

# The Complexity of Epistemic Model Checking: Clock Semantics and Branching Time<sup>1</sup>

X. Huang and R. van der Meyden<sup>2</sup>

**Abstract.** In the clock semantics for epistemic logic, two situations are indistinguishable for an agent when it makes the same observation and the time in the situations is the same. The paper characterizes the complexity of model checking branching time logics of knowledge in finite state systems with respect to the clock semantics.

## 1 Introduction

Epistemic logic has been shown to provide a useful formalism for reasoning about systems in which the representation of agents' states of uncertainty is a critical factor [6]. The range of problems to which epistemic logic has been applied encompasses distributed and multi-agent systems, computer security, diagnosis and recoverability. These applications have motivated the development of verification techniques based on logics that combine temporal and epistemic features.

In particular, model checking combinations of temporal and epistemic logics has been a topic of recent interest. In model checking, the verification problem is treated as the problem of checking that a formula is satisfied in a given model. For suitably restricted representations of models and specification languages this problem can be shown to be decidable. Model checkers implement this decidable problem using a variety of sophisticated heuristics and symbolic implementation techniques.

Operators in the logic of knowledge can be given a number of different semantics, depending on the basic information from which an agent derives what it knows. For example, we can treat agent's knowledge as those facts that it can derive from just its current observation. This semantics is basis of much of the literature on model checking temporal and epistemic logics. However, the model checking problem can also be shown to be decidable for stronger interpretations of knowledge. In this paper, we consider the model checking problem for an interpretation of knowledge in which an agent's knowledge is taken to be what it can derive from its current observation plus the current clock value. We call this the clock semantics for knowledge. The significance of the clock semantics is that many systems are built with clocks, and they are used in protocols, e.g., for timeouts. The observational semantics is too weak to capture information present in clock values.

The model checking problem with respect to the clock semantics has previously been shown to be decidable for linear time temporal logics extended by knowledge operators by Engelhardt et al [5]. In this paper we consider the effect of taking the temporal basis for

the specification language to be instead a branching time temporal logic. We show that this combination also leads to a decidable model checking problem, and characterise its complexity for a number of different fragments of a logic  $CTL^*K_n^{clk}$  that combines the linear time operators, branching operators and epistemic operators.

The main results of the paper are presented in Table 1, which gives the complexity (both upper and lower bounds) of each of the model checking problems we consider. For purposes of comparison, we include in this table the known results for the complexity of the linear time epistemic logic with respect to clock semantics ( $LTLK_n^{clk}$ ).

| Logic                                | Bound          | Combined Complexity       | Model Complexity          | Formula Complexity   |
|--------------------------------------|----------------|---------------------------|---------------------------|----------------------|
| $LTLK_n^{clk}$<br>(from [5])         | upper<br>lower | PSPACE<br>PSPACE          | PSPACE<br>PH              | PSPACE<br>PSPACE     |
| $CTL^*K_n^{clk}$<br>(this paper)     | upper<br>lower | PSPACE<br>PSPACE          | PSPACE<br>PH              | PSPACE<br>PSPACE     |
| $CTLK_n^{clk}$<br>(this paper)       | upper<br>lower | PSPACE<br>PSPACE          | PSPACE<br>PH              | LOGSPACE<br>LOGSPACE |
| $CTL \neg K_n^{clk}$<br>(this paper) | upper<br>lower | $P^{NP}$<br>$P^{NP[log]}$ | $P^{NP}$<br>$P^{NP[log]}$ | LOGSPACE<br>LOGSPACE |

**Table 1.** Complexity Results

It turns out that the addition of branching operators to the combination  $LTLK_n^{clk}$  of linear time temporal logic and epistemic logic (giving the logic  $CTL^*K_n^{clk}$ ) does not increase the computational complexity of model checking, which remains PSPACE-complete. More interestingly, the model checking problem for the fragment  $CTLK_n^{clk}$  based in the branching time logic CTL has the same complexity as for the linear time case ( $LTLK_n^{clk}$ ), viz., PSPACE-complete. Prima facie, this result is a little surprising since the complexities of model checking linear time temporal logic and branching time logic in the absence of epistemic operators are known to be different, viz., PTIME for the branching time temporal logic CTL and PSPACE-complete for linear time temporal logic LTL.

However, our general result masks some subtleties, and a more careful analysis reveals differences between the linear time and branching time cases. In particular, a difference is apparent if one considers the complexity of model checking as a function of the size of the model, or as a function of the size of the formula. *Formula complexity* is the complexity of model checking a varying formula when the model is held fixed: this gives a measure of the complexity of the model checking as a function of the size of the formula. Alternately, *model complexity* is the complexity of the model checking problem when a formula is held fixed and the model is varied:

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<sup>2</sup> School of Computer Science and Engineering, University of New South Wales, email: {xiaowei, meyden}@cse.unsw.edu.au

this gives a measure of the complexity of the model checking problem as a function of the size of the model. We find that the complexity difference between CTL and LTL continues to be reflected when one adds epistemic operators in the case of formula complexity, where  $\text{CTL}_n^{\text{clk}}$  is LOGSPACE-complete and  $\text{LTLK}_n^{\text{clk}}$  is PSPACE-complete (which is the same as the formula complexities for CTL and LTL respectively.) However, with respect to model complexity there is no change.

We also explore the impact of a further restriction on the set of branching time operators, taking these to be just  $EF$  (at some future time in some branch) and  $EX$  (at some successor), giving the logic  $\text{CTL}^-K_n^{\text{clk}}$ . Here, we show that the complexity of model checking falls down to a low level of the polynomial hierarchy, viz  $\text{P}^{\text{NP}}$ .

The structure of the paper is as follows. In section 2, we define the syntax and semantics of the logics that we study, as well as the model checking problems that we consider. Section 3 proves the main complexity results of the paper. Section 4 discusses related work; in particular, we argue that some previous work on bounded model checking  $\text{CTLK}_n^{\text{clk}}$  is incorrect.

