# Decidability of Predicate Logics with Team Semantics 

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#### Abstract

We study the complexity of predicate logics based on team semantics. We show that the satisfiability problems of two-variable independence logic and inclusion logic are both NEXPTIMEcomplete. Furthermore, we show that the validity problem of two-variable dependence logic is undecidable, thereby solving an open problem from the team semantics literature. We also briefly analyse the complexity of the Bernays-Schönfinkel-Ramsey prefix classes of dependence logic.


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## 1 Introduction

The satisfiability problem of two-variable logic $\mathrm{FO}^{2}$ was shown to be NEXPTIME-complete in [9]. The extension of two-variable logic with counting quantifiers, FOC $^{2}$, was proved decidable in [10, 22], and it was subsequently shown to be NEXPTIME-complete in [23]. Research on extensions and variants of two-variable logic is currently very active. Recent research efforts have mainly concerned decidability and complexity issues in restriction to particular classes of structures and also questions related to different built-in features and operators that increase the expressivity of the base language. Recent articles in the field include for example [1], [4], [13], [16], [24], and several others.

In this article we study two-variable fragments of logics based on team semantics. Team semantics was originally conceived in [15] in the context of independence friendly (IF) logic [14]. In [25], Väänänen introduced dependence logic, which is a novel approach to IF logic based on new atomic formulas $=\left(x_{1}, \ldots x_{k}, y\right)$ stating that the interpretation of the variable $y$ is functionally determined by the interpretations of the variables $x_{1}, \ldots, x_{k}$.

After the introduction of dependence logic, research on logics based on team semantics has been active. Several different logics with different applications have been suggested. In particular, team semantics has proved to be a powerful framework for studying different kinds of dependency notions. Independence logic [11] is a variant of dependence logic that extends first-order logic by new atomic formulas $x_{1}, \ldots, x_{k} \perp y_{1}, \ldots, y_{l}$ with the intuitive meaning that the interpretations of the variables $x_{1}, \ldots, x_{k}$ are informationally independent of the interpretations of the variables $y_{1}, \ldots, y_{l}$. Inclusion logic [6] extends first-order logic by atomic formulas $x_{1}, \ldots, x_{k} \subseteq y_{1}, \ldots, y_{k}$, whose intuitive meaning is that tuples interpreting the
variables $x_{1}, \ldots, x_{k}$ are also tuples interpreting $y_{1}, \ldots, y_{k}$. Currently dependence, independence and inclusion logics are the three most important and most widely studied systems based on team semantics.

Both dependence logic and independence logic are equiexpressive with existential secondorder logic (see [25], [11]), and thereby capture NP. Curiously, inclusion logic is equiexpressive with greatest fixed point logic (see [7]), and thereby characterizes P on finite ordered models. While the descriptive complexity of most known logics based on team semantics is understood reasonably well, the complexity of related satisfiability problems has received somewhat less attention. The satisfiability problem of the two-variable fragment of dependence logic and IF-logic have been studied in [18]. It is shown that while the two-variable IF-logic is undecidable, the corresponding fragment of dependence logic is NEXPTIME-complete.

In this article we establish that the satisfiablity problems of the two-variable fragments of independence and inclusion logics are likewise NEXPTIME-complete. This result is established via proving a more general theorem that implies also a range of other decidability results for a variety of team-semantics-based logics with generalized dependency notions. Furthermore, we prove that the validity problem of two-variable dependence logic is undecidable; this result is the main result of the paper. The problem has been open for some time in the team semantics literature and has been explicitly posed in, e.g., [5], [18], [26], and elsewhere.

In addition to studying two-variable logics, we study the Bernays-Schönfinkel-Ramsey prefix class, i.e., sentences with the quantifier prefix $\exists^{*} \forall^{*}$. We show that - as in the case of ordinary first-order logic - the prefix class $\exists^{*} \forall^{*}$ of $\mathrm{FO}(\mathcal{A})$ is decidable for any uniformly polynomial time computable class $\mathcal{A}$ of generalized dependencies closed under substructures. We prove inclusion in 2NEXPTIME, and furthermore, for vocabularies of fixed arity, we show NEXPTIME-completeness. We also prove a partial converse of the result concerning $\operatorname{logics} \mathrm{FO}(\mathcal{A})$ with a decidable prefix class $\exists^{*} \forall^{*}$, see Theorem 22 .

## 2 Preliminaries

The domain of a structure $\mathfrak{A}$ is denoted by $A$. We assume that the reader is familiar with first-order logic FO. The extension of FO with counting quantifiers $\exists^{\geq i}$ is denoted by FOC. The two-variable fragments $\mathrm{FO}^{2}$ and $\mathrm{FOC}^{2}$ are the fragments of FO and FOC with formulas in which only the variables $x$ and $y$ appear. We let $\Sigma_{1}^{1}$ denote the fragment of formulas of second-order logic of the form $\exists X_{1} \ldots \exists X_{k} \varphi$, where $X_{1}, \ldots, X_{k}$ are relation symbols and $\varphi$ a first-order formula. $\Sigma_{1}^{1}\left(\mathrm{FOC}^{2}\right)$ is the extension of $\mathrm{FOC}^{2}$ consisting of formulas of the form $\exists X_{1} \ldots \exists X_{k} \chi$, where $X_{1}, \ldots, X_{k}$ are relation symbols and $\chi$ a formula of $\mathrm{FOC}^{2}$.

### 2.1 Logics based on team semantics

Let $\mathbb{Z}_{+}$denote the set of positive integers, and let $\operatorname{VAR}=\left\{v_{i} \mid i \in \mathbb{Z}_{+}\right\}$be the set of exactly all first-order variable symbols. We mainly use metavariables $x, y, z, x_{1}, x_{2}$, etc., in order to refer to variable symbols in VAR. We let $\bar{x}, \bar{y}, \bar{z}, \bar{x}_{1}, \bar{x}_{2}$, etc., denote finite nonempty tuples of variable symbols, i.e., tuples in $\operatorname{VAR}^{n}$ for some $n \in \mathbb{Z}_{+}$. When we study two-variable logics, we use the metavariables $x$ and $y$, and assume they denote distinct variables in VAR.

Let $D \subseteq$ VAR be a finite, possibly empty set. Let $\mathfrak{A}$ be a model. We do not allow for models to have an empty domain, so $A \neq \emptyset$. A function $s: D \rightarrow A$ is called an assignment with codomain $A$. If $\bar{x}=\left(x_{1}, \ldots, x_{n}\right)$, we denote $\left(s\left(x_{1}\right), \ldots, s\left(x_{n}\right)\right)$ by $s(\bar{x})$. We let $s[a / x]$ denote the variable assignment with the domain $D \cup\{x\}$ and codomain $A$ defined such
that $s[a / x](y)=a$ if $y=x$, and $s[a / x](y)=s(y)$ if $y \neq x$. Let $T \in \mathcal{P}(A)$, where $\mathcal{P}$ denotes the power set operator. We define $s[T / x]=\{s[a / x] \mid a \in T\}$.

