ATL with Strategy Contexts: Expressiveness and Model Checking

Arnaud Da Costa¹, François Laroussinie², and Nicolas Markey¹

¹ Lab. Spécification & Vérification, ENS Cachan & CNRS, France
² LIAFA, Univ. Paris Diderot - Paris 7 & CNRS, France

Abstract

We study the alternating-time temporal logics ATL and ATL* extended with strategy contexts: these make agents commit to their strategies during the evaluation of formulas, contrary to plain ATL and ATL* where strategy quantifiers reset previously selected strategies.

We illustrate the important expressive power of strategy contexts by proving that they make the extended logics, namely ATLsc and ATL* sc, equally expressive: any formula in ATL* sc can be translated into an equivalent, linear-size ATLsc formula. Despite the high expressiveness of these logics, we prove that their model-checking problems remain decidable by designing a tree-automata-based algorithm for model-checking ATL* sc on the full class of n-player concurrent game structures.

1 Introduction

Temporal logics and model checking. Thirty years ago, temporal logics (LTL, CTL) have been proposed for specifying properties of reactive systems, with the aim of automatically checking that those properties hold for these systems [18, 10, 19]. This model-checking approach to formal verification has been widely studied, with powerful algorithms and implementations, and successfully applied in many situations.

Alternating-time temporal logic (ATL). In the last ten years, temporal logics have been extended with the ability of specifying controllability properties of multi-agent systems: the evolution of a multi-agent system depends on the concurrent actions of several agents, and ATL extends CTL with strategy quantifiers [4]: it can express properties such as agent A has a strategy to keep the system in a set of safe states, whatever the other agents do.

Nesting strategy quantifiers. Assume that, in the formula above, “safe states” are those from which agent B has a strategy to reach her goal state qB infinitely often, and consider the system depicted on Fig. 1, where the circled states are controlled by player A (meaning that Player A selects the transition to be fired from those state) and the square state is controlled by player B. It is easily seen that this game contains no “safe state”: after each visit to qB, Player A can decide to take the system to the rightmost state, from which qB is not reachable. It follows that Player A has no strategy to keep the system in safe states.

Now, assume that Player A commits to always select the transition to the left, when the system is in the initial (double-circled) state. Then under this strategy, it suffices for Player B to always go to qB when the system is in the square state in order to achieve her goal of visiting qB infinitely often. The difference with the previous case is that here, Player B takes advantage of Player A’s strategy in order to achieve her goal.
Both interpretations of our original property can make sense, depending on the context. However, the original semantics of ATL cannot capture the second interpretation: strategy quantifications in ATL “reset” previous strategies. While this is very convenient algorithmically (and makes ATL model-checking polynomial-time for some game models), it prevents ATL from expressing many interesting properties of games (especially non-zero-sum games).

In [7], we introduced an alternative semantics for ATL, where strategy quantifiers store strategies in a context. Those strategies then apply for evaluating the whole subformula, until they are explicitly removed from the context or replaced with a new strategy. We demonstrated the high expressiveness of this new semantics by showing that it can express important requirements, e.g., existence of equilibria or dominating strategies.

**Our contribution.** This work is a continuation of [7]. Our contribution in this paper is twofold: on the one hand, we prove that $\text{ATL}^\tl$ is not more expressive than $\text{ATL}^\tl$: this is a theoretical argument witnessing the expressive power of strategy contexts; it complements the more practical arguments presented in [7]. On the other hand, we develop an algorithm for $\text{ATL}^\tl$ model-checking, based on alternating tree automata. Our algorithm uses a novel encoding of strategies into the execution tree of the underlying concurrent game structures. This way, it is valid for the whole class of concurrent game structures and without restrictions on strategies, contrary to previously existing algorithms on related extensions of ATL.

**Related work.** In the last three years, several approaches have been proposed to increase the expressiveness of ATL and ATL$^\tl$.

- **Strategy logic** [8, 9] extends LTL with first-order quantification over strategies. This allows for very expressive constructs: for instance, the property above would be written as $\exists \sigma_A. [G (\exists \sigma_B. (G F q_B)(\sigma_A, \sigma_B))] (\sigma_A)$. This logic was only studied on two-player turn-based games in [8, 9], where a non-elementary algorithm is given. The algorithm we propose in this paper could be adapted to handle strategy logic in multi-player concurrent games.

- **QD$\mu$** [17] is a second-order extension of the propositional $\mu$-calculus augmented with decision modalities. In terms of expressiveness, fixpoints allow for richer constructs than CTL- or LTL-based approaches. Again, model-checking has been proved to be decidable, but only over the class of alternating transition systems (as defined in [3]).

- **Stochastic game logic** [6] is an extension of ATL similar to ours, but in the stochastic case. It is proved undecidable in the general case, and decidable when strategy quantification is restricted to memoryless (randomized or deterministic) strategies.

- several other semantics of ATL, related to ours, are discussed in [1, 2]. A $\Delta^P$-algorithm is proposed there for a subclass of our logic (where strategies stored in the context are irrevocable and cannot be overwritten), but no proof of correctness is given. In [20], an NP algorithm is proposed for the same subclass, but where strategy quantification is restricted to memoryless strategies.

By lack of space, some proofs are omitted in this paper, but they are detailed in [11].

## 2 ATL with Strategy Contexts

### 2.1 Concurrent game structures.

Concurrent game structures [4] are a multi-player extension of classical Kripke structures. Their definition is as follows:
Definition 1. A Concurrent Game Structure (CGS for short) \( C \) is an 7-tuple \( \langle \text{Loc}, \text{Lab}, \delta, \text{Agt}, \mathcal{M}, \text{Mov}, \text{Edg} \rangle \) where:

- \( (\text{Loc}, \text{Lab}, \delta) \) is a finite Kripke structure, where \( \text{Loc} \) is the set of locations, \( \text{Lab} : \text{Loc} \to 2^{\text{AP}} \) is a labelling function, and \( \delta \subseteq \text{Loc} \times \text{Loc} \) is the set of transitions;
- \( \text{Agt} = \{ A_1, \ldots, A_p \} \) is a finite set of agents (or players);
- \( \mathcal{M} \) is a finite, non-empty set of moves;
- \( \text{Mov} : \text{Loc} \times \text{Agt} \to \mathcal{P}(\mathcal{M}) \setminus \{ \emptyset \} \) defines the (finite) set of possible moves of each agent in each location.
- \( \text{Edg} : \text{Loc} \times \mathcal{M}^{\text{Agt}} \to \delta \) is a transition table; with each location \( \ell \) and each set of moves of the agents, it associates the resulting transition, which is required to depart from \( \ell \).