## 2 Syntax and Semantics

We work with logics that combine temporal logics and the logic of knowledge and common knowledge for  $n$  agents. All the logics that we consider are fragments of the logic  $\text{CTL}^*K_n^{\text{clk}}$ . Let  $\text{Prop}$  be a set of atomic propositions and  $\text{Ags} = \{1, \dots, n\}$  be a set of  $n$  agents. The syntax of the logic  $\text{CTL}^*K_n^{\text{clk}}$  is given by the following grammar:

$$\begin{aligned} \phi ::= & p \mid \neg\phi \mid \phi_1 \wedge \phi_2 \mid X\phi \mid F\phi \mid \phi_1 U \phi_2 \mid G\phi \mid \\ & E\phi \mid A\phi \mid \\ & K_i\phi \mid C_G\phi \end{aligned}$$

where  $p \in \text{Prop}$  and  $i \in \text{Ags}$  and  $G \in \mathcal{P}(\text{Ags}) \setminus \{\emptyset\}$ . The first line gives basic propositional logic plus linear time temporal operators that refer to the future. Intuitively,  $X\phi$  says that  $\phi$  holds at the next time,  $F\phi$  says that  $\phi$  holds at some future time,  $\phi_1 U \phi_2$  says that  $\phi_1$  holds until  $\phi_2$  does, and  $G\phi$  says that  $\phi$  holds at all times in the future. The operators in the second line are from branching time temporal logic and refer to possible alternate futures:  $A\phi$  says that  $\phi$  holds in all possible futures and  $E\phi$  says that  $\phi$  holds in some possible future. The final line gives epistemic operators:  $K_i\phi$  says that agent  $i$  knows  $\phi$  and  $C_G\phi$  says that  $\phi$  is common knowledge to the group of agents  $G$ . If we take just the first line of this grammar we have the linear time temporal logic LTL, and adding the second line gives the branching temporal logic  $\text{CTL}^*$ . The logic  $\text{LTLK}_n^{\text{clk}}$  is the fragment obtained by taking just the first and the third lines of the grammar. The branching time temporal logic CTL is obtained from  $\text{CTL}^*$  by placing a restriction on the permitted combinations of linear and branching time temporal operators: we replace these in the grammar by the restricted cases  $QX\phi$ ,  $Q\phi_1 U \phi_2$ ,  $QF\phi$  and  $QG\phi$ , where  $Q$  is either  $A$  or  $E$ . Adding the epistemic operators to CTL gives the logic  $\text{CTLK}_n^{\text{clk}}$ . The fragment  $\text{CTL}^-K_n^{\text{clk}}$  is obtained by combining the epistemic operators with the branching temporal operators  $EX$  and  $EF$ . Dually, this logic contains  $AX$  and  $AG$ , so can express the important class of *epistemic safety* properties.

To give semantics to all these logics it suffices to give semantics to  $\text{CTL}^*K_n^{\text{clk}}$ . We do this using a variant of interpreted systems [6], specialised to the clock semantics. Let  $S$  be a set, which we call the set of global states. A *run* over  $S$  is a function  $r : \mathbb{N} \rightarrow S$ . A *point* is a pair  $(r, m)$  where  $r$  is a run and  $m \in \mathbb{N}$ . Given a set  $\mathcal{R}$  of runs, we define  $\text{Points}(\mathcal{R})$  to be the set of all points of runs  $r \in \mathcal{R}$ . An *interpreted system* for  $n$  agents is a tuple  $I = (\mathcal{R}, \sim_1, \dots, \sim_n, \pi)$ , where  $\mathcal{R}$  is a

set of runs over  $S$ , each  $\sim_i$  is an equivalence relation on  $\text{Points}(\mathcal{R})$  (called agent  $i$ 's *indistinguishability relation*) and  $\pi : S \rightarrow \mathcal{P}(\text{Prop})$  is an interpretation function. We say that a run  $r'$  is *equivalent to a run  $r$  up to time  $m \in \mathbb{N}$*  if  $r'(k) = r(k)$  for  $0 \leq k \leq m$ .

We can define a general semantics of  $\text{CTL}^*K_n^{\text{clk}}$  by means of a relation  $I, (r, m) \models \phi$ , where  $I$  is an interpreted system,  $(r, m)$  is a point of  $I$  and  $\phi$  is a formula. This relation is defined inductively as follows:

- $I, (r, m) \models p$  if  $p \in \pi(r(m))$ ,
- $I, (r, m) \models \neg\phi$  if not  $I, (r, m) \models \phi$
- $I, (r, m) \models \phi_1 \vee \phi_2$  if  $I, (r, m) \models \phi_1$  or  $I, (r, m) \models \phi_2$
- $I, (r, m) \models E\phi$  if there exists a run  $r' \in \mathcal{R}$  equivalent to  $r$  up to time  $m$  such that  $I, (r', m) \models \phi$
- $I, (r, m) \models A\phi$  if for all runs  $r' \in \mathcal{R}$  equivalent to  $r$  up to time  $m$ , we have  $I, (r', m) \models \phi$
- $I, (r, m) \models X\phi$  if  $I, (r, m+1) \models \phi$
- $I, (r, m) \models \phi_1 U \phi_2$  if there exists  $m' \geq m$  such that  $I, (r, m') \models \phi_2$ , and  $I, (r, k) \models \phi_1$  for  $m \leq k < m'$ .
- $I, (r, m) \models G\phi$  if  $I, (r, k) \models \phi$  for all  $k \geq m$
- $I, (r, m) \models K_i\phi$  if for all points  $(r', m')$  of  $I$  such that  $(r, m) \sim_i (r', m')$ , we have  $I, (r', m') \models \phi$
- $I, (r, m) \models C_G\phi$  if for all sequences of points  $(r, m) = (r_0, m_0), (r_1, m_1), \dots, (r_k, m_k)$  of  $I$ , such that for each  $j = 0 \dots k-1$ , there exists  $i \in G$  such that  $(r_j, m_j) \sim_i (r_{j+1}, m_{j+1})$ , we have  $I, (r_k, m_k) \models \phi$ .

For the knowledge operators, this semantics is essentially the same as the usual interpreted systems semantics. For the temporal operators, it corresponds to a semantics for branching time known as the *bundle semantics* [2, 21]. To specialise this general semantics to the clock semantics, we suppose that we have for each agent  $i$  an observation function  $O_i : S \rightarrow \mathcal{O}$ , for some set  $\mathcal{O}$ , such that  $O_i(s)$  represents agent  $i$ 's observation in state  $s$ . Say that the equivalence relations  $\sim_i$  in the system are *derived from the observation functions*  $O_i$  when  $(r, m) \sim_i (r', m')$  iff  $m = m'$  and  $O_i(r(m)) = O_i(r'(m'))$ . A system is a *clock system* if there exist observation functions from which its indistinguishability relations  $\sim_i$  are derived. Intuitively, in such systems, an agent's knowledge is determined from its current observation plus the clock value.