Let $D \subseteq$ VAR be a finite, possibly empty set of first-order variable symbols. Let $X$ be a set of assignments $s: D \rightarrow A$. Such a set $X$ is a team with the domain $D$ and codomain $A$. Note that the empty set is a team, as is the set $\{\emptyset\}$ containing only the empty assignment. The team $\emptyset$ does not have a unique domain; any finite subset of VAR is a domain of $\emptyset$. The domain of the team $\{\emptyset\}$ is $\emptyset$.

Let $X$ be a team with the domain $D$ and codomain $A$. Let $T \subseteq A$. We define $X[T / x]=$ $\{s[a / x] \mid a \in T, s \in X\}$. Let $F: X \rightarrow \mathcal{P}(A)$ be a function. We define $X[F / x]=$ $\bigcup_{s \in X} s[F(s) / x]$. Let $C \subseteq A$. We define $X \upharpoonright C=\{s \in X \mid s(x) \in C$ for all $x \in D\}$.
Let $X$ be a team with domain $D$. Let $k \in \mathbb{Z}_{+}$, and let $y_{1}, \ldots, y_{k}$ be variable symbols. Assume that $\left\{y_{1}, \ldots, y_{k}\right\} \subseteq D$. We define $\operatorname{rel}\left(X,\left(y_{1}, \ldots, y_{k}\right)\right)=\left\{\left(s\left(y_{1}\right), \ldots, s\left(y_{k}\right)\right) \mid s \in X\right\}$.

Let $\tau$ be a relational vocabulary, i.e., a vocabulary containing relation symbols only. (In this article we consider only relational vocabularies.) The syntax of a logic based on team semantics is usually given in negation normal form. We shall also follow this convention in the current article. For this reason, we define the syntax of first-order logic as follows.

$$
\varphi::=R(\bar{x})|\neg R(\bar{x})| x_{1}=x_{2}\left|\neg x_{1}=x_{2}\right|\left(\varphi_{1} \vee \varphi_{2}\right)\left|\left(\varphi_{1} \wedge \varphi_{2}\right)\right| \exists x \varphi \mid \forall x \varphi,
$$

where $R \in \tau$. The first four formula formation rules above introduce first-order literals to the language. Below we shall consider logics $\mathrm{FO}(\mathcal{A})$, where the above syntax is extended by clauses of the type $A_{Q}\left(\bar{y}_{1}, \ldots, \bar{y}_{k}\right)$. Here $A_{Q}$ is (a symbol corresponding to) a generalized atom in $\mathcal{A}$ and each $\bar{y}_{i}$ is a tuple of variables. Before considering such novel atoms, let us define lax team semantics for first-order logic.

- Definition $1([15,25])$. Let $\mathfrak{A}$ be a model and $X$ a team with codomain $A$. The satisfaction relation $\mathfrak{A} \models_{X} \varphi$ is defined as follows.

1. If $\varphi$ is a first-order literal, then $\mathfrak{A} \models_{X} \varphi$ iff for all $s \in X: \mathfrak{A}, s \models_{\text {FO }} \varphi$. Here $\models_{\text {FO }}$ refers to the ordinary Tarskian satisfaction relation of first-order logic.
2. $\mathfrak{A} \models_{X} \psi \wedge \varphi$ iff $\mathfrak{A} \models_{X} \psi$ and $\mathfrak{A} \models_{X} \varphi$.
3. $\mathfrak{A} \models_{X} \psi \vee \varphi$ iff there exist teams $Y$ and $Z$ such that $X=Y \cup Z, \mathfrak{A} \models_{Y} \psi$, and $\mathfrak{A} \models_{Z} \varphi$.
4. $\mathfrak{A} \models_{X} \exists x \psi$ iff $\mathfrak{A} \models_{X[F / x]} \psi$ for some $F: X \rightarrow(\mathcal{P}(A) \backslash\{\emptyset\})$.
5. $\mathfrak{A} \models_{X} \forall x \psi$ iff $\mathfrak{A} \models_{X[A / x]} \psi$.

Finally, a sentence $\varphi$ is true in a model $\mathfrak{A}(\mathfrak{A} \models \varphi)$ if $\mathfrak{A} \models_{\{\emptyset\}} \varphi$.
Proposition $2([15,25])$. Let $\psi$ be a formula of first-order logic. We have $\mathfrak{A} \models_{X} \psi$ iff $\mathfrak{A}, s \models_{\text {FO }} \psi$ for all $s \in X$.

In this paper we consider first-order logic extended with generalized dependency atoms. Before formally introducing the notion of a generalized dependency atom, we recall some particular atoms familiar from the literature related to team semantics.

Dependence atoms $=(\bar{x}, y)$, inspired by the slashed quantifiers of Hintikka and Sandu [14], were introduced by Väänänen [25]. The intuitive meaning of the atom $=(\bar{x}, y)$ is that the value of the variable $y$ depends solely on the values of the variables in $\bar{x}$. The semantics for dependence atoms is defined as follows: $\mathfrak{A} \models_{X}=(\bar{x}, y)$ iff $\forall s, s^{\prime} \in X$ : if $s(\bar{x})=$ $s^{\prime}(\bar{x})$ then $s(y)=s^{\prime}(y)$. Dependence logic (D) is the extension of first-order logic with dependence atoms.

While dependence atoms of dependence logic declare dependences between variables, independence atoms, introduced by Grädel and Väänänen [11], do just the opposite; independence atoms are used to declare independencies between variables. Independence atom
is an atomic formula of the form $\left(x_{1}, \ldots, x_{k}\right) \perp_{\left(z_{1}, \ldots, z_{t}\right)}\left(y_{1}, \ldots, y_{l}\right)$ with the intuitive meaning that for any fixed interpretation of the variables $z_{1}, \ldots, z_{t}$, the interpretations of the variables $x_{1}, \ldots, x_{k}$ are independent of the interpretations of the variables $y_{1}, \ldots, y_{l}$. The semantics for independence atoms is defined as follows:

$$
\begin{aligned}
& \mathfrak{A} \models_{X}\left(x_{1}, \ldots, x_{k}\right) \perp_{\left(z_{1}, \ldots, z_{t}\right)}\left(y_{1}, \ldots, y_{l}\right) \text { iff } \forall s, s^{\prime} \in X \exists s^{\prime \prime} \in X: \bigwedge_{i \leq t} s\left(z_{i}\right)=s^{\prime}\left(z_{i}\right) \\
& \text { implies that } \bigwedge_{i \leq k} s^{\prime \prime}\left(x_{i}\right)=s\left(x_{i}\right) \wedge \bigwedge_{i \leq t} s^{\prime \prime}\left(z_{i}\right)=s\left(z_{i}\right) \wedge \bigwedge_{i \leq l} s^{\prime \prime}\left(y_{i}\right)=s^{\prime}\left(y_{i}\right) .
\end{aligned}
$$