The size \( |C| \) of a CGS \( C \) is \( |\text{Loc}| + |\text{Edg}| \), where \( |\text{Edg}| \) is the size of the transition table\(^1\).

The intended behaviour of a CGS is as follows [4]: in a location \( \ell \), each player \( A_i \) in \( \text{Agt} \) chooses one among her possible moves \( m_i \) in \( \text{Mov}(\ell, A_i) \); the next transition to be fired is given by \( \text{Edg}(\ell, (m_1, \ldots, m_p)) \). We write \( \text{Next}(\ell) \) for the set of all transitions corresponding to possible moves from \( \ell \), and \( \text{Next}(\ell, A_j, m_j) \), with \( m_j \in \text{Mov}(\ell, A_j) \), for the restriction of \( \text{Next}(\ell) \) to possible transitions from \( \ell \) when player \( A_j \) plays the move \( m_j \). We extend \( \text{Mov} \) and \( \text{Next} \) to coalitions (i.e., sets of agents) in the natural way:

- Given \( A \subseteq \text{Agt} \) and \( \ell \in \text{Loc} \), \( \text{Mov}(\ell, A) \) denotes the set of possible moves for coalition \( A \) from \( \ell \). Those moves are composed of one single move per agent of the coalition, i.e., \( m = (m_a)_{a \in A} \).
- Given \( m = (m_a)_{a \in A} \in \text{Mov}(\ell, A) \), we let \( \text{Next}(\ell, A, m) \) denote the restriction of \( \text{Next}(\ell) \) to locations reachable from \( \ell \) when every player \( A_j \in A \) makes the move \( m_{A_j} \).

A (finite or infinite) path of \( C \) is a sequence \( \rho = \ell_0 \ell_1 \ldots \) of locations such that for any \( i \), \( \ell_{i+1} \in \text{Next}(\ell_i) \). Finite paths are also called history. The length of a history \( \rho = \ell_0 \ell_1 \ldots \ell_n \) is \( n \). We write \( \rho^{i \rightarrow j} \) for the part of \( \rho \) between \( \ell_i \) and \( \ell_j \) (inclusive). In particular, \( \rho^{i \rightarrow j} \) is empty iff \( j < i \). We simply write \( \rho^i \) for \( \rho^{i \rightarrow i} \), denoting the i + 1-st location \( \ell_i \) of \( \rho \). We also define \( \text{first}(\rho) = \rho^0 \), and, if \( \rho \) has finite length \( n \), \( \text{last}(\rho) = \rho^n \). Given a history \( \pi \) of length \( n \) and a path \( \rho \) s.t. \( \text{last}(\pi) = \text{first}(\rho) \), the concatenation of \( \pi \) and \( \rho \) is the path \( \tau = \pi \cdot \rho \) s.t. \( \tau_0^{n \rightarrow \infty} = \pi \) and \( \tau^{n \rightarrow \infty} = \rho \) (notice that the last location of \( \pi \) and the first location of \( \rho \) are “merged”).

A strategy for a player \( A_i \in \text{Agt} \) is a function \( f_i \) that maps any history to a possible move for \( A_i \), i.e., satisfying \( f_i(\ell_0 \ldots \ell_m) \in \text{Mov}(\ell_m, A_i) \). A strategy for a coalition \( A \) of agents is a mapping assigning a strategy to each agent in the coalition. The set of strategies for \( A \) is denoted \( \text{Strat}(A) \). The domain of \( F_A \in \text{Strat}(A) \) (denoted \( \text{dom}(F_A) \)) is \( A \). Given a coalition \( B \), the strategy \( (F_A)|_B \) (resp. \( (F_A)\setminus_B \)) denotes the restriction of \( F_A \) to the coalition \( A \cap B \) (resp. \( A \setminus B \)).

Let \( \rho \) be a history of length \( n \). A strategy \( F_A = (f_j)_{A_j \in A} \) for some coalition \( A \) induces a set of paths from \( \rho \), called the outcomes of \( F_A \) after (or from) \( \rho \), and denoted \( \text{Out}(\rho, F_A) \): a path \( \pi = \rho \cdot \ell_1 \ell_2 \ldots \) is in \( \text{Out}(\rho, F_A) \) iff, writing \( \ell_0 = \text{last}(\rho) \), for all \( i \geq 0 \) there exists a set of moves \( (m_k^i)_{A_k \in \text{Agt}} \) such that \( m_k^i \in \text{Mov}(\ell_i, A_k) \) for all \( A_k \in \text{Agt} \), \( m_k^i = f_{A_k}(\pi_0^{i+n+k}) \) if \( A_k \in A \), and \( \ell_{i+1} \in \text{Next}(\ell_i, A_k, (m_k^i)_{A_k \in \text{Agt}}) \). We write \( \text{Out}_\infty(\rho, F_A) \) for the set of infinite outcomes of \( F_A \) after \( \rho \). Note that \( \text{Out}(\rho, F_A) \subseteq \text{Out}(\rho, (F_A)|_B) \) for any two coalitions \( A \) and \( B \), and that \( \text{Out}(\rho, F_A) \) represents the set of all paths starting with \( \rho \).

\(^1\) Our results would still hold (with the same complexity) if we consider symbolic CGSs [13], where the transition table is encoded succinctly as boolean formulas.
It is also possible to combine two strategies $F \in \text{Strat}(A)$ and $F' \in \text{Strat}(B)$, resulting in a strategy $F \circ F' \in \text{Strat}(A \cup B)$ defined as follows: $(F \circ F')|_{A_j}(\rho)$ (resp. $F'|_{A_j}(\rho)$) if $A_j \in A$ (resp. $A_j \in B \setminus A$).

Finally, given a strategy $F$ and a history $\rho$, we define the strategy $F^{\rho}$ corresponding to the behaviour of $F$ after prefix $\rho$: it is defined, for any history $\pi$ with $\text{last}(\rho) = \text{first}(\pi)$, as $F^{\rho}(\pi) = F(\rho \cdot \pi)$.