For *model checking* we require the decidability of the problem of checking that a formula holds in a model. In order to do so, we require a finite state representation for the model. Interpreted systems are unsuitable for this, since they are based on infinite runs. We therefore treat interpreted systems as *generated* from an alternate finite representation. Define a (finite) *model* to be a tuple  $M = (S, I, \Rightarrow, O, \pi)$  where  $S$  is a (finite) set of states,  $I \subseteq S$  is the set of initial states,  $\Rightarrow \subseteq S \times S$  is a serial temporal transition relation,  $O = \{O_i\}_{i \in \text{Ags}}$  is a family of observation functions  $O_i : S \rightarrow \mathcal{O}$ , and  $\pi : S \rightarrow \mathcal{P}(\text{Prop})$  is a propositional interpretation. We write  $\mathcal{E}$  for the set of all finite models

Given a model  $M$  with states  $S$ , we may construct clock system  $I(M) = (\mathcal{R}, \sim_1, \dots, \sim_n, \pi)$  over global states  $S$ , as follows. The component  $\pi$  in  $I(M)$  is identical to that in  $M$ . The set of runs is defined as follows. We say that a *fullpath* from a state  $s$  is an infinite sequence of states  $s_0 s_1 \dots$  such that  $s_0 = s$  and  $s_i \Rightarrow s_{i+1}$  for all  $i \geq 0$ . We use  $\text{Path}(s)$  to denote the set of all fullpaths from state  $s$ . A *run* of the system is a fullpath  $s_0 s_1 \dots$  with  $s_0 \in I$ . We define  $\mathcal{R}$  to be the set of runs of  $M$ . Finally, we take the indistinguishability relations  $\sim_i$  to be the relations derived from the observation functions  $O_i$  in the environment.

A formula  $\phi$  is said to *hold* in a model  $M$ , written  $M \models \phi$ , if

$I(M), (r, 0) \models \phi$  for all  $r \in \mathcal{R}$ . The model checking problem we study is defined as follows: given a finite model  $M$  and a formula  $\phi$ , determine if  $M \models \phi$ . We are interested in this problem for a range of different languages  $\mathcal{L}$ , and as a function of its parameters as well as in general. More precisely, the *combined complexity* of model checking  $\mathcal{L}$  is the complexity of the set  $\{(M, \phi) \in \mathcal{E} \times \mathcal{L} \mid M \models \phi\}$ . The *model complexity* of a fixed formula  $\phi$  is the complexity of the set  $\{M \in \mathcal{E} \mid M \models \phi\}$ . This gives a measure of the complexity of model checking as a function of the size of the model. The *formula complexity* of  $\mathcal{L}$  for a fixed model  $M$  is the complexity of the set  $\{\phi \in \mathcal{L} \mid M \models \phi\}$ . This captures the contribution to the complexity of model checking that derives from the formula.

### 3 Complexity

We now consider the complexity of the model checking problems introduced in the previous section, for the language  $\text{CTL}^*K_n^{\text{clk}}$  and several of its sublanguages. For purposes of comparison, we first recall a number of known results for model checking temporal and epistemic logics.

#### Previous Results

Our logics build on the temporal logics LTL, CTL and  $\text{CTL}^*$ , whose model checking complexities are already well understood. In the case of  $\text{CTL}^*$ , the combined complexity is known to be PSPACE-complete [3, 4]. Indeed, this is already the case for the linear time logic LTL [1]. On the other hand, the logic CTL has a combined complexity in PTIME [3, 13]. A fortiori, the model and formula complexities of CTL are also in PTIME. For both LTL and  $\text{CTL}^*$ , the model complexity is PTIME [22, 8, 13], and the formula complexity is PSPACE-complete [13]. Thus, the complexity of model checking linear time temporal logic derives primarily from the contribution made by the formula. Since the formula we wish to check is typically small (but the model may be large) this result explains the feasibility in practice of LTL model checking even in the face of its PSPACE-hardness in the sense of combined complexity.

Given these results, we easily obtain that the combined complexity and the formula complexity of  $\text{LTL}K_n^{\text{clk}}$  are at least PSPACE hard. It is shown in [5] that, in fact, the addition of epistemic operators (interpreted with respect to clock semantics) does not necessarily add to the complexity of LTL. Both the combined and formula complexities of  $\text{LTL}K_n^{\text{clk}}$  turn out to be in PSPACE (hence PSPACE-complete, since the fragment LTL is already PSPACE-hard). A difference is found in the case of model complexity, however. It is shown that whereas LTL has PTIME model complexity, for each level  $\Pi_k^p$  of the polynomial hierarchy, there exists a formula  $\phi \in \text{LTL}K_n^{\text{clk}}$  such that  $\{M \in \mathcal{E} \mid M \models \phi\}$  is  $\Pi_k^p$ -hard. An exact characterization of the model complexity of  $\text{LTL}K_n^{\text{clk}}$  remains an open problem. However, it is possibly a very hard problem: note that whereas QBF is PSPACE complete, and the QBF formulas starting with  $\forall$  and having  $k$ -alternations correspond to  $\Pi_k^p$ , it is not known whether the polynomial hierarchy is equal to PSPACE. It appears that the gap between the upper and lower bounds for model complexity of  $\text{LTL}K_n^{\text{clk}}$  may be related to this problem.