Independence logic (Ind) is the extension of first-order logic with independence atoms.
Galliani [6] introduced inclusion and exclusion atoms. The intuitive meaning of the inclusion atom $\left(x_{1}, \ldots, x_{n}\right) \subseteq\left(y_{1}, \ldots, y_{n}\right)$ is that tuples interpreting the variables $x_{1} \ldots, x_{n}$ are also tuples interpreting $y_{1}, \ldots, y_{n}$. The intuitive meaning of the exclusion atom $\left(x_{1}, \ldots, x_{n}\right) \mid$ $\left(y_{1}, \ldots, y_{n}\right)$ on the other hand is that tuples interpreting the variables $x_{1} \ldots, x_{n}$ and the tuples interpreting $y_{1}, \ldots, y_{n}$ are distinct. The semantics for inclusion atoms and exclusion atoms is defined as follows:

$$
\mathfrak{A} \models_{X} \bar{x} \subseteq \bar{y} \text { iff } \forall s \in X \exists s^{\prime} \in X: s(\bar{x})=s^{\prime}(\bar{y}), \quad \mathfrak{A} \models_{X} \bar{x} \mid \bar{y} \text { iff } \forall s, s^{\prime} \in X: s(\bar{x}) \neq s^{\prime}(\bar{y}) .
$$

The extension of first-order logic with inclusion atoms (exclusion atoms) is called inclusion logic (exclusion logic) and denoted by Inc (Exc). The extension of first-order logic with both inclusion atoms and exclusion atoms is called inclusion/exclusion logic and denoted by Inc/Exc.

### 2.2 Generalized atoms

In this section we first give the well known definition of generalized quantifiers (Lindström quantifiers [21]). We then show how each generalized quantifier naturally gives rise to a generalized atom. Finally, we discuss on some fundamental properties of first-order logic extended with generalized atoms. Generalized atoms were first defined in [20].

Let $\left(i_{1}, \ldots, i_{n}\right)$ be a nonempty sequence of positive integers. A generalized quantifier of the type $\left(i_{1}, \ldots, i_{n}\right)$ is a class $\mathcal{C}$ of structures $\left(A, B_{1}, \ldots, B_{n}\right)$ such that the following conditions hold.

1. $A \neq \emptyset$, and for each $j \in\{1, \ldots, n\}$, we have $B_{j} \subseteq A^{i_{j}}$.
2. If $\left(A^{\prime}, B_{1}^{\prime}, \ldots, B_{n}^{\prime}\right) \in \mathcal{C}$ and if there is an isomorphism $f: A^{\prime} \rightarrow A^{\prime \prime}$ from $\left(A^{\prime}, B_{1}^{\prime}, \ldots, B_{n}^{\prime}\right)$ to another structure $\left(A^{\prime \prime}, B_{1}^{\prime \prime}, \ldots, B_{n}^{\prime \prime}\right)$, then $\left(A^{\prime \prime}, B_{1}^{\prime \prime}, \ldots, B_{n}^{\prime \prime}\right) \in \mathcal{C}$.

Let $Q$ be a generalized quantifier of the type $\left(i_{1}, \ldots, i_{n}\right)$. Let $\mathfrak{A}$ be a model with the domain $A$. We define $Q^{\mathfrak{A}}$ to be the set $\left\{\left(B_{1}, \ldots, B_{n}\right) \mid\left(A, B_{1}, \ldots, B_{n}\right) \in Q\right\}$.

Let $n$ be a positive integer. Let $Q$ be a generalized quantifier of the type $\left(i_{1}, \ldots, i_{n}\right)$. Extend the syntax of first-order logic with atomic expressions of the type $A_{Q}\left(\bar{y}_{1}, \ldots, \bar{y}_{n}\right)$, where each $\bar{y}_{j}$ is a tuple of variables of length $i_{j}$. Let $X$ be a team whose domain contains all variables occurring in the tuples $\bar{y}_{1}, \ldots, \bar{y}_{n}$. Extend team semantics such that $\mathfrak{A} \models_{X} A_{Q}\left(\bar{y}_{1}, \ldots, \bar{y}_{n}\right)$ if and only if $\left(\operatorname{rel}\left(X, \bar{y}_{1}\right), \ldots, \operatorname{rel}\left(X, \bar{y}_{n}\right)\right) \in Q^{\mathfrak{A}}$. The generalized quantifier $Q$ defines a generalized atom $A_{Q}$ of the type $\left(i_{1}, \ldots, i_{n}\right)$.

A generalized atom $A_{Q}$ is downwards closed if for all $\mathfrak{A}, X$ and $\bar{y}_{1}, \ldots, \bar{y}_{k}$, it holds that if $\mathfrak{A} \models_{X} A_{Q}\left(\bar{y}_{1}, \ldots, \bar{y}_{k}\right)$ and $Y \subseteq X$, then $\mathfrak{A} \models_{Y} A_{Q}\left(\bar{y}_{1}, \ldots, \bar{y}_{k}\right)$. Similarly, a generalized atom $A_{Q}$ is closed under substructures if for all $\mathfrak{A}, X$ and $\bar{y}_{1}, \ldots, \bar{y}_{k}$, it holds that if $\mathfrak{A} \models_{X} A_{Q}\left(\bar{y}_{1}, \ldots, \bar{y}_{k}\right)$, $\mathfrak{A}^{\prime}:=\mathfrak{A} \upharpoonright B$ and $X^{\prime}:=X \upharpoonright B$ for some $B \subseteq A$, then we have $\mathfrak{A}^{\prime} \models_{X^{\prime}} A_{Q}\left(\bar{y}_{1}, \ldots, \bar{y}_{k}\right)$. Finally, a generalized atom $A_{Q}$ is universe independent if for all $\mathfrak{A}, \mathfrak{B}, X$ and $\bar{y}_{1}, \ldots, \bar{y}_{k}$, where
both $A$ and $B$ are codomains for $X$, it holds that $\mathfrak{A} \models_{X} A_{Q}\left(\bar{y}_{1}, \ldots, \bar{y}_{k}\right)$ if and only if $\mathfrak{B} \models_{X} A_{Q}\left(\bar{y}_{1}, \ldots, \bar{y}_{k}\right)$.

Let $\varphi$ be a formula of first-order logic, possibly extended with generalized atoms. The set $\operatorname{Fr}(\varphi)$ of free variables of $\varphi$ is defined in the same way as in first-order logic. The set $\operatorname{Fr}\left(A_{Q}\left(\bar{y}_{1}, \ldots, \bar{y}_{k}\right)\right)$ of course contains exactly all variable that occur in the tuples $\bar{y}_{i}$. The satisfiability problem of a (possibly team-semantics-based) logic $L$ takes as an input a sentence of $L$ and asks whether $\mathfrak{A} \models \varphi$ for some model $\mathfrak{A}$. The validity problem asks, given a sentence $\varphi$, whether $\mathfrak{A} \models \varphi$ for all models $\mathfrak{A}$.