### 2.2 Alternating-time temporal logics.

The logics ATL and ATL* have been defined in [4] as extensions of CTL and CTL* with strategy quantification. Following [7], we further extend them with strategy contexts:

**Definition 2.** The syntax of $\text{ATL}_{sc}^*$ is defined by the following grammar:

$$\text{ATL}_{sc}^* \ni \varphi, \psi ::= p \mid \neg \varphi \mid \varphi \lor \psi \mid \langle A \rangle \varphi_p \mid \langle \cdot \rangle \varphi_s$$

$$\varphi_p, \psi_p ::= \varphi \mid \neg \varphi_p \mid \varphi_p \lor \psi_p \mid X \varphi_p \mid \varphi_p U \psi_p$$

with $p \in \text{AP}$ and $A \subseteq \text{Agt}$. Formulas defined as $\varphi_s$ are called state-formulas, while $\varphi_p$ defines path-formulas. The logic $\text{ATL}_{sc}^*$ is obtained by restricting the grammar of $\text{ATL}_{sc}^*$ path-formulas as follows:

$$\varphi_p, \psi_p ::= \neg \varphi_p \mid X \varphi_s \mid \varphi_s U \psi_s.$$  

That a formula $\varphi$ in $\text{ATL}_{sc}^*$ (or $\text{ATL}_{sc}$) holds (initially) along a computation $\rho$ of a CGS $C$ under a strategy context $F$ (i.e., a preselected strategy for some of the players, hence belonging to some $\text{Strat}(A)$ for a coalition $A$), denoted $C, \rho \models_F \varphi$, is defined as follows:

$$C, \rho \models_F p \iff p \in \text{Lab} (\text{first}(\rho))$$

$$C, \rho \models_F \neg \varphi \iff C, \rho \not\models_F \varphi$$

$$C, \rho \models_F \varphi \lor \psi \iff C, \rho \models_F \varphi \text{ or } C, \rho \models_F \psi$$

$$C, \rho \models_F \langle A \rangle \varphi_p \iff \exists F_A \in \text{Strat}(A). \forall p' \in \text{Out}^{\infty} (\text{first}(\rho), F_A \circ F). C, \rho' \models_{F_A \circ F} \varphi_p$$

$$C, \rho \models_F \langle \cdot \rangle \varphi_s \iff C, \rho \models_{F \circ A} \varphi_s$$

$$C, \rho \models_F X \varphi_p \iff C, \rho^{i \rightarrow \infty} \models_{F_p \circ \rho^{i \rightarrow \infty}} \varphi_p$$

$$C, \rho \models_F \varphi_p U \psi_p \iff \exists i \geq 0. C, \rho^{i \rightarrow \infty} \models_{F_p \circ \rho^{i \rightarrow \infty}} \psi_p \text{ and } \forall 0 \leq j < i. C, \rho^{j \rightarrow \infty} \models_{F_p \circ \rho^{j \rightarrow \infty}} \varphi_p$$

We define the following shorthands, which will be useful in the sequel: $\top \overset{\text{def}}{=} p \lor \neg p$, $\bot \overset{\text{def}}{=} \neg \top$, $F \varphi \overset{\text{def}}{=} T U \varphi$, $G \varphi \overset{\text{def}}{=} \neg F \neg \varphi$, $\langle A \rangle \varphi_s \overset{\text{def}}{=} \langle A \rangle (\perp U \varphi_s)$ and $\langle A \rangle \varphi \overset{\text{def}}{=} \langle \text{Agt}(\langle A \rangle \varphi) \rangle$.

**Example 3** (see [7] for more examples). We illustrate the usefulness of strategy contexts with some examples. First, the last shorthand $\langle A \rangle$ is the classical ATL* strategy quantifier (where each quantification resets the context), so that $\text{ATL}_{sc}$ and $\text{ATL}_{sc}^*$ encompass ATL and ATL*, respectively.

$\text{ATL}_{sc}$ can also express qualitative equilibria properties, for instance Nash equilibria. Given the (non-zero-sum) objectives $\Phi_1$ and $\Phi_2$ of players 1 and 2, Nash equilibria are strategy profiles where none of the player can unilaterally improve her payoff. In other terms, if the Player-1 strategy in the context is not winning against the Player-2 strategy, then there is no Player-1 winning strategy against this particular strategy of Player 2 (and symmetrically). Thus, the existence of a Nash equilibrium can be expressed as

$$\langle A_1, A_2 \rangle \left[ (\langle A_1 \rangle \Phi_1 \rightarrow \Phi_1) \land (\langle A_2 \rangle \Phi_2 \rightarrow \Phi_2) \right]$$
As another example, we mention the interaction between a server \( S \) and different clients \( C_i \), where we may want to express that the server can be programmed in such a way that each client \( C_i \) has a strategy to have its request granted. This could be written as

\[
\langle \cdot \rangle S \left[ \bigwedge_i (\text{req}_i \rightarrow \langle A_i \rangle F \text{grant}_i) \right]
\]

As stated in Lemma 4, the truth value of a state formula \( \varphi_s \) depends only on the strategy context \( F \) and the first state of the computation \( \rho \) where it is interpreted (thus we may simply write \( C, \text{first}(\rho) \models_F \varphi_s \) when it raises no ambiguity).

\[\blacktriangleright\text{Lemma 4.}\]

Let \( C \) be a CGS, and \( F \in \text{Strat}(A) \) be a strategy context. For any state formula \( \varphi_s \), and for any two infinite paths \( \rho \) and \( \rho' \) with \( \text{first}(\rho) = \text{first}(\rho') \), it holds

\[
C, \rho \models_F \varphi_s \iff C, \rho' \models_F \varphi_s.
\]

\[\blacktriangleright\text{Proof.}\]

The proof is by induction on the structure of \( \varphi_s \): the result obviously holds for atomic propositions, and it is clearly preserved by boolean combinations and by the \( \langle A \rangle \) operator. Finally, if \( \varphi_s = \langle A \rangle \psi_s \), the result is immediate as the semantics only involves the first location of the path along which the formula is being evaluated.

\[\blacktriangleright\text{Remark.}\]

It must be noted that contrary to ATL, it is not possible to restrict to memoryless strategies (i.e., that only depend on the current state) for \( \text{ATL}_{sc} \) formulas. For example, the formula \( \langle \cdot \rangle A \left[ \langle \emptyset \rangle F P \land \langle \emptyset \rangle F P' \right] \) is equivalent in a standard Kripke structure (seen as a CGS with one single player \( A \)) to the CTL* formula \( E(\overline{F} P \land \overline{F} P') \) that may require strategies with memory. The next section provides more results on the extra expressiveness brought in by strategy contexts.

### 3 The expressive power of strategy contexts

As shown in [7], adding strategy contexts in formulas increases the expressive power of logics: \( \text{ATL}_{sc} \) (resp. \( \text{ATL}_{sc}^\star \)) is strictly more expressive than ATL (resp. ATL*). Game Logic (see [4]) can also be translated into \( \text{ATL}_{sc}^\star \) (while the converse is not true). In this section, we present some new results on the expressiveness of \( \text{ATL}_{sc} \).

#### 3.1 Alternating bisimulation.

Contrary to ATL, ATL*, GL or AMC, our logics are not alternating-bisimulation (see [5]) invariant, indeed we have: There exists two CGSs \( C \) and \( C' \), with an alternating-bisimulation linking two states \( \ell_0 \) of \( C \) and \( \ell'_0 \) in \( C' \), and an \( \text{ATL}_{sc} \) formula \( \varphi \) such that \( C, \ell_0 \models \varphi \) and \( C', \ell'_0 \not\models \varphi \).