#### $\text{CTL}^*K_n^{\text{clk}}$

We now show that we can derive complexity bounds for  $\text{CTL}^*K_n^{\text{clk}}$  from the known results for  $\text{LTL}K_n^{\text{clk}}$ . Given a model  $M$  and a  $\text{CTL}^*K_n^{\text{clk}}$  formula  $\phi$ , we will show that the model checking problem

$M \models \phi$  is equivalent to another model checking problem  $M' \models \phi'$ , where  $\phi'$  is a  $\text{LTL}K_{n+1}^{\text{clk}}$  formula and  $M'$  and  $\phi'$  are of polynomial size w.r.t.  $M$  and  $\phi$ , respectively. With this equivalence, we can move all the complexity results for  $\text{LTL}K_n^{\text{clk}}$  to  $\text{CTL}^*K_n^{\text{clk}}$ . To obtain model  $M'$ , we add an agent  $\top$  to the model  $M$  and define its observations by  $O_\top(s) = s$  for each state  $s \in S$ . We take  $\phi'$  to be the formula obtained from  $\phi$  by replacing all  $A$  operators with  $K_\top$  and all  $E$  operators with  $\neg K_\top \neg$ . The correctness of the equivalence claim now follows by an induction using the argument in the following result:

**Lemma 1** *For all points  $(r, m)$  of  $I(M)$ , and  $\phi \in \text{CTL}^*K_n^{\text{clk}}$  we have  $I(M), (r, m) \models A\phi$  iff  $I(M'), (r, m) \models K_\top\phi$ .*

**Proof:** The direction from right to left is trivial. It is easily seen that the semantics of  $I(M), (r, m) \models A\phi$  refers to a subset of the points  $(r', m')$  referred to by the semantics of  $I(M), (r, m) \models K_\top\phi$ . (Note that in both cases the fact that we are using clock semantics means that  $m = m'$ .) Conversely, suppose that  $I(M), (r, m) \models A\phi$ . Considering  $I(M'), (r, m) \models K_\top\phi$ , let  $(r', m')$  be a point such that  $(r, m) \sim_\top (r', m')$ . Then  $m = m'$  and  $r(m) = O_\top(r(m)) = O_\top(r'(m')) = r'(m')$ . It follows that the sequence  $r'' = r(0)r(1) \dots r(m)r'[m+1..\infty]$  is a run. Since the operators in  $\phi$  refer only to the future, an easy induction shows that  $I(M'), (r'', m) \models \phi$  iff  $I(M'), (r', m) \models \phi$ . Note that  $r''$  is equivalent to time  $m$  to  $r$ . Hence, by assumption that  $I(M), (r, m) \models A\phi$ , we obtain that  $I(M), (r'', m) \models \phi$ , hence  $I(M'), (r'', m) \models \phi$ , from which it follows using the above observation that  $I(M'), (r', m) \models \phi$ . This completes the proof that  $I(M'), (r, m) \models K_\top\phi$ .  $\square$

This transformation immediately enables us to derive upper bounds of PSPACE for the combined, model and formula complexity of  $\text{CTL}^*K_n^{\text{clk}}$ . Since  $\text{LTL}K_n^{\text{clk}}$  is a sublanguage of  $\text{CTL}^*K_n^{\text{clk}}$ , lower bounds of PSPACE,  $\Pi_k^p$  (for any  $k$ , by some formula) and PSPACE respectively also follow directly.

#### $\text{CTL}K_n^{\text{clk}}$

The combined complexity and model complexity upper bounds for  $\text{CTL}^*K_n^{\text{clk}}$  are also upper bounds for  $\text{CTL}K_n^{\text{clk}}$ , because  $\text{CTL}K_n^{\text{clk}}$  is a sublanguage of  $\text{CTL}^*K_n^{\text{clk}}$ .

The logic CTL has formula complexity of LOGSPACE [13], which is lower than the PSPACE-complete formula complexity of LTL, as noted above. We now show that this difference remains reflected in the extended logic  $\text{CTL}K_n^{\text{clk}}$ . We present an algorithm that shows that this logic has LOGSPACE formula complexity, lower than the PSPACE-complete formula complexity for the extended logic  $\text{CTL}^*K_n^{\text{clk}}$  and matching the LOGSPACE formula complexity for CTL model checking.

Given a fixed model  $M = (S, I, \Rightarrow, \sim_1, \dots, \sim_n, \pi)$ , we construct another model  $M' = (S \times \mathcal{P}(S), I', \Rightarrow', \sim'_1, \dots, \sim'_n, \pi')$ , with states of the form  $(s, P)$  where  $s \in S$  and  $P \subseteq S$ , such that

1.  $(s, P) \in I'$  iff  $s \in I$  and  $P = I$ ,
2.  $(s, P) \Rightarrow' (t, Q)$  iff  $s \Rightarrow t$  and  $Q = \{t \in S \mid \exists s \in P (s \Rightarrow t)\}$ ,
3.  $(s, P) \sim'_i (t, Q)$  iff  $P = Q$  and  $s \sim_i t$ ,
4.  $\pi'(s, P) = \pi(s)$ .

Though  $M'$  might be exponentially larger than  $M$ , it is also a fixed structure. Now  $M \models \phi$  is equivalent to  $M' \models \phi$  for observational semantics defined as follows.

1.  $M', (s, P) \models p$  iff  $p \in \pi'(s, P)$ .

2.  $M', (s, P) \models \neg\phi$  iff not  $M', (s, P) \models \phi$ .
3.  $M', (s, P) \models EX\phi$  iff there exists  $(t, Q)$  such that  $M', (t, Q) \models \phi$  and  $(s, P) \Rightarrow' (t, Q)$ .
4.  $M', (s, P) \models E[\phi_1 U \phi_2]$  iff there exists a path  $(s, P) = (s_0, P_0) \Rightarrow' (s_1, P_1) \Rightarrow' \dots \Rightarrow' (s_m, P_m)$  such that  $M', (s_m, P_m) \models \phi_2$  and  $M', (s_i, P_i) \models \phi_1$  for  $0 \leq i \leq m-1$ .
5.  $M', (s, P) \models EG\phi$  iff there exists an infinite path  $(s, P) = (s_0, P_0) \Rightarrow' (s_1, P_1) \Rightarrow' \dots$ , such that  $M', (s_i, P_i) \models \phi_1$  for all  $i \geq 0$ .
6.  $M', (s, P) \models K_i\phi$  iff  $M', (t, Q) \models \phi$  for all  $(t, Q)$  with  $(s, P) \sim'_i (t, Q)$ .
7.  $M', (s, P) \models C_G\phi$  iff for all sequences  $(s, P) = (s_0, P_0), (s_1, P_1), \dots, (s_m, P_m)$  with  $m \geq 0$  and  $(s_i, P_i) \sim_j (s_{i+1}, P_{i+1})$  for  $j \in G$  and  $0 \leq i \leq m-1$ , we have  $M', (s_m, P_m) \models \phi$ .