Let $k \in \mathbb{Z}_{+}$and let $A_{Q}$ be a generalized atom of the type $\left(i_{1}, \ldots, i_{n}\right)$, where $i_{j} \leq k$ for each $j$. Let $\varphi\left(R_{1}, \ldots, R_{n}\right)$ be a sentence of $\Sigma_{1}^{1}\left(\mathrm{FOC}^{k}\right)$ with unquantified relation symbols $R_{1}, \ldots, R_{n}$ of arities $i_{1}, \ldots, i_{n}$, respectively. Assume that for all models $\mathfrak{A}$ and teams $X$ with codomain $A$ and domain containing the variables in $A_{Q}\left(\bar{x}_{1}, \ldots, \bar{x}_{n}\right)$, we have $\mathfrak{A} \models_{X} A_{Q}\left(\bar{x}_{1}, \ldots, \bar{x}_{n}\right)$ iff

$$
\left(\mathfrak{A}, R_{1}:=\operatorname{rel}\left(X, \bar{x}_{1}\right), \ldots, R_{n}:=\operatorname{rel}\left(X, \bar{x}_{n}\right)\right) \models_{\text {fo }} \varphi\left(R_{1}, \ldots, R_{n}\right) .
$$

Then we say that the atom $A_{Q}$ is definable in $\Sigma_{1}^{1}\left(\mathrm{FOC}^{k}\right)$.
We now show that, for any generalized atom $A_{Q}$, the $\operatorname{logic} \mathrm{FO}\left(A_{Q}\right)$ has the so-called locality property. We also show that, for a downwards closed atom $A_{Q}$, all formulas of FO $\left(A_{Q}\right)$ satisfy the downwards closure property. These two properties have previously turned out to be very useful in the study of dependence logic.

Let $X$ be a team with domain $\left\{x_{1}, \ldots, x_{k}\right\}$, and let $V \subseteq\left\{x_{1}, \ldots, x_{k}\right\}$. We denote by $X(V)$ the team $\{s \upharpoonright V \mid s \in X\}$ with the domain $V$. The following proposition shows that the truth of an $\mathrm{FO}\left(A_{Q}\right)$-formula depends only on the interpretations of the variables occurring free in the formula. The proof uses the fact that generalized atoms satisfy the claim by definition. Otherwise the proof is identical to the corresponding proof given in [6].
Proposition 3 (Locality). Let $A_{Q}$ be a generalized atom and $\varphi \in \mathrm{FO}\left(A_{Q}\right)$ a formula. If $V \supseteq \operatorname{Fr}(\varphi)$, then $\mathfrak{A} \models_{X} \varphi$ if and only if $\mathfrak{A} \models_{X(V)} \varphi$.

The next proposition is also very useful. The proof is almost identical to the corresponding proof for dependence logic, see [25]. The additional case for generalized atoms follows by the assumption of downwards closure.

- Proposition 4 (Downward closure). Let $A_{Q}$ be a downwards closed generalized atom. Suppose $\varphi$ is an $\mathrm{FO}\left(A_{Q}\right)$-formula, $\mathfrak{A}$ a model, and $Y \subseteq X$ teams. Then $\mathfrak{A} \models_{X} \varphi$ implies $\mathfrak{A} \models_{Y} \varphi$.


## 3 Satisfiability problems of logics $\mathrm{FO}^{2}(\mathcal{A})$

In this section we show that for any finite collection $\mathcal{A}$ of $\Sigma_{1}^{1}\left(\mathrm{FOC}^{2}\right)$-definable atoms $A_{Q}$, both $\operatorname{Sat}\left(\mathrm{FO}^{2}(\mathcal{A})\right)$ and $\operatorname{FinSat}\left(\mathrm{FO}^{2}(\mathcal{A})\right)$ are NEXPTIME-complete. Our proof relies on a translation from $\mathrm{FO}^{2}(\mathcal{A})$ into $\Sigma_{1}^{1}\left(\mathrm{FOC}^{2}\right)$ and the fact that $\operatorname{SAT}\left(\mathrm{FOC}^{2}\right)$ and $\operatorname{FinSAT}\left(\mathrm{FOC}^{2}\right)$ are NEXPTIME-complete [23].

We start by establishing a more general translation. We show that for every $k \geq 1$ and every $\Sigma_{1}^{1}\left(\mathrm{FOC}^{k}\right)$ definable atom $A_{Q}$, we have $\mathrm{FO}^{k}\left(A_{Q}\right) \leq \Sigma_{1}^{1}\left(\mathrm{FOC}^{k}\right)$. Note that strictly speaking $\mathrm{FO}^{k}\left(A_{Q}\right)$ uses only one atom $A_{Q}$ instead of a finite collection $\mathcal{A}$ of atoms, but our proof below generalizes directly to the case with a finite collection of atoms. The reason for considering a single atom is simply to keep the notation light.

When considering $k$-variable logic, we let $\left\{x_{1}, \ldots, x_{k}\right\}$ denote the $k$ distinct variables used in the syntax of the logic, and we let $\operatorname{rel}(X)$ denote $\operatorname{rel}\left(X,\left(x_{1}, \ldots, x_{k}\right)\right)$. The proof of the following lemma (see the full version in arXiv [19]) significantly modifies and extends the argument establishing Lemma 3.3.14 of [26]. See also [18] and Theorem 6.2 in [25].

- Lemma 5. Assume that $k, t \geq 1$. Let $\tau$ be a relational vocabulary, let $R \notin \tau$ be a $k$-ary relation symbol and let $A_{Q}$ be a $\Sigma_{1}^{1}\left(\mathrm{FOC}^{k}\right)$-definable atom of type $\left(i_{1}, \ldots, i_{t}\right)$, where $i_{j} \leq k$ for each $j$. For every formula $\varphi \in \operatorname{FO}^{k}\left(A_{Q}\right)$ there exists a sentence $\varphi^{*} \in \Sigma_{1}^{1}\left(\mathrm{FOC}^{k}\right)(\tau \cup\{R\})$ such that for every model $\mathfrak{A}$ and team $X$ with codomain $A$ and $\operatorname{dom}(X)=\left\{x_{1}, \ldots, x_{k}\right\}$, we have

$$
\begin{equation*}
\mathfrak{A} \models_{X} \varphi \quad \text { iff } \quad(\mathfrak{A}, \operatorname{rel}(X)) \models \varphi^{*}, \tag{1}
\end{equation*}
$$

where $(\mathfrak{A}, \operatorname{rel}(X))$ is the expansion $\mathfrak{A}^{\prime}$ of $\mathfrak{A}$ into the vocabulary $\tau \cup\{R\}$ such that $R^{\mathfrak{A}^{\prime}}:=\operatorname{rel}(X)$. Moreover $\varphi^{*}$ is computable from $\varphi$ in polynomial time.

- Theorem 6. For every $k \geq 1$ and for every $\Sigma_{1}^{1}\left(\mathrm{FOC}^{k}\right)$-definable atom $A_{Q}$ it holds that $\mathrm{FO}^{k}\left(A_{Q}\right) \leq \Sigma_{1}^{1}\left(\mathrm{FOC}^{k}\right)$, i.e., for every sentence of $\mathrm{FO}^{k}\left(A_{Q}\right)$, there exists an equivalent sentence of $\Sigma_{1}^{1}\left(\mathrm{FOC}^{k}\right)$.

Proof. Let $\tau$ be a relational vocabulary, $k \geq 1$, and $A_{Q}$ a $\Sigma_{1}^{1}\left(\mathrm{FOC}^{k}\right)$-definable atom. Let $\varphi$ be an $\mathrm{FO}^{k}\left(A_{Q}\right)(\tau)$-sentence and $\varphi^{*}=\exists R_{1} \ldots \exists R_{n} \psi$ the related $\Sigma_{1}^{1}\left(\mathrm{FOC}^{k}\right)(\tau \cup\{R\})$-sentence given by Lemma 5 . The following conditions are equivalent.