#### 3.2 Relative expressiveness of \( \text{ATL}_{sc} \) and \( \text{ATL}_{sc}^\star \).

Surprisingly, strategy contexts bring \( \text{ATL}_{sc} \) to the same expressiveness as \( \text{ATL}_{sc}^\star \). This was already exemplified above, with the CTL* formula \( E(\overline{F} P \land \overline{F} P') \). We can extend this approach to any \( \text{ATL}_{sc}^\star \) formula: the idea is to

1. first use full strategy contexts (by adding universally quantified strategies) in order to be able to insert the \( \langle \emptyset \rangle \) modality before every temporal modality, and

2. ensure that for every nested strategy quantifier \( \langle A \rangle \), Coalition \( A \) cannot take advantage of the added strategies.
Given an ATL\textsuperscript{sc} formula $\Phi$ and a coalition $B$, we define $\hat{\Phi}^{\lbrack B\rbrack}$ inductively as follows:

\[
\begin{align*}
\hat{p}^{\lbrack B\rbrack} & \overset{\text{def}}{=} p \\
\neg \hat{\varphi}^{\lbrack B\rbrack} & \overset{\text{def}}{=} \neg \hat{\varphi}^{\lbrack B\rbrack} \\
\hat{\forall}^{\lbrack B\rbrack} \Phi & \overset{\text{def}}{=} \langle \varnothing \rangle \hat{\forall}^{\lbrack B\rbrack} \hat{\varphi}^{\lbrack B\rbrack} \\
\hat{\exists}^{\lbrack B\rbrack} \Phi & \overset{\text{def}}{=} \langle A \rangle \hat{\exists}^{\lbrack B\rbrack} \hat{\varphi}^{\lbrack B\rbrack} \\
\hat{\varphi}^{\lbrack A \cup B\rbrack} & \overset{\text{def}}{=} \langle A \rangle \hat{\exists}^{\lbrack B\rbrack} \hat{\varphi}^{\lbrack A \cup B\rbrack} \\
\hat{\varphi}^{\lbrack B \setminus A\rbrack} & \overset{\text{def}}{=} \langle A \rangle \hat{\exists}^{\lbrack B\rbrack} \hat{\varphi}^{\lbrack B \setminus A\rbrack}
\end{align*}
\]

Clearly, $\hat{\Phi}^{\lbrack B\rbrack}$ is an ATL\textsuperscript{sc} formula. The idea behind this translation is that a state-formula $\hat{\varphi}^{\lbrack A\rbrack}$ interpreted in a strategy context $F$ only depends on $F_{\mid A}$. We then have:

\begin{itemize}
\item \textbf{Lemma 5.} Let $C$ be a CGS, $l$ be one of its locations, and $F$ be a strategy context. Then for any ATL\textsuperscript{sc} formula $\varphi$, for any strategy context $G$ s.t. $\text{dom}(G) = \text{Agt} \setminus \text{dom}(F)$, and for any outcome $\pi \in \text{Out}^\infty(l, G \circ F)$, it holds: $C, \pi \models_F \varphi \iff C, \pi \models_{G \circ F} \hat{\varphi}^{\lbrack \text{dom}(F)\rbrack}$. Moreover, if $\varphi$ is a state-formula, this result extends to any strategy context $G$ s.t. $\text{dom}(G) \cap \text{dom}(F) = \emptyset$.
\end{itemize}

Since our transformation does not depend on the underlying CGS, we get:

\begin{itemize}
\item \textbf{Theorem 6.} Given a set of agents $\text{Agt}$, any ATL\textsuperscript{sc} formula $\varphi$ can be translated into an equivalent (under the empty context) ATL\textsuperscript{sc} formula $\hat{\varphi}$ for any CGS based on $\text{Agt}$.
\end{itemize}

Another consequence of the previous result is that any ATL\textsuperscript{*} state-formula $\varphi$ can be translated into the equivalent ATL\textsuperscript{sc} formula $\hat{\varphi}^{\lbrack \varnothing\rbrack}$ in polynomial time. Thus we have:

\begin{itemize}
\item \textbf{Corollary 7.} Model-checking ATL\textsuperscript{sc} is \textbackslash EXPTIME-hard.
\end{itemize}

\section{From ATL\textsuperscript{sc} to alternating tree automata}

The main result of this section is the following:

\begin{itemize}
\item \textbf{Theorem 8.} Model-checking ATL\textsuperscript{sc} formulas with at most $k$ nested strategy quantifiers can be achieved in $(k + 1)\text{EXPTIME}$. The program complexity (i.e., the complexity of model-checking a fixed ATL\textsuperscript{sc} formula) is in EXPTIME.
\end{itemize}

The proof mainly consists in building an alternating tree automaton from a formula and a CGS. Similar approaches have already been proposed for strategy logic \cite{9} or qD\mu \cite{17}, but they were only valid for subclasses of CGSs: strategy logic was only studied on turn-based games, while the algorithm for qD\mu was restricted to ATSs \cite{3}. In both cases, the important point is that strategies are directly encoded as trees, with as many successors of a node as the number of possible moves from the corresponding node. With this representation, it is required that two different successors of a node correspond to two different states (which is the case for ATSs, hence for turn-based games): if this is not the case, the tree automaton may accept strategies that do not only depend on the sequence of states visited in the history, but also on the sequence of moves proposed by the players. Our encoding is different: we work on the execution tree of the CGS under study, and label each node with possible moves of the players. We then have to focus on branches that correspond to outcomes of selected strategies, and check that they satisfy the requirement specified by the formula. Before presenting the detailed proof, we first introduce alternating tree automata and fix notations.

\subsection{Trees and alternating tree automata}

Let $\Sigma$ and $S$ be two finite sets. A $\Sigma$-labelled $S$-tree is a pair $T = (T, l)$, where...
\( T \subseteq S^* \) is a non-empty set of finite words on \( S \) satisfying the following constraints: for any non-empty word \( n = m \cdot s \) in \( T \) with \( m \in T \) and \( s \in S \), the word \( m \) is also in \( T \);

\( l: T \to \Sigma \) is a labeling function.

Given such a tree \( \mathcal{T} = (T, l) \) and a node \( n \in T \), the set of directions from \( n \) in \( T \) is the set \( \text{dir}_T(n) = \{ s \in S \mid n \cdot s \in T \} \). The set of successors of \( n \) in \( T \) is \( \text{succ}_T(n) = \{ n \cdot s \mid s \in \text{dir}_T(n) \} \). We use \( \mathcal{T}_n \) to denote the subtree rooted in \( n \). An \( S \)-tree is complete if \( T = S^* \), i.e., if \( \text{dir}_T(n) = S \) for all \( n \in T \). We may omit the subscript \( T \) when it is clear from the context.