By [13], for CTL,  $M \models \phi$  can be checked in space  $O(|M| \cdot \log|\phi|)$ . With respect to the observational semantics, the addition of epistemic operators to CTL does not increase the complexity of model checking, since we may easily reduce the epistemic transitions to a special type of temporal transition. Therefore, the formula complexity of  $\text{CTLK}_n^{\text{clk}}$  is in LOGSPACE.

Lower bounds for model checking  $\text{CTLK}_n^{\text{clk}}$  can be obtained from the proof of lower bounds for  $\text{LTLK}_n^{\text{clk}}$  in [5] by noting that the proofs of these lower bounds use structures of the following special form.

**Definition 1** Let  $M = (S, I, \Rightarrow, \sim_1, \dots, \sim_n, \pi)$  be a model. Say that  $M$  is a lasso-bundle if for any  $s \in S$ , there exists a unique state  $s'$  such that  $s \Rightarrow s'$ .  $M$  is a lasso-structure if  $M$  is a lasso-bundle and  $I$  is a singleton set.

For lasso-structures, the semantics of LTL and CTL are known to coincide [9, 17]. An LTL formula can be evaluated in a lasso-bundled model by a CTL model checker by prefixing each temporal operator by an A/E path quantifier which results in a CTL formula. The reverse also holds.

We show that a similar result applies to  $\text{LTLK}_n^{\text{clk}}$  and  $\text{CTLK}_n^{\text{clk}}$ . Define a transformation function  $f$  mapping a normalized  $\text{LTLK}_n^{\text{clk}}$  formula (in which all negative operators occur in front of atomic propositions) into a  $\text{CTLK}_n^{\text{clk}}$  formula.

1.  $f(p) = p$
2.  $f(\neg\phi) = \neg f(\phi)$
3.  $f(\phi_1 \vee \phi_2) = f(\phi_1) \vee f(\phi_2)$
4.  $f(X\phi) = AXf(\phi)$ ,  $f(F\phi) = AFf(\phi)$ ,  $f(G\phi) = AGf(\phi)$
5.  $f(\phi_1 U \phi_2) = A[f(\phi_1) U f(\phi_2)]$
6.  $f(Y\phi) = Yf(\phi)$ , where  $Y \in \{K_i, C_G\}$ .

The transformation prefixes each temporal operator with a universal operator A. The following proposition concludes that this transformation preserves the satisfiability of formulae under lasso-bundled models.

**Proposition 1** For any lasso-bundled model  $M$  and any  $\text{LTLK}_n^{\text{clk}}$  formula  $\phi$ , and run  $r$  of  $I(M)$ , we have  $I(M), (r, n) \models \phi \Leftrightarrow I(M), (r, m) \models f(\phi)$ .

**Proof:** By a straightforward induction on the construction of  $\phi$ . Note that it follows from the fact that  $M$  is lasso-bundled that if  $r'$  is a run equivalent to  $r$  up to time  $n$  then in fact  $r' = r$ . This implies that  $I(M), (r, m) \models \phi$  iff  $I(M), (r, m) \models A\phi$  iff  $I(M), (r, m) \models E\phi$ .  $\square$

The lower bounds of  $\text{LTLK}_n^{\text{clk}}$  logic are proved in [5] by a reduction from the satisfiability problem of QBF. For any alternation

depth  $k$ , a  $\text{LTLK}_n^{\text{clk}}$  formula  $\phi_k$  is constructed, such that for each QBF formula  $\Psi$  of alternation depth  $k$ , a lasso-bundled model  $M_\Psi$  can be constructed, such that  $\Psi$  is satisfiable iff  $I(M_\Psi) \models \phi_k$ . Now by Proposition 1,  $\phi_k$  can be further reduced to a  $\text{CTLK}_n^{\text{clk}}$  formula  $f(\phi_k)$ , and  $\Psi$  is satisfiable iff  $I(M_\Psi) \models f(\phi_k)$ . Therefore, we can move the lower bounds of combined complexity and model complexity for  $\text{LTLK}_n^{\text{clk}}$  to  $\text{CTLK}_n^{\text{clk}}$ . In particular, we conclude that the combined complexity is PSPACE-hard, and for each  $k$  there exists a formula whose model complexity is  $\Pi_k^p$ -hard. The lower bound for formula complexity can be deemed as LOGSPACE, since every problem in LOGSPACE is complete under log-space reductions.

## CTL<sup>-</sup>K<sub>n</sub><sup>clk</sup>

We reduce the problem of model checking  $\text{CTL}^{\text{-}}\text{K}_n^{\text{clk}}$  to a fragment of Presburger Arithmetic, called extended Min-Max Arithmetic (MMA) in [7]. We use the notation  $[i, j] = \{i, i+1, \dots, j\}$  and  $[j] = [1, j]$ . An extended MMA dag-formula  $\alpha$  based in a finite partial order  $(X, \leq)$  is a collection of definitions of variables  $(\alpha_i)_{i \in X}$ , where the definition for each  $\alpha_i$  is one of the following, with  $i > j, k$  and  $\sim \in \{\leq, \geq\}$  and  $m, n \in \mathbb{N}$  in each case: (1)  $\equiv m \bmod n$  with  $n > 0$  and  $m \in \mathbb{Z}/n\mathbb{Z}$ , (2)  $\sim n$ , (3)  $\neg\alpha_j$ , (4)  $\alpha_j \wedge \alpha_k$ , (5)  $\sim \max(\alpha_j, n)$ , (6)  $m \sim \max(\alpha_j, n)$ , (7)  $\alpha_j \odot a$ , with  $\odot \in \{+, -\}$  and  $a \in \mathbb{N}$  written in unary. The semantics of  $\alpha_i$  in the context of an extended MMA formula  $\alpha$ , written as  $\llbracket \alpha_i \rrbracket$ , is the set of natural numbers satisfying the definition of  $\alpha_i$ , e.g.,  $\llbracket \sim \max(\alpha_j, n) \rrbracket = \{k \in \mathbb{N} \mid k \sim \max(\llbracket \alpha_j \rrbracket \cap [0, n])\}$ , and  $\llbracket m \sim \max(\alpha_j, n) \rrbracket$  is either  $\mathbb{N}$  or  $\emptyset$  depending on whether  $m \sim \max(\llbracket \alpha_j \rrbracket \cap [0, n])$ , and  $\llbracket \alpha_j \odot a \rrbracket = \{k \mid k = n \odot a \geq 0, n \in \llbracket \alpha_j \rrbracket\}$ . The following two lemmas are from [7].