1. $\mathfrak{A} \models \varphi$.
2. $\mathfrak{A} \models_{X} \varphi$ for some nonempty team $X$ such that $\operatorname{dom}(X)=\left\{x_{1}, \ldots, x_{k}\right\}$.
3. $(\mathfrak{A}, \operatorname{rel}(X)) \models \varphi^{*}$ for some nonempty team $X$ such that $\operatorname{dom}(X)=\left\{x_{1}, \ldots, x_{k}\right\}$.
4. $(\mathfrak{A}, R) \models \exists R_{1} \ldots \exists R_{n}\left(\exists x_{1} \ldots \exists x_{k} R\left(x_{1}, \ldots, x_{k}\right) \wedge \psi\right)$ for some $R \subseteq A^{k}$.
5. $\mathfrak{A} \models \exists R \exists R_{1} \ldots \exists R_{n}\left(\exists x_{1} \ldots \exists x_{k} R\left(x_{1}, \ldots, x_{k}\right) \wedge \psi\right)$.

The equivalence of 1 and 2 follows from Proposition 3 and the fact that $\operatorname{Fr}(\varphi)=\emptyset$. By Lemma 5, conditions 2 and 3 are equivalent. The equivalence of 3 and 4 follows from the fact that $\varphi^{*}=\exists R_{1} \ldots \exists R_{n} \psi$. The conditions 4 are 5 clearly equivalent.

- Theorem 7. Let $A_{Q}$ be a $\Sigma_{1}^{1}\left(\mathrm{FOC}^{2}\right)$-definable generalized atom. Then $\operatorname{Sat}\left(\mathrm{FO}^{2}\left(A_{Q}\right)\right)$ and $\operatorname{FinSAt}\left(\mathrm{FO}^{2}\left(A_{Q}\right)\right)$ are NEXPTIME-complete.

Proof. Since the translation $\varphi \mapsto \varphi^{*}$ is computable in polynomial time and (finite) satisfiability of $\Sigma_{1}^{1}\left(\mathrm{FOC}^{2}\right)$ can be checked in NEXPTIME [23], we conclude that both $\operatorname{SAT}\left(\mathrm{FO}^{2}\left(A_{Q}\right)\right)$ and $\operatorname{FinSat}\left(\mathrm{FO}^{2}\left(A_{Q}\right)\right)$ are in NEXPTIME. On the other hand, since $\mathrm{FO}^{2} \leq \mathrm{FO}^{2}\left(A_{Q}\right)$ by Proposition 2, and since both $\operatorname{SAT}\left(\mathrm{FO}^{2}\right)$ and $\operatorname{FinSat}\left(\mathrm{FO}^{2}\right)$ are NEXPTIME-hard [9], it follows that both $\operatorname{Sat}\left(\mathrm{FO}^{2}\left(A_{Q}\right)\right)$ and also $\operatorname{FinSat}\left(\mathrm{FO}^{2}\left(A_{Q}\right)\right)$ are as well.

The result of Theorem 7 can be directly generalized to concern finite collections $\mathcal{A}$ of generalized atoms. The proof of the following theorem is practically the same as that of Theorem 7.

- Theorem 8. Let $\mathcal{A}$ be a finite collection of $\Sigma_{1}^{1}\left(\mathrm{FOC}^{2}\right)$-definable generalized atoms. The satisfiability and the finite satisfiability problems of $\mathrm{FO}^{2}(\mathcal{A})$ are NEXPTIME-complete.

We shall next make use of Theorem 8 in order to show that the satisfiability and the finite satisfiability problems of two-variable fragments of dependence logic, inclusion logic, exclusion logic and independence logic are NEXPTIME-complete. The result for two-variable dependence logic was already established in [18]. Note that when regarded as generalized atoms, each of the dependency notions above correspond to a collection of generalized atoms; for example the atomic formulas $=(x, y)$ and $=(x, y, z)$ refer to two different atoms, one of type (2) and the other of type (3). However, in order to capture the two-variable fragments of of these logics, we only need a finite number of generalized atoms for each logic, as we shall see. We define $\varphi_{\text {const }}:=\exists \leq 1 x R(x), \varphi_{\text {dep }}:=\forall x \exists \leq 1 y R(x, y), \varphi_{\text {inc }}:=\forall x \forall y(R(x, y) \rightarrow$
$S(x, y)), \quad \varphi_{e x c}:=\forall x \forall y(R(x, y) \rightarrow \neg S(x, y)), \varphi_{\text {ind }}:=\forall x \forall y((\exists y R(x, y) \wedge \exists x R(x, y)) \rightarrow$ $R(x, y))$.

The formulas $\varphi_{\text {const }}, \varphi_{\text {dep }}, \varphi_{\text {inc }}, \varphi_{\text {exc }}$ and $\varphi_{\text {ind }}$ define the generalized atoms $A_{\text {const }}$ of type (1), $A_{\text {dep }}$ of type (2), $A_{\text {inc }}$ of type (2,2), $A_{\text {exc }}$ of type (2,2), and $A_{\text {ind }}$ of type (2), respectively.

- Theorem 9. The satisfiability and finite satisfiability problems of the two-variable fragments of dependence logic, inclusion logic, exclusion logic, inclusion/exclusion logic, and independence logic are all NEXPTIME-complete.

Proof. The proof proceeds via polynomial time translations $\mathrm{D}^{2} \rightarrow \mathrm{FO}^{2}\left(\left\{A_{\text {const }}, A_{\text {dep }}\right\}\right)$, $\mathrm{Inc}^{2} \rightarrow \mathrm{FO}^{2}\left(A_{\text {inc }}\right), \mathrm{Exc}^{2} \rightarrow \mathrm{FO}^{2}\left(A_{\text {exc }}\right), \quad \operatorname{Inc} / \mathrm{Exc}^{2} \rightarrow \quad \mathrm{FO}^{2}\left(A_{\text {inc }}, A_{\text {exc }}\right)$, $\mathrm{Ind}^{2} \rightarrow \mathrm{FO}^{2}\left(\left\{A_{\text {const }}, A_{\text {dep }}, A_{\text {ind }}\right\}\right)$ that preserve equivalence. The result then follows from Lemma 8 and the fact that the generalised atoms $A_{\text {const }}, A_{\text {dep }}, A_{\text {exc }}, A_{\text {inc }}, A_{\text {ind }}$ are all $\Sigma_{1}^{1}\left(\mathrm{FOC}^{2}\right)$-definable. For the full proof, see [19].

## 4 Undecidability via non-tiling

In this section we introduce structures and methods that we will later employ to prove undecidability of the validity problem of two-variable dependence logic. Curiously, all attempts (by us or known to us) to use the standard ( $\Pi_{1}^{0}$-complete) tiling problem for the undecidability proof have failed; we will instead use the ( $\Sigma_{1}^{0}$-complete) non-tiling problem in our arguments below.