The set of infinite paths of \( \mathcal{T} \) is the set \( \text{Path}_T = \{ s_0 \cdot s_1 \cdots \in S^\omega \mid \forall i \in \mathbb{N}, s_0 \cdot s_1 \cdots \cdot s_i \in T \} \). Given such an infinite path \( \pi = (s_i)_{i \in \mathbb{N}} \), we write \( l(\pi) \) for the infinite sequence \( (l(s_i))_{i \in \mathbb{N}} \in \Sigma^\omega \), and \( \text{Inf}(l(\pi)) \) for the set of letters in \( \Sigma \) that appear infinitely often along \( l(\pi) \).

Assume that \( \Sigma = \Sigma_1 \times \Sigma_2 \), and pick a \( \Sigma \)-labelled \( S \)-tree \( \mathcal{T} = (T, l) \). For all \( n \in T \), we write \( l(n) = (l_1(n), l_2(n)) \) with \( l_i(n) \in \Sigma_i \) for \( i \in \{1, 2\} \). Then for \( i \in \{1, 2\} \), the projection of \( \mathcal{T} \) on \( \Sigma_i \), denoted by \( \text{proj}_{\Sigma_i}(\mathcal{T}) \), is the \( \Sigma_i \)-labelled \( S \)-tree \((T, l_i)\). Two \( \Sigma \)-labelled \( S \)-trees are \( \Sigma_i \)-equivalent if their projections on \( \Sigma_i \) are equal. These notions naturally extend to more complex alphabets, of the form \( \prod_{i \in I} \Sigma_i \).

We now define alternating tree automata, which will be used in the proof. This requires the following definition: the set of positive boolean formulas over a finite set \( P \) of propositional variables is the set of formulas generated by: \( \text{PBF}(P) \) \( \triangleright= \{ \zeta \mid \zeta \wedge \zeta \mid \zeta \vee \zeta \mid T \mid l \} \), where \( p \) ranges over \( P \). That a valuation \( v: P \to \{ \top, \bot \} \) satisfies a formula in \( \text{PBF}(P) \) is defined in the natural way. We abusively say that a subset \( P' \) of \( P \) satisfies a formula \( \varphi \in \text{PBF}(P) \) iff the valuation \( \mathbf{1}_{P'} \) (mapping the elements of \( P' \) to \( \top \) and the elements of \( P \setminus P' \) to \( \bot \)) satisfies \( \varphi \). Since negation is not allowed, if \( P' \models \varphi \) and \( P' \subseteq P'' \), then also \( P'' \models \varphi \).

**Definition 9.** Let \( S \) and \( \Sigma \) be two finite sets. An alternating \( S \)-tree automaton on \( \Sigma \), or \( \langle S, \Sigma \rangle \)-ATA, is a 4-tuple \( \mathcal{A} = (Q, q_0, \tau, \text{Acc}) \) where \( Q \) is a finite set of states, \( q_0 \in Q \) is the initial state, \( \Sigma \) is a finite alphabet, \( \tau: Q \times \Sigma \to \text{PBF}(S \times Q) \) is the transition function, and \( \text{Acc}: Q^\omega \to \{ \top, \bot \} \) is the acceptance function.

A non-deterministic \( S \)-tree automaton on \( \Sigma \), or \( \langle S, \Sigma \rangle \)-NTA, is a \( \langle S, \Sigma \rangle \)-ATA in which conjunctions are not allowed for defining the transition function. The size of \( \mathcal{A} \), denoted by \( |\mathcal{A}| \), is the number of states in \( Q \).

Let \( \mathcal{A} = (Q, q_0, \tau, \text{Acc}) \) be an \( \langle S, \Sigma \rangle \)-ATA, and \( \mathcal{T} = (T, l) \) be a \( \Sigma \)-labelled \( S \)-tree. An execution tree of \( \mathcal{A} \) on \( \mathcal{T} \) is a \( T \times Q \)-labelled \( S \times Q \)-tree \( \mathcal{E} = (E, p) \) such that \( p(e) = (e, q_0) \), and for each node \( e \in E \) with \( p(e) = (t, q) \), the set \( \text{dir}_E(e) = \{ (s_0, q_0), (s_1, q_1), \ldots, (s_n, q_n) \} \subseteq S \times Q \) satisfies \( \tau(q, l(t)) \), and for all \( 0 \leq i \leq n \), the node \( e \cdot (s_i, q_i) \) is labelled with \( (t \cdot s_i, q_i) \). We write \( p(e \cdot (s_i, q_i)) = t \cdot s_i \) and \( p(e \cdot (s_i, q_i)) = q_i \) for the two components of the labelling function.

An execution tree is accepting if \( \text{Acc}(pQ(\pi)) = \top \) for any infinite path \( \pi \in (S \times Q)^\omega \) in \( \text{Path}_E \). A tree \( \mathcal{T} \) is accepted by \( \mathcal{A} \) iff there exists an accepting execution tree of \( \mathcal{A} \) on \( \mathcal{T} \).

4.2 Unwinding of a CGS

Let \( \mathcal{C} = (\mathsf{Loc}, \mathsf{Lab}, \delta, \mathsf{Agt}, \mathcal{M}, \mathsf{Mov}, \mathsf{Edg}) \) be an \( n \)-player CGS, where we assume w.l.o.g. that \( \delta = \mathsf{Loc} \times \mathsf{Loc} \) and \( \mathsf{Mov}(\ell, A_i) = \mathcal{M} \) for any state \( \ell \) and any player \( A_i \). Let \( \ell_0 \) be a state of \( \mathcal{C} \).
For each location $\ell \in \Loc$, we define $\Sigma(\ell) = \{\ell\} \times \{\Lab(\ell)\} \times \{\Edg(\ell)\}$, and $\Sigma^+(\ell) = \Sigma(\ell) \times (\mathcal{M} \cup \{\bot\})^{\abs{\Edg} + 2^{|p_o| + |p_p| + |p_r|}}$, where $\bot$ is a special symbol not in $\mathcal{M}$ and $p_o$, $p_p$ and $p_r$ are three fresh propositions not in $\AP$.

