* 35:677–716.
- Wang, Z.; Wang, K.; Topor, R. W.; and Pan, J. Z. 2010. Forgetting for knowledge bases in DL-Lite. *Annals of Mathematics and Artificial Intelligence* 58(1-2):117–151.
- Wang, Y.; Zhang, Y.; Zhou, Y.; and Zhang, M. 2012. Forgetting in logic programs under strong equivalence. In *Principles of Knowledge Representation and Reasoning: Proceedings of the Thirteenth International Conference*, 643–647. Rome, Italy: AAAI Press.
- Wang, Y.; Wang, K.; and Zhang, M. 2013. Forgetting for answer set programs revisited. In *IJCAI 2013, Proceedings of the 23rd International Joint Conference on Artificial Intelligence*, 1162–1168. Beijing, China: IJCAI/AAAI.
- Wang, Y. 2015. On forgetting in tractable propositional fragments. <http://arxiv.org/abs/1502.02799>.
- Wong, K.-S. 2009. *Forgetting in Logic Programs*. Ph.D. Dissertation, The University of New South Wales.
- Zhang, Y., and Foo, N. Y. 2006. Solving logic program conflict through strong and weak forgettings. *Artificial Intelligence* 170(8-9):739–778.
- Zhang, Y., and Zhou, Y. 2009. Knowledge forgetting: Properties and applications. *Artificial Intelligence* 173(16-17):1525–1537.
- Zhang, Y.; Foo, N. Y.; and Wang, K. 2005. Solving logic program conflict through strong and weak forgettings. In *Ijcai-05, Proceedings of the Nineteenth International Joint Conference on Artificial Intelligence, Edinburgh, Scotland, Uk, July 30-August*, 627–634.
- Zhang, L.; Hustadt, U.; and Dixon, C. 2008. First-order resolution for ctl. Technical report, Citeseer.
- Zhang, L.; Hustadt, U.; and Dixon, C. 2009. A refined resolution calculus for ctl. In *International Conference on Automated Deduction*, 245–260. Springer.
- Zhang, L.; Hustadt, U.; and Dixon, C. 2014. A resolution calculus for the branching-time temporal logic ctl. *ACM Transactions on Computational Logic (TOCL)* 15(1):1–38.
- Zhao, Y., and Schmidt, R. A. 2017a. Role forgetting for alcoh(δ)-ontologies using an ackermann-based approach. In *Proceedings of the 26th International Joint Conference on Artificial Intelligence*, 1354–1361. AAAI Press.
- Zhao, Y., and Schmidt, R. A. 2017b. Role forgetting for alcoh(Δ)-ontologies using an ackermann-based approach. In *Proceedings of the 26th International Joint Conference on Artificial Intelligence, IJCAI’17*, 1354–1361. AAAI Press.