**Lemma 2** Given an extended MMA dag-formula  $\alpha = (\alpha_i)_{i \in X}$ , there exist numbers  $L_i$  and  $T_i$ , computable in PTIME, such that, if  $n_1, n_2 > T_i$  and  $n_1 \equiv n_2 \bmod L_i$  then  $n_1 \in \llbracket \alpha_i \rrbracket \Leftrightarrow n_2 \in \llbracket \alpha_i \rrbracket$ .

**Lemma 3** Given  $n \in \mathbb{N}$  and an extended MMA dag-formula  $\alpha$ , whether  $n \in \llbracket \alpha_i \rrbracket$  can be decided with complexity  $\Delta_2^P = P^{NP}$ .

Another result we use is Chrobak's normal form for nondeterministic finite automata [18]:

**Lemma 4** Given a model  $M$ , we may compute in polynomial time a collection of arithmetic progressions  $R(s, t) = \{a_i + b_i \mathbb{N} \mid i = 1..k\}$  such that  $\cup R(s, t) = \{n \in \mathbb{N} \mid s \Rightarrow^n t\}$ ,  $|R(s, t)| \leq |S|^2$  and  $a_i = O(|S|^2)$  and  $b_i = O(|S|)$ .

We write  $M(\psi)$  for  $\{(s, n) \in S \times \mathbb{N} \mid \exists r(I(M), (r, n) \models \psi \text{ and } r(n) = s)\}$ . (Note that  $I(M), (r, n) \models \psi$  and  $r(n) = s$  implies  $I(M), (r', n) \models \psi$  for all runs  $r'$  with  $r'(n) = s$ .) Let  $\text{subf}(\phi)$  be the set of subformulas of  $\phi$ , and let  $\mathbb{T}$  denote tautology.

**Lemma 5** Given a formula  $\phi \in \text{CTL}^{\text{-}}\text{K}_n^{\text{clk}}$  and a model  $M$ , let  $X = S \times (\text{subf}(\phi) \cup \{\mathbb{T}\})$ , equipped with the partial order defined by  $(s_1, \psi_1) \geq (s_2, \psi_2)$  iff  $s_1 = s_2$  and  $\psi_2 \in \text{subf}(\psi_1) \cup \{\mathbb{T}\}$ . There exists an extended MMA-dag  $(\alpha_x)_{x \in X}$ , computable in time polynomial in  $|M| + |\phi|$  such that for all  $(s, \psi) \in X$ , we have  $\llbracket \alpha_{(s, \psi)} \rrbracket = \{n \in \mathbb{N} \mid (s, n) \in M(\psi)\}$ .

**Proof:** We define  $\alpha_{(s, \psi)}$  by induction on  $\geq$ . First, note that by Lemma 4, for each pair  $s, t \in S$ , we may write the set  $\{n \mid s \Rightarrow^n t\}$  as a polynomial size union  $R(s, t)$  of arithmetic progressions  $a + b\mathbb{N}$ . Thus, we may take

$$\alpha_{(t, \mathbb{T})} = \bigwedge_{s \in I} \bigwedge_{a+b\mathbb{N} \in R(s, t)} \equiv a \bmod b$$

for the base case. In the inductive case, we proceed as follows:

1.  $\psi = p$ : here we take  $\alpha_{(s,p)} = \alpha_{(s,\mathbb{T})}$  if  $p \in \pi(s)$ , else we take  $\alpha_{(s,p)} = \mathbb{F}$  (where  $\mathbb{F} = (> 1) \wedge (< 1)$  so that  $\llbracket \mathbb{F} \rrbracket = \emptyset$ ).
2.  $\psi = \psi_1 \wedge \psi_2$ : here  $\alpha_{(s,\psi)} = \alpha_{(s,\psi_1)} \wedge \alpha_{(s,\psi_2)}$
3.  $\psi = \neg\psi_1$ : here  $\alpha_{(s,\psi)} = \alpha_{(s,\mathbb{T})} \wedge \neg\alpha_{(s,\psi_1)}$
4.  $\psi = K_i\psi_1$ : here

$$\alpha_{(s,\psi)} = \alpha_{(s,\mathbb{T})} \wedge \bigwedge_{t \in S, s \sim_i t} (\alpha_{(t,\mathbb{T})} \Rightarrow \alpha_{(t,\psi_1)})$$

5.  $\psi = C_G\psi_1$ : here we take

$$\alpha_{(s,\psi)} = \alpha_{(s,\mathbb{T})} \wedge \bigwedge_{t \in S} (\gamma(s, t, G, |S|) \Rightarrow \alpha_{(t,\psi_1)})$$

where  $\gamma(s, t, G, k)$  is an extended MMA-dag formula expressing times at which  $t$  is reachable from  $s$  in  $k$  steps through the relation  $\cup_{i \in G} \sim_i$ . This can be represented efficiently by the following induction: in the base cases,  $\gamma(s, t, G, 0) = \alpha_{(s,\mathbb{T})}$  if  $s = t$  and  $\mathbb{F}$  otherwise,  $\gamma(s, t, G, 1) = \alpha_{(s,\mathbb{T})} \wedge \alpha_{(t,\mathbb{T})}$  if  $s \sim_i t$  for some  $i \in G$ , otherwise  $\gamma(s, t, G, 1) = \mathbb{F}$ . Inductively, for  $k \geq 2$ , we define

$$\gamma(s, t, G, k) = \bigvee_{s' \in S} (\gamma(s, s', G, \lceil k/2 \rceil) \wedge \gamma(s', t, G, \lfloor k/2 \rfloor)).$$

6.  $\psi = EX\psi_1$ : here  $\alpha_{(s,\psi)} = \alpha_{(s,\mathbb{T})} \wedge \bigvee_{s \Rightarrow t} (\alpha_{(t,\psi_1)} - 1)$ .
7.  $\psi = EF\psi_1$ : we break this case down into several possibilities. Suppose that  $(s, n) \in M(EF\psi_1)$ . Then there exist  $t \in S$  and  $n' \in \mathbb{N}$  such that  $s \Rightarrow^{n'-n} t$  and  $(t, n') \in M(\psi_1)$ .