The unwinding of the $\Sigma^-$-labelled complete $\Loc$-tree $U = (\ell, v)$ where $U = \Loc^*$ and $v(u) \in \Sigma(\text{last}(\ell_0 \cdot u))$ for all $u \in U$. An extended unwinding of $\mathcal{C}$ from $\ell_0$ is a $\Sigma^+_\ell$-labelled complete $\Loc$-tree $U'$ such that $\proj_{\Sigma^-}(U') = U$. For each letter $\sigma$ of $\Sigma^+_\ell$, we write $\sigma_{\Loc}$, $\sigma_{\AP}$, $\sigma_{\Edg}$, $\sigma_{\str}$ and $\sigma_\ell$ for the five components, and extend this subscripting notation for the labelling functions of trees (written $l_{\Loc}$, $l_{\AP}$, $l_{\Edg}$, $l_{\str}$ and $l_\ell$).

In the sequel, we identify a node $u$ of $U$ (which is a finite word over $\Loc$) with the finite path $\ell_0 \cdot u$ of $\mathcal{C}$. Notice that this sequence of states of $\mathcal{C}$ may correspond to no real path of $\mathcal{C}$, in case it involves a transition that is not in the image of $\Edg$.

With $\mathcal{C}$ and $\ell_0$, we associate a $\langle \Loc, \Sigma^+_\ell \rangle$-APT $\mathcal{A}_{\mathcal{C}, \ell_0} = \langle \Loc, \ell_0, \Omega, \ell \rangle$ s.t. $\Loc = \{ \ell \mid \ell \in \Loc \}$, $\ell_0$ is the initial state, $\Omega$ constantly equals 0 (hence any valid execution tree is accepting), and given a state $\ell \in \Loc$ and a letter $\sigma \in \Sigma^+_\ell$, the transition function is defined as follows: if $\sigma \in \Sigma^+(\ell)$, we let $\tau(\ell, \sigma) = \bigwedge_{v \in \Loc} (\ell', \overline{\sigma})$, and otherwise, we let $\tau(\ell, \sigma) = \bot$.

**Lemma 10.** Let $\mathcal{C}$ be a CGS and $\ell_0$ be a state of $\mathcal{C}$. Let $\mathcal{T} = (T, l)$ be a $\Sigma^+_\ell$-labelled $\Loc$-tree. Then $\mathcal{A}_{\mathcal{C}, \ell_0}$ accepts $\mathcal{T}$ iff $\proj_{\Sigma^-}(\mathcal{T})$ is the unwinding of $\mathcal{C}$ from $\ell_0$.

In the sequel, we also use automaton $\mathcal{A}_\mathcal{C}$, which accepts the union of all $\mathcal{L}(\mathcal{A}_{\mathcal{C}, \ell_0})$ when $\ell_0$ ranges over $\Loc$. It is easy to come up with such an automaton, e.g. with Lemma 13 below.

### 4.3 Strategy quantification

Let $\mathcal{T} = (T, l)$ be a $\Sigma^+_\ell$-labeled complete $\Loc$-tree accepted by $\mathcal{A}_{\mathcal{C}, \ell_0}$. Such a tree defines partial strategies for each player: for $A \in \Agt$, and for each node $n \in T$, we define $\text{str}_A(\ell_0 \cdot n) = l_{\str}(n)(A) \in \mathcal{M} \cup \{\bot\}$. For $D \subseteq \Agt$, we write $\text{str}_D$ for $\langle \text{str}_A \rangle_{A \in D}$.

As a first step, for each $D \subseteq \Agt$, we build a $\langle \Loc, \Sigma^+_\ell \rangle$-APT $\mathcal{A}_{\text{str}}(D)$ which will ensure that for all $A \in D$, $\text{str}_A$ is really a strategy for player $A$, i.e., never returns $\bot$. This automaton has only one state $q_0$, with $\tau(q_0, \sigma) = \bigwedge_{\ell \in \Loc} (\ell, q_0)$ provided that $\sigma_{\str}(A) \neq \bot$ for all $A \in D$. Otherwise, $\tau(q_0, \sigma) = \bot$. Finally, $\mathcal{A}_{\text{str}}$ accepts all trees having a valid execution tree (i.e., $\Omega$ constantly equals 0). The following result is straightforward:

**Lemma 11.** Let $\mathcal{C}$ be a CGS, $\ell_0$ be a location of $\mathcal{C}$, and $D \subseteq \Agt$. Let $\mathcal{T} = (T, l)$ be a $\Sigma^+_\ell$-labelled complete $\Loc$-tree accepted by $\mathcal{A}_{\mathcal{C}, \ell_0}$. Then $\mathcal{T}$ is accepted by $\mathcal{A}_{\text{str}}(D)$ iff for each player $A \in D$, $\text{str}_A$ never returns $\bot$.

We now build an automaton for checking that proposition $p_o$ labels outcomes of $\mathcal{T}$. More precisely, let $D \subseteq \Agt$ be a set of players. The automaton $\mathcal{A}_{\text{out}}(D)$ will accept $\mathcal{T}$ iff $p_o$ labels exactly the outcomes of strategies $\text{str}_A$ for players $A \in D$. This is achieved by the following two-player automaton $\mathcal{A}_{\text{out}}(D) = (Q, q_0, \Omega, \ell, q_0, q_\mathcal{E}, \mathcal{E})$ of the initial state, $\Omega$ constantly equals 0, and the transition function is defined as follows: if $p_o \notin \sigma$, then $\tau(q_0, \sigma) = \bot$ and $\tau(q_\mathcal{E}, \sigma) = \bigwedge_{\ell \in \Loc} (\ell, q_\mathcal{E})$; otherwise, $\tau(q_\mathcal{E}, \sigma) = \bot$ and

$$\tau(q_\mathcal{E}, \sigma) = \bigwedge_{\ell \in \text{Next}(\sigma, D)} (\ell, q_\mathcal{E}) \land \bigwedge_{\ell \notin \text{Next}(\sigma, D)} (\ell, q_\mathcal{E})$$

Notice that $|\Sigma^-| = |\Loc|$ and $|\Sigma^+_\ell|$ is linear in the size of the input, as we assume an explicit representation of the $\Edg$ function [13].
The size of $A$ will use in our construction.

Next where

Lemma 13.

$\text{Lemma 13.}$

Sketch of proof.

Lemma 17.

$\text{Lemma 17.}$

4.5 Transforming an $\text{ATL}_{sc}$ formula into an alternating tree automaton

$\text{Lemma 17.}$

Let $C$ be a CGS with finite state space $\text{Loc}$. Let $\psi$ be an $\text{ATL}_{sc}$-formula, and $D \subseteq \text{Agt}$ be a coalition. We can build a $(\text{Loc}, \Sigma^+_C)$-APT $A_{\psi, D}$ s.t.