The grid is the structure $\mathfrak{G}=\left(\mathbb{N}^{2}, V, H\right)$, where $V=\left\{((i, j),(i, j+1)) \in \mathbb{N}^{2} \times \mathbb{N}^{2} \mid i, j \in \mathbb{N}\right\}$ and $H=\left\{((i, j),(i+1, j)) \in \mathbb{N}^{2} \times \mathbb{N}^{2} \mid i, j \in \mathbb{N}\right\}$. A function $t: 4 \longrightarrow \mathbb{N}$ is called a tile type. Define the set TILES $:=\left\{P_{t} \mid t\right.$ is a tile type $\}$ of unary relation symbols. The unary relation symbols in the set TILES are called tiles. The number $t(0)$ is the top colour, $t(1)$ the right colour, $t(2)$ the bottom colour, and $t(3)$ the left colour of $P_{t}$.

Let $T$ be a finite nonempty set of tiles and $V$ and $H$ binary relation symbols. We say that a structure $\mathfrak{A}=(A, V, H)$ is $T$-tilable, if there exists an expansion of $\mathfrak{A}$ to the vocabulary $\{H, V\} \cup\left\{P_{t} \mid P_{t} \in T\right\}$ such that the following conditions hold for all $u, v \in A$.

1. The point $u$ belongs to the extension of exactly one symbol $P_{t}$ in $T$.
2. If $u H v, P_{t}(u)$ and $P_{s}(v)$, then the right colour of $P_{t}$ is the same as the left colour of $P_{s}$.
3. If $u V v, P_{t}(u)$ and $P_{s}(v)$, then the top colour of $P_{t}$ is the same as the bottom colour of $P_{s}$.

We will next define the tiling problem and the non-tiling problem. Let $\mathcal{F}$ denote the set of finite, nonempty subsets of TILES. We define $\mathcal{T}:=\{T \in \mathcal{F} \mid \mathfrak{G}$ is $T$-tilable $\}$ and $\overline{\mathcal{T}}^{\prime}:=\{T \in \mathcal{F} \mid \mathfrak{G}$ is not $T$-tilable $\}$. The tiling problem (non-tiling problem, resp.) is the membership problem of the set $\mathcal{T}\left(\overline{\mathcal{T}}^{\prime}\right.$, resp.) with the input set $\mathcal{F}$.

- Theorem 10 ([2]). The tiling problem is $\Pi_{1}^{0}$-complete.

The non-tiling problem is the complement of the tiling problem. Thus the following corollary follows.

- Corollary 11. The non-tiling problem is $\Sigma_{1}^{0}$-complete.

The proof of the following lemma is straightforward.

- Lemma 12. There is a computable function associating each input $T$ to the non-tiling problem with an $\mathrm{FO}^{2}$-sentence $\varphi_{T}$ of the vocabulary $\tau:=\{H, V\} \cup T$ such that for every structure $\mathfrak{A}$ of the vocabulary $\{H, V\}$, the structure $\mathfrak{A}$ is not $T$-tilable iff for every expansion $\mathfrak{A}^{*}$ of $\mathfrak{A}$ to the vocabulary $\tau$, it holds that $\mathfrak{A}^{*} \models \varphi_{T}$.
- Definition 13. Let $\tau=\{V, H\}$ be a vocabulary where $V$ and $H$ are binary relation symbols. Let $\mathfrak{A}=(A, V, H)$ be a $\tau$-structure. We say that $\mathfrak{A}$ is gridlike if the below conditions hold.

1. The extension of $V$ in $\mathfrak{A}$ is serial (i.e., $\forall x \in A \exists y \in A$ s.t. $V(x, y))$.
2. The extension of $H$ in $\mathfrak{A}$ is serial (i.e., $\forall x \in A \exists y \in A$ s.t. $H(x, y))$.
3. If $a, b, c, b^{\prime}, c^{\prime} \in A$ are such that $V(a, b), H(b, c), H\left(a, b^{\prime}\right)$, and $V\left(b^{\prime}, c^{\prime}\right)$, then $c=c^{\prime}$.

Note that it follows from the above definition that in gridlike structures, for every point $a$, there exist points $b, c$ and $d$ such that $H(a, b), V(a, c), V(b, d)$, and $H(c, d)$.

Let $\tau$ be the vocabulary of gridlike structures and $U, P, Q, C$ unary relation symbols. We say that a $\tau \cup\{U, P, Q, C\}$-structure $\mathfrak{A}$ is striped and gridlike if the $\tau$-reduct of $\mathfrak{A}$ is gridlike, the extensions of $P$ and $Q$ in $\mathfrak{A}$ are distinct singleton sets, the extension of $U$ in $\mathfrak{A}$ is the union of the extensions of $P$ and $Q$, and $\mathfrak{A}$ has the following property (intuitively $C$ creates stripes in $\mathfrak{A}$ ):

$$
\begin{equation*}
(H(a, b) \Rightarrow(C(a) \Leftrightarrow C(b))) \text { and }(V(a, b) \Rightarrow(C(a) \Leftrightarrow \neg C(b))) \tag{2}
\end{equation*}
$$

The following lemma can be now proven by a simple inductive argument.

- Lemma 14. If $\mathfrak{A}$ is striped and gridlike, then there exists a homomorphism from the grid into $\mathfrak{A}$.
- Lemma 15. Let $T$ be an input to the non-tiling problem. The grid is non-T-tilable iff (the $\{H, V\}$-reduct of) every striped gridlike structure is non-T-tilable.

Proof. The direction from left to right follows from Lemma 14 in a straightforward way. The converse holds since the grid is an $\{\mathrm{H}, \mathrm{V}\}$-reduct of a striped gridlike structure.

## 5 The validity problem of $\mathrm{D}^{2}$ is undecidable

In this section we give a reduction from the non-tiling problem to the validity problem of $\mathrm{D}^{2}$.
Let $\tau=\{V, H, C, U, P, Q\}$ be the vocabulary of striped gridlike structures. We will first define a formula $\varphi_{\text {non-grid }}$ of $\mathrm{D}^{2}$ such that $\mathfrak{A}$ is not striped and gridlike iff $\mathfrak{A} \models \varphi_{\text {non-grid }}$. We first notice that the first two conditions of Definition 13 are easy to deal with. Define $\varphi_{\text {non-serial }}:=\exists x \forall y \neg V(x, y) \vee \exists x \forall y \neg H(x, y)$. The third condition of Definition 13 is nontrivial. In the below construction, we will use the predicates $P, Q, U$ for counting (only). We will first show how to force the extensions of $P$ and $Q$ to be distinct singletons and the extension of $U$ to be the union of $P$ and $Q$. The next formulae will be used for dealing with the cases where this does not hold.

$$
\begin{aligned}
\varphi_{\text {non-singleton }}(X):= & \forall x \neg X(x) \vee \exists x \exists y(X(x) \wedge X(y) \wedge \neg x=y) \\
\varphi_{\text {non-distinct }}(X, Y):= & \exists x(X(x) \wedge Y(x)) \\
\varphi_{\text {non-union }}(X, Y, Z):= & \exists x(X(x) \wedge(\neg Y(x) \vee \neg Z(x))) \vee \exists x(\neg X(x) \wedge(Y(x) \vee Z(x))) \\
\varphi_{|U| \neq 2}:= & \varphi_{\text {non-singleton }}(P) \vee \varphi_{\text {non-singleton }}(Q) \vee \varphi_{\text {non-distinct }}(P, Q) \\
& \vee \varphi_{\text {non-union }}(U, P, Q) .
\end{aligned}
$$

It is easy to check that the $\tau$-models $\mathfrak{A}$ such that $\mathfrak{A} \not \vDash \varphi_{|U| \neq 2}$ are exactly those models where the extensions of $P$ and $Q$ are distinct singletons and the extension of $U$ is the union of the extensions of $P$ and $Q$ (and thus the cardinality of the extension of $U$ is 2 ).