It follows using Lemma 4 that there exists  $a + b\mathbb{N} \in R(s, t)$  such that  $n' - n \in a + b\mathbb{N}$ . If  $b = 0$ , then plainly we have  $n \in \alpha_{(t,\psi_1)} - a$ . Consider the case where  $b \neq 0$ , so  $n' - n \equiv a \pmod{b}$ . Moreover, inductively, there exist, by Lemma 2, numbers  $T$  and  $L$  such that  $n_1, n_2 > T$  and  $n_1 \equiv n_2 \pmod{L}$  implies  $n_1 \in \llbracket \alpha_{(t,\psi_1)} \rrbracket$  iff  $n_2 \in \llbracket \alpha_{(t,\psi_1)} \rrbracket$ . We consider the different possible values of  $c$ , the residue of  $n' \pmod{b}$ . There are two possibilities for a given  $c$ : either there exist an infinite number of  $n' \equiv c \pmod{b}$  such that  $(n', t) \in M(\psi_1)$ , or there exist a finite number such values.

We first note that in the case of an infinite number of such values, there must exist such a value in the range  $T < n' \leq T + L$ , by the periodicity condition on  $\alpha_{(t,\psi_1)}$ . Writing  $\beta(t, \psi_1) = \alpha_{(t,\psi_1)} \wedge (\equiv c \pmod{\gcd(b, L)})$ , and  $\inf(t, \psi_1, b, c) = T < \max(\beta(t, \psi_1), t + L)$ , we find that  $\inf(t, \psi_1, b, c)$  expresses that there exist an infinite number of values of  $n' \equiv c \pmod{b}$  such that  $(n', t) \in M(\psi_1)$ . (Note that if  $(m, t) \in M(\psi_1)$  and  $m \equiv c \pmod{\gcd(b, L)}$ , then by the Chinese Remainder theorem, the simultaneous equation  $x \equiv c \pmod{b}$  and  $x \equiv m \pmod{L}$  has a solution, which may be taken to be greater than  $T$ , so that by periodicity of  $\alpha_{(t,\psi_1)}$  we have  $(x, t) \in M(\psi_1)$  and  $x \equiv c \pmod{b}$ .) Thus, we may represent the case of  $n$  with an infinite number of  $n' \equiv c \pmod{b}$  using the formula

$$\chi_{\infty}(t, \psi_1, b, c) = \inf(t, \psi_1, b, c) \wedge (\equiv c - a \pmod{b}).$$

In the case of a finite number of such values, note that if  $\llbracket \inf(t, \psi_1, b, c) \rrbracket = \emptyset$  then for all  $n' \in \llbracket \alpha_{(t,\psi_1)} \rrbracket$  with  $n' \equiv c \pmod{b}$  we have  $n' < T$ . (If there exists such a value  $n'$  larger than  $T$  then by  $L$ -periodicity we can find one in the range  $[T, T + L]$  with  $n' \equiv c \pmod{\gcd(b, L)}$ .) Thus, in this case we can express the possible values of  $n$  by the expression

$$\chi_{<\infty}(t, \psi_1, b, c) = \neg \inf(t, \psi_1, b, c) \wedge \leq \max((\alpha_{(t,\psi_1)} \equiv c \pmod{b}), t).$$

Putting the pieces together, we may define  $\alpha_{(s,\psi)}$  to be

$$\bigvee_{t \in S} \left( \bigvee_{a+b\mathbb{N} \in R(s,t), b=0} (\alpha_{(t,\psi_1)} - a) \vee \bigvee_{a+b\mathbb{N} \in R(s,t), 0 \leq c < b} \chi_{\infty}(t, \psi_1, b, c) \vee \chi_{<\infty}(t, \psi_1, b, c) \right).$$

Combining this result with Lemma 3, we obtain the upper bound of  $P^{NP}$  for the combined complexity of model checking  $\text{CTL}^-K_n^{\text{clk}}$ .

By way of lower bound, we show that the combined complexity of model checking  $\text{CTL}^-K_n^{\text{clk}}$  is hard for  $P^{NP[\log]}$ , i.e., the complexity class corresponding to PTIME computations with a logarithmic number of queries to an NP oracle. (The same lower bound also holds for model complexity, by a slightly more elaborate proof that we leave for a longer version of this paper.) This is equivalent to the class of PTIME computations with *independent* queries to an NP oracle, i.e., the construction of an oracle query may not depend on the outcome of any other oracle query. The following problem ODD-SAT is known to be complete for this class: given boolean formulas  $\phi_1, \dots, \phi_n$ , determine if there exists an odd  $i$  such that  $\phi_1, \dots, \phi_i$  are satisfiable and  $\phi_{i+1}, \dots, \phi_n$  are not satisfiable.

To encode this as a model checking problem, let  $p_1 \dots p_k$  be the first  $k$  primes, and let  $M_k$  be the model with the transition relation composed of a set of independent cycles  $s_{i,1} \Rightarrow s_{i,2} \dots \Rightarrow s_{i,p_i} \Rightarrow s_{i,1}$  with the  $i$ -th such cycle of length  $p_i$ . Let the proposition  $q_i$  hold only at state  $s_{i,1}$ , for  $i = 1..k$ . Take the set of initial states to be  $I = \{s_{1,1}, \dots, s_{k,1}\}$ . Suppose that there is a single agent 1 that is not able to distinguish any of these states, i.e.  $O_1(s) = O_1(t)$  for all states  $s, t$ . The size of  $M_k$  is polynomial in  $k$ , since this also holds for the size of the  $k$ -th prime.