- for any $\Sigma^+_C$-labelled complete $\text{Loc}$-tree $T$ accepted by $A_C$ and by $A_{\text{strat}(D)}$, it holds

$$T \in \mathcal{L}(A_{\psi, D}) \iff C, l_{\text{Loc}}(e) \models_{\text{strat}_D} \psi;$$

- for any two $\Sigma^+_C$-labelled complete $\text{Loc}$-tree $T$ and $T'$ s.t. $\text{proj}_{\Sigma^+_C}(T) = \text{proj}_{\Sigma^+_C}(T')$, with $\Sigma^+_C = \Sigma_C \times (\mathcal{M} \cup \{\bot\})^{\text{Agt}}$, we have

$$T \in \mathcal{L}(A_{\psi, D}) \iff T' \in \mathcal{L}(A_{\psi, D}).$$

The size of $A_{\psi, D}$ is at most $d$-exponential, where $d$ is the number of (nested) strategy quantifiers in $\psi$. Its index is $d-1$-exponential.

Sketch of proof. The proof proceeds by induction on the structure of formula $\psi$. The case of atomic propositions is straightforward. Applying Lemma 13, we immediately get the result for the case when $\varphi$ is a boolean combination of subformulas.
We now sketch the proof for the case when \( \psi = \langle A \rangle X \varphi \). The case formulas \( \langle A \rangle \varphi_1 U \varphi_2 \) and \( \langle A \rangle \varphi_1 R \varphi_2 \) could be handled similarly. The idea of the construction is as follows: we use automaton \( A_{\text{out}}(D \cup A) \) to label outcomes with \( p_o \), \( A_{\varphi,D\cup\Delta A} \) to label nodes where \( \varphi \) holds, and build an intermediate automaton \( A_f \) to check that all the outcomes satisfy \( X \varphi \).

We then project out the strategy of coalition \( A \), which yields the automaton for \( \langle A \rangle X \varphi \).

Assume that we have already built the automaton \( A_{\varphi,D\cup\Delta A} \) (inductively). Applying Lemma 16 to \( A_{\varphi,D\cup\Delta A} \) with the extra proposition \( p_r \), we get an automaton \( B_{\varphi,D\cup\Delta A} \) such that, given a tree \( T = \langle T, l \rangle \) accepted by \( A_c \) and \( A_{\text{strat}}(D \cup A) \), it holds

\[
T \in \mathcal{L}(B_{\varphi,D\cup\Delta A}) \iff \forall n \in T. (p_r \in l(n) \Rightarrow C, l_{\text{Loc}}(n) \models_{\text{strat}_{D\cup\Delta A}} \varphi).
\]  

(1)

In order to check that all the outcome satisfy \( X \varphi \), we simply have to build an automaton \( A_f \) for checking the \( \text{CTL}^* \) property \( A(G p_o \rightarrow X p_r) \). We refer to [12] for this classical construction. This automaton \( A_f \) has the following property: for any \( \Sigma^c_+ \)-labelled Loc-tree \( T = \langle T, l \rangle \), we have

\[
T \in \mathcal{L}(A_f) \iff \exists \epsilon \in \mathcal{L}(A(G p_o \rightarrow X p_r))
\]  

(2)

Now, let \( H \) be the product of \( A_{\text{strat}}(A) \), \( A_{\text{out}}(D \cup A) \), \( A_f \) and \( B_{\varphi,D\cup\Delta A} \), and let \( T \) be a tree accepted by \( A_c \) and \( A_{\text{strat}}(D) \). If \( T \) is accepted by \( H \), then then \( D \cup A \subseteq \text{dom}(T) \) and all the outcomes of the strategy \( \text{strat}_{D\cup\Delta A} \) from \( l_{\text{Loc}}(\epsilon) \) satisfy \( X \varphi \).

The converse does not hold in general, but we prove a weaker form: from \( T = \langle T, l \rangle \), accepted by \( A_c \) and \( A_{\text{strat}}(D) \), and such that \( D \cup A \subseteq \text{dom}(T) \) and the outcomes of \( \text{strat}_{D\cup\Delta A} \) from \( l_{\text{Loc}}(\epsilon) \) satisfy \( X \varphi \), we build \( T' = \langle T, l' \rangle \) such that \( \text{proj}_{\Sigma^c_+}(T) = \text{proj}_{\Sigma^c_+}(T') \), and \( T' \) is accepted by \( H \). To do this, it suffices to modify the labelling of \( T \) with \( p_o \) and \( p_r \), in such a way that \( T' \) is accepted by \( A_{\text{out}}(D \cup A) \) and \( B_{\varphi,D\cup\Delta A} \). This ensures that \( C, l_{\text{Loc}}(\epsilon) \models_{\text{strat}_{\Delta A}^{T'}} (\varnothing) X \varphi \), and that \( T' \) is also accepted by \( A_f \). In the end, we have that for any tree \( T = \langle T, l \rangle \) accepted by \( A_c \) and \( A_{\text{strat}}(D) \),

\[
D \cup A \subseteq \text{dom}(T) \quad \text{and} \quad \exists \epsilon, T' \text{ s.t. } \text{proj}_{\Sigma^c_+}(T') = \text{proj}_{\Sigma^c_+}(T) \text{ and } T' \in \mathcal{L}(H).
\]  

(3)

Applying Lemma 14, we get a \( (\Sigma^c_+, \text{Loc}) \)-NPT \( N \) such that \( \mathcal{L}(N) = \mathcal{L}(H) \). We can then apply Lemma 15 for \( \Sigma^c_+ = (\Sigma_c \times (M \cup \{\bot\}))^{\text{p.o}} \times ((M \cup \{\bot\})^4 \times 2^{p_o-p_r-p_r}) \) on the NPT \( N \); the resulting \( (\Sigma^c_+, \text{Loc}) \)-NPT \( P \) accepts all trees \( T \) whose labelling on \( (M \cup \{\bot\})^4 \times 2^{p_o-p_r-p_r} \) can be modified in order to have the tree accepted by \( N \). Then \( P \) satisfies both properties of the Lemma: the second property directly follows from the use Lemma 15. For the first one, pick \( T = \langle T, l \rangle \) accepted by \( A_c \) and by \( A_{\text{strat}}(D) \). If \( T \) is accepted by \( P \), then from Lemma 15, there exists a tree \( T' = \langle T', l' \rangle \), with the same labelling as \( T \) on \( \Sigma_c \times (M \cup \{\bot\})^{\text{p.o}} \times (M \cup \{\bot\})^4 \), and accepted by \( N \). Since \( \mathcal{L}(N) = \mathcal{L}(H) \), and from (3), we get that \( D \cup A \subseteq \text{dom}(T') \) and \( C, l_{\text{Loc}}(\epsilon) \models_{\text{strat}_{\Delta A}^{T'}} (\varnothing) X \varphi \). Thus \( \text{strat}_{D\cup\Delta A}^{T'} \) is a strategy for coalition \( A \), and it witnesses the fact that \( C, l_{\text{Loc}}(\epsilon) \models_{\text{strat}_{D\cup\Delta A}^{T'}} (\langle A \rangle X \varphi) \), and we get the desired result since \( \text{strat}_{D\cup\Delta A}^{T} = \text{strat}_{D\cup\Delta A}^{T'} \). Conversely, if \( C, l_{\text{Loc}}(\epsilon) \models_{\text{strat}_{D\cup\Delta A}^{T'}} (\langle A \rangle X \varphi) \), then we can modify the labelling of \( T \) with a witnessing strategy for \( A \), obtaining a tree \( T' \) such that \( C, l_{\text{Loc}}(\epsilon) \models_{\text{strat}_{\Delta A}^{T'}} (\varnothing) X \varphi \). From (3), \( T' \) can in turn be modified into a tree \( T'' \), with \( \text{proj}_{\Sigma^c_+}(T'') = \text{proj}_{\Sigma^c_+}(T') \), in