We will now show how to enforce Equation (2). The formula $\varphi_{\text {non-stripes }}$ below takes care of the cases where (2) does not hold. Define

$$
\varphi_{\text {non-stripes }}:=\exists x \exists y((H(x, y) \wedge(C(x) \leftrightarrow \neg C(y))) \vee(V(x, y) \wedge(C(x) \leftrightarrow C(y)))) .
$$

We are now ready to show how to deal with models that violate the last condition of Definition 13. To understand the intended meaning of the following formula, assume that the extension of $U$ is of size two and that the condition given by Equation (2) holds. Note also that from (2) it follows that if such points $c$ and $c^{\prime}$ exist that violate the last condition of Definition 13 , then $c$ and $c^{\prime}$ agree about $C$, i.e., we have $C(c)$ iff $C\left(c^{\prime}\right)$. We first deal with the case where $C(c)$ and $C\left(c^{\prime}\right)$ both hold. We denote by $\varphi_{\text {non- } C^{+}-j o i n}$ the following formula (whose meaning is fully explained in the proof of Lemma 16):

$$
\begin{aligned}
& \forall x(\neg U(x) \vee \exists y(C(y) \wedge=(y, x) \wedge \exists x(=(x, y) \wedge((=(x) \wedge H(x, y)) \vee(=(x) \wedge V(x, y))) \\
& \wedge \exists y(=(y) \wedge(V(y, x) \vee H(y, x)) \wedge \neg C(y))))))
\end{aligned}
$$

To deal with the case where $\neg C(c)$ and $\neg C\left(c^{\prime}\right)$, we define the formula $\varphi_{\text {non }-C^{-}-\text {join }}$ which is obtained from $\varphi_{\text {non }-C^{+-} \text {join }}$ by simultaneously replacing each $C(x)$ and $C(y)$ by $\neg C(x)$ and $\neg C(y)$, respectively. Finally, we define that $\varphi_{\text {non-join }}:=\varphi_{\text {non }-C^{+}-\text {join }} \vee \varphi_{\text {non- } C^{-}-\text {join }}$ and $\varphi_{\text {non-grid }}:=\varphi_{\text {non-serial }} \vee \varphi_{|U| \neq 2} \vee \varphi_{\text {non-stripes }} \vee \varphi_{\text {non-join }}$.

- Lemma 16. Let $\tau=\{V, H, C, U, P, Q\}$ be the vocabulary of striped gridlike structures. Let $\mathfrak{A}$ be a $\tau$-structure such that the extension of $U$ is of cardinality 2 . Assume the condition (2) holds. Then $\mathfrak{A} \models \varphi_{\text {non-join }}$ iff the last condition of Definition 13 fails in $\mathfrak{A}$.
Proof. From (2) it follows that if such $c$ and $c^{\prime}$ exist in $\mathfrak{A}$ that violate the last condition of Definition 13, then $c$ and $c^{\prime}$ agree on $C$. We will show that
$\mathfrak{A} \models \varphi_{\text {non- } C^{+}-\text {join }}$ iff the last condition of Definition 13 fails in $\mathfrak{A}$ for some $c, c^{\prime}$ s.t.

$$
\begin{equation*}
C(c) \& C\left(c^{\prime}\right) \tag{3}
\end{equation*}
$$

The analogous argument for $\varphi_{\text {non }-C^{-}-\text {join }}$ and the case where $\neg C(c)$ and $\neg C\left(c^{\prime}\right)$ hold is similar.

Below we denote by $\left\{\left(x_{1}, v_{1}\right), \ldots,\left(x_{k}, v_{k}\right)\right\}$ the variable assignment that maps $x_{i}$ to $v_{i}$ for each $i$. Let $u, u^{\prime}$ be the elements that are in the extension of $U$ in $\mathfrak{A}$. We thus have $\mathfrak{A} \models \varphi_{\text {non }-C^{+}-\text {join }}$ iff

$$
\left.\left.\left.\left.\begin{array}{rl}
\mathfrak{A} \models_{X_{1}} \exists y(C(y) & \wedge=(y, x) \wedge \exists x(=(x, y) \wedge((=(x) \wedge H(x, y)) \\
& \vee(=(x) \wedge V(x, y)))
\end{array}\right) \nexists y(=(y) \wedge(V(y, x) \vee H(y, x)) \wedge \neg C(y))\right)\right)\right), ~ \$
$$

where $X_{1}=\left\{\{(x, u)\},\left\{\left(x, u^{\prime}\right)\right\}\right\}$. Now, recalling that dependence logic has the downwards closure property (cf. proposition 4), we observe that the above holds if and only if there exist distinct (distinctness being due to the atom $=(y, x))$ points $c, c^{\prime}$ in the extension of $C$ such that

$$
\begin{aligned}
\mathfrak{A} \models_{X_{2}} \exists x(=(x, y) \wedge((=(x) \wedge H(x, y)) & \vee(=(x) \wedge V(x, y))) \\
& \wedge \exists y(=(y) \wedge(V(y, x) \vee H(y, x)) \wedge \neg C(y)))),
\end{aligned}
$$

where $X_{2}=\left\{\{(x, u),(y, c)\},\left\{\left(x, u^{\prime}\right),\left(y, c^{\prime}\right)\right\}\right\}$. The above holds if and only if there exist distinct points $b, b^{\prime}$ of $\mathfrak{A}$ such that $H(b, c)$ and $V\left(b^{\prime}, c^{\prime}\right)$ (or $V(b, c)$ and $H\left(b^{\prime}, c^{\prime}\right)$ in which case the argument is analogous) and

$$
\left.\mathfrak{A} \models_{X_{3}} \exists y(=(y) \wedge(V(y, x) \vee H(y, x)) \wedge \neg C(y))\right),
$$

where $X_{3}=\left\{\{(x, b),(y, c)\},\left\{\left(x, b^{\prime}\right),\left(y, c^{\prime}\right)\right\}\right\}$. The above holds if and only if there exists a point $a$ in $\mathfrak{A}$ such that $\neg C(a),(V(a, b)$ or $H(a, b))$ and $\left(V\left(a, b^{\prime}\right)\right.$ or $\left.H\left(a, b^{\prime}\right)\right)$. Since $C(c)$ and $C\left(c^{\prime}\right)$ hold, it follows from the assumption that (2) holds that $C(b)$ and $\neg C\left(b^{\prime}\right)$. Now since also $\neg C(a)$ holds, it follows again from (2) that $V(a, b)$ and $H\left(a, b^{\prime}\right)$. When all of the above is combined, we obtain (3). The analogous condition where $\neg C(c)$ and $\neg C\left(c^{\prime}\right)$ is proved similarly. Since (2) holds for $\mathfrak{A}$, any points $c$ and $c^{\prime}$ of $\mathfrak{A}$ that violate the last condition of Definition 13, must agree on $C$. Thus the lemma holds.