Note that the sets  $P_n = \{t \in S \mid \exists s \in I (s \Rightarrow^n t)\}$  pass through all combinations of sets of size  $k$ , with one element selected from each cycle. Thus, we may use the formulas  $\neg K_1 \neg q_1, \dots, \neg K_1 \neg q_k$ , whose truth value at a state  $s$  at time  $n$  depends only on  $P_n$  and not  $s$ , to represent an assignment to  $k$  boolean variables, and satisfiability of the boolean formula  $\phi(x_1, \dots, x_k)$  can be represented by the formula  $\phi^* = EF(\phi(\neg K_1 \neg q_1, \dots, \neg K_1 \neg q_k))$ . Finally, we encode the problem ODD-SAT by the formula

$$\bigvee_{i=1..n, i \text{ odd}} \bigwedge_{j=1..i} \phi_j^* \wedge \bigwedge_{j=i+1..n} \neg \phi_j^*.$$

This lower bound does not quite match the upper bound. We do not know at present whether either can be improved.

## 4 Related Work

The model checking complexities for temporal logics are collected in Table 2, which is taken from [13].

| Logic            | Combined Complexity | Model Complexity   | Formula Complexity |
|------------------|---------------------|--------------------|--------------------|
| LTL              | PSPACE-complete     | NLOGSPACE-complete | PSPACE-complete    |
| CTL*             | PSPACE-complete     | NLOGSPACE-complete | PSPACE-complete    |
| CTL              | PTIME-complete      | NLOGSPACE-complete | LOGSPACE           |
| CTL <sup>-</sup> | PTIME-complete      | NLOGSPACE-complete | LOGSPACE           |

**Table 2.** Complexity Results for Temporal Logics

When combined with epistemic logic under the observational semantics, the model checking complexities do not increase, e.g., for  $\text{CTLK}_n^{\text{obs}}$ , PTIME-completeness is shown in [6] and [10], for upper bound and lower bound, respectively.

When combined with epistemic logic under the perfect recall semantics, many subtleties are introduced. Linear time temporal epistemic logics on synchronous systems or asynchronous systems are analyzed in [19] and [20], and [16, 15, 14] explore branching time temporal epistemic logics.

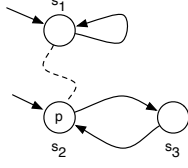


Figure 1. A model

For epistemic logic under clock semantics, [5] provides complexity results for its combination with LTL. The present paper complements this work by considering instead several branching time logics. A bounded model checking (BMC) algorithm for the universal fragment of  $CTL^*K_n^{clk}$  logic, abbreviated as  $ACTL^*K_n^{clk}$ , is presented in [11, 12]. However, it appears to be incorrect. A flaw of their algorithm occurs on formulas of the form  $AFK_i\phi$ , which in negated dual form  $EG\neg K_i\neg\psi$  requires finding a witness run on which  $\neg K_i\neg\psi$  holds at all times. For bounded model checking temporal logics, runs can be represented as lassos  $s_0 \Rightarrow s_1 \Rightarrow \dots \Rightarrow s_k \Rightarrow s_l$  where  $0 \leq l \leq k$  and  $k \leq |S|$ . To satisfy the formula  $EG\neg K_i\neg\psi$  would then require that we satisfy  $\neg K_i\neg\psi$  at each point in the run. For the observational semantics for knowledge, this can be done by finding for each time  $m$  with  $0 \leq m \leq k$  another lasso  $t_0 \Rightarrow t_1 \Rightarrow \dots \Rightarrow t_{k'} \Rightarrow t_{l'}$  with  $s_m \sim_i t_{m'}$  for some  $m'$  with  $\psi$  holding at the state  $t_{m'}$ . However, the clock semantics requires that the states  $s_m$  and  $t_{m'}$  occur at the same time. The approach to this in [11, 12] is to check that the witness state  $t_{m'}$  can occur at the same time as the state  $s_m$ .

However, this is not sufficient, rather, we require such a witness for each time  $n$  such that  $s_k$  can occur at time  $n$  in the original lasso. An example is shown in Figure 1. The model  $M = (S, I, \Rightarrow, \sim_a, \pi)$ , where  $S = \{s_1, s_2, s_3\}$ ,  $I = \{s_1, s_2\}$ ,  $\Rightarrow$  is shown in Figure 1,  $\sim_a$  is the least equivalence relation with  $s_1 \sim_a s_2$  and  $\pi(s_2) = \{p\}$ ,  $\pi(s_1) = \pi(s_3) = \emptyset$ . For the  $ACTL^*K_n^{clk}$  formula  $\phi = AFK_a\neg p$ , the algorithm of [11, 12] would conclude that  $M \not\models \phi$  by finding a ‘counterexample’ lasso of length 1, where the dual  $\neg\phi \equiv EG\neg K_a\neg p$  is resolved on loop  $s_1 \Rightarrow s_1$ , and the fact that  $\neg K_a\neg p$  must hold at  $s_1$  is witnessed by the fact that  $p$  holds at state  $s_2 \sim_a s_1$  on the lasso  $s_2 \Rightarrow s_3 \Rightarrow s_2$ . However, though  $\neg K_a\neg p$  is satisfiable on the lasso  $s_1 \Rightarrow s_1$  at time 0, it is not satisfiable at time 1 since state  $s_2$  is not possible at time 1.

By comparison, the algorithms discussed above for  $CTLK_n^{clk}$  and  $LTLK_n^{clk}$  use set-level loops. More specifically, the state-level lasso  $s_1 \Rightarrow s_1$  should be lifted to a set-level lasso  $(s_1, \{s_1, s_2\}) \Rightarrow (s_1, \{s_1, s_3\}) \Rightarrow (s_1, \{s_1, s_2\})$ , where the second element in a pair is the set of possible states at a given time. Now  $\neg K_a\neg p$  is not satisfiable on  $(s_1, \{s_1, s_3\})$  because state  $s_2$  is not possible at that time. Therefore, the loop  $s_1 \Rightarrow s_1$  does not provide a valid counterexample for formula  $\phi$ . Indeed, it can be easily checked that  $M \models \phi$ .

While we believe that lasso-structures can be used to provide counter-examples for bounded model checking of  $ACTL^*K_n^{clk}$ , the fact that the model complexity of  $CTLK_n^{clk}$  is hard for the polynomial hierarchy (compared with the NLOGSPACE-complete model complexities of LTL and CTL) strongly suggests that the claim in [11, 12] that lassos of length at most  $|M|$  suffice cannot be upheld. Rather, the apparent necessity of set-level lassos suggests that the bound should be of the order of  $2^{|M|}$ . We leave a more detailed investigation of this issue for further work.

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