\footnote{The “release” modality \( R \) is the dual of \( U \), defined by \( \varphi_1 R \varphi_2 \equiv \neg[(\neg \varphi_1) U (\neg \varphi_2)] \). Notice that \( X \) is self-dual as we only evaluate formulas along infinite outcomes.}
such a way that \( T'' \in \mathcal{L}(H) \). Finally, since the projections of \( T'' \) and \( T \) coincide on \((\Sigma_\mathcal{C} \times (\mathcal{M} \cup \{ \bot \})_{\text{agt} \langle - A \rangle} \), it holds that \( T \) is accepted by \( \mathcal{P} \). This concludes the proof for \( \langle A \rangle X \varphi \).

The proofs for the “until” and “release” modalities follow the same lines, using \( p_l \) and \( p_r \) as extra atomic propositions for the left- and right-hand subformulas, and modifying automaton \( A_f \) so that it accepts trees satisfying \( A(G p_0 \rightarrow p_l U p_r) \) and \( A(G p_0 \rightarrow p_l R p_r) \), respectively. Finally, when \( \psi = \psi_A(\varphi, \psi) \), we let \( A(\varphi, D) = A_{\varphi, D \langle - A \rangle} \), which is easily proved to satisfy both requirements.

Unless \( A = \emptyset \), the construction of the automaton for \( \langle A \rangle X \varphi \) (or \( \langle A \rangle \varphi_1 U \varphi_2 \) or \( \langle A \rangle \varphi_1 R \varphi_2 \)) involves an exponential blowup in the size and index of the automata for the subformulas, and the index is bilinear in the size and index of these automata. In the end, for a formula involving \( d \) nested non-empty strategy quantifiers, the automaton has size \( d \)-exponential and index \( d - 1 \)-exponential.

\[ \square \]

**Corollary 18.** Given an \( \text{ATL}_{sc} \) formula \( \varphi \), a CGS \( \mathcal{C} \) and a state \( \ell_0 \) of \( \mathcal{C} \), we can build an alternating parity tree automaton \( A \) s.t. \( \mathcal{L}(A) \neq \emptyset \) iff \( \mathcal{C}, \ell_0 \models_\emptyset \varphi \). Moreover, \( A \) has size \( d \)-exponential and index \( d - 1 \)-exponential, where \( d \) is the number of nested non-empty strategy quantifiers.

**Proof.** It suffices to take the product of the automaton \( A_{\varphi, \emptyset} \) (from Lemma 17) with \( A_{\mathcal{C}, \ell_0} \).

In case this \((\text{Loc}, \Sigma^+_{\mathcal{C}})\text{-APT}\) accepts a tree \( T \), Lemma 17 entails that \( \mathcal{C}, \ell_0 \models_\emptyset \varphi \). Conversely, if \( \mathcal{C}, \ell_0 \models \varphi \), then the extended unwinding tree \( T = \langle T, l \rangle \) of \( \mathcal{C} \) from \( \ell_0 \) in which \( l_{\text{at}}(n) = \bot \) for all \( n \in T \) is accepted by \( A_{\mathcal{C}, \ell_0} \) (and, trivially, by \( A_{\text{strat}(\emptyset)} \)), and from Lemma 17, it is also accepted by \( A_{\varphi, \emptyset} \).

\[ \square \]

**Proof of Theorem 8.** The first statement directly follows from the previous corollary, since emptiness of alternating parity tree automata \( A \) can be checked in time \( \exp(O(|A| \times \text{idr}(A))) \).

For the second statement, notice that the size and index of \( A_{\varphi, \emptyset} \) in the proof of Corollary 18 do not depend on the CGS \( \mathcal{C} \). Hence the automaton \( A \) of Corollary 18 has size linear in \( |\mathcal{C}| \), and can be computed in time exponential in \( |\mathcal{C}| \) (because \( A_{\text{out}}(D) \) requires the computation of \( \text{Next}(\sigma, D) \)). Non-emptiness is then checked in time exponential in \( |\mathcal{C}| \).

\[ \square \]

**Remark.** Our algorithm can easily be modified in order to handle \( \text{ATL}_{sc}^* \). One solution is to rely on Theorem 6, but our translation from \( \text{ATL}_{sc}^* \) to \( \text{ATL}_{sc} \) may double the number of nested non-empty strategy quantifiers. The algorithm would then be in \((2k + 1)-\text{EXPTIME}\), where \( k \) is the number of nested strategy quantifications. Another solution is to adapt our construction, by replacing each state-subformula with a fresh atomic proposition, and build the automaton \( A_f \) for a more complex CTL$^*$ formula. This results in a \((k + 1)-\text{EXPTIME}\) algorithm. In both cases, the program complexity is unchanged, in \( \text{EXPTIME} \).

Similarly, our algorithm could be modified to handle strategy logic [9]. One important difference is that strategy logic may require to store several strategies per player in the tree, while \( \text{ATL}_{sc} \) only stores one strategy per player. This would then be reflected in a modified version of the \( \text{Next} \) function we use when building \( A_{\text{out}}(D) \), where we should also indicate which strategies we use for which player.

**5 Conclusions**

Strategy contexts provide a very expressive extension of the semantics of \( \text{ATL} \), as we witnessed by the fact that \( \text{ATL}_{sc} \) and \( \text{ATL}_{sc}^* \) are equally expressive. We also designed a tree-automata-
based algorithm for model-checking both logics on the whole class of CGSs, based on a novel encoding of strategies as a tree.

Our algorithms involve a non-elementary blowup in the size of the formula, which we currently don’t know if it can be avoided. Trying to establish lower-bounds on the complexity of the problems is part of our future works. Regarding expressiveness, ATLsc can distinguish between alternating-bisimilar CGSs, and we are also looking for a behavioural equivalence that could characterize the distinguishing power of ATLsc.

References