The next lemma follows from Lemma 16 together with the observations made above.

- Lemma 17. Let $\tau=\{V, H, C, U, P, Q\}$ be the vocabulary of striped gridlike structures and let $\mathfrak{A}$ be a $\tau$-model. Then $\mathfrak{A}$ is striped and gridlike iff $\mathfrak{A} \not \vDash \varphi_{\text {non-grid }}$.
- Theorem 18. The validity problem for $\mathrm{D}^{2}$ is undecidable (more precisely, $\Sigma_{1}^{0}$-hard).

Proof. We give a computable reduction from the non-tiling problem to the validity problem of $\mathrm{D}^{2}$. Since the former is $\Sigma_{1}^{0}$-complete (Corollary 11), we obtain $\Sigma_{1}^{0}$-hardness for the latter.

If $T$ is an input to the non-tiling problem, then $\varphi_{T}$ denotes the $\mathrm{FO}^{2}$-sentence given by Lemma 12 and $\varphi_{\text {non-T-tiling }}:=\left(\varphi_{\text {non-grid }} \vee \varphi_{T}\right)$. Let $\tau$ be as defined in Lemma 17. Let $C_{\tau, T}$ denote the class of all $\tau \cup T$-structures and let $\mathcal{C}_{s-\text { gridlike }}^{\tau, T}$ be the class of exactly all expansions of striped gridlike structures to the vocabulary $\tau \cup T$.

Let $T$ be an input to the non-tiling problem. We will show that the grid is non-Ttilable iff the $\mathrm{D}^{2}$-sentence $\varphi_{\text {non-T-tiling }}$ is valid. By definition, $\varphi_{\text {non-T-tiling }}$ is valid iff $\mathfrak{A} \models \varphi_{\text {non-grid }} \vee \varphi_{T}$ holds for every $\mathfrak{A} \in \mathcal{C}_{\tau, T}$. Since $\varphi_{\text {non-grid }}$ and $\varphi_{T}$ are sentences, the right-hand side of this equivalence is equivalent to the claim that

$$
\begin{equation*}
\forall \mathfrak{A} \in \mathcal{C}_{\tau, T}: \mathfrak{A} \models \varphi_{\text {non-grid }} \text { or } \mathfrak{A} \models \varphi_{T} . \tag{4}
\end{equation*}
$$

By Lemma $17, \mathfrak{B}^{*} \models \varphi_{\text {non-grid }}$ holds for every $\tau$-reduct $\mathfrak{B}^{*}$ of $\mathfrak{B} \in \mathcal{C}_{\tau, T}$ that is not striped and gridlike. Hence for every $\mathfrak{B} \in \mathcal{C}_{\tau, T}$ such that the $\tau$-reduct $\mathfrak{B}^{*}$ of $\mathfrak{B}$ is not striped and gridlike, it holds that $\mathfrak{B} \models \varphi_{\text {non-grid }}$. Thus (4) is equivalent to the claim that

$$
\begin{equation*}
\forall \mathfrak{A} \in \mathcal{C}_{s-\text { gridlike }}^{\tau, T}: \mathfrak{A} \models \varphi_{\text {non-grid }} \text { or } \mathfrak{A} \models \varphi_{T} . \tag{5}
\end{equation*}
$$

Now let $\mathfrak{B}$ be an arbitrary striped and gridlike $\tau$-structure. By Lemma $17, \mathfrak{B} \notin \varphi_{\text {non- grid }}$. Thus $\mathfrak{B}^{*} \not \models \varphi_{\text {non-grid }}$ for every expansion $\mathfrak{B}^{*}$ of $\mathfrak{B}$ to the vocabulary $\tau \cup T$. From this it follows that (5) is equivalent to the claim that

$$
\begin{equation*}
\forall \mathfrak{A} \in \mathcal{C}_{s-\text { gridlike }}^{\tau, T}: \mathfrak{A} \models \varphi_{T} . \tag{6}
\end{equation*}
$$

Thus, by Lemma 12, (6) holds if and only if every striped gridlike structure is non- $T$-tilable. Finally, from Lemma 15 it follows that this is equivalent to the claim that the grid is non- $T$-tilable.

## 6 Satisfiability of $\exists^{*} \forall^{*}$-formulas

In this section we consider the complexity of satisfiability for sentences of dependence logic and its variants in the prefix class $\exists^{*} \forall^{*}$. For first-order logic, the satisfiability and finite satisfiability problems of the prefix class $\exists^{*} \forall^{*}$ are known to be NEXPTIME-complete. The results hold for both the case with equality and the case without equality, see [3].

Let $\mathcal{A}$ be a collection of generalized atoms. We denote by $\exists^{*} \forall^{*}[\mathcal{A}]$ the class of sentences of $\operatorname{FO}(\mathcal{A})$ of the form $\exists x_{0} \cdots \exists x_{n} \forall y_{0} \cdots \forall y_{m} \theta$, where $\theta$ is a quantifier-free formula whose
generalized atoms are in $\mathcal{A}$. It is worth noting that, depending on the set $\mathcal{A}$, the expressive power and complexity of sentences in $\exists^{*} \forall^{*}[\mathcal{A}]$ can vary considerably even when $\mathcal{A}$ is finite and contains only computationally non-complex atoms. For example, there are universal sentences of dependence logic that define NP-complete problems [17]. Furthermore, every sentence of inclusion logic is equivalent to a sentence with a prefix of the form $\exists^{*} \forall^{1}[12]$ implying that the satisfiability problem of the $\exists^{*} \forall^{*}$-fragment of inclusion logic is undecidable.

Recall that we say that a formula $\varphi$ is closed under substructures if for all $\mathfrak{A}$ and $X$ it holds that if $\mathfrak{A} \models_{X} \varphi, \mathfrak{A}^{\prime}:=\mathfrak{A} \upharpoonright B$ and $X^{\prime}:=X \upharpoonright B$ for some $B \subseteq A$, then we have $\mathfrak{A}^{\prime} \models{ }_{X^{\prime}} \varphi$.

- Lemma 19. Let $\mathcal{A}$ be a collection of generalized atoms that are closed under substructures. Then the following conditions hold.

1. Suppose $\varphi \in \mathrm{FO}[\mathcal{A}]$ is of the form $\forall y_{0} \cdots \forall y_{m} \theta$, where $\theta$ is quantifier-free. Then $\varphi$ is closed under substructures.
2. Let $\varphi \in \exists^{*} \forall^{*}[\mathcal{A}]$ be a sentence. Then, if $\varphi$ is satisfiable, $\varphi$ has a model with at most $\max \{1, k\}$ elements, where $k$ refers to the number of existentially quantified variables in $\varphi$.

Proof. We will first prove claim (1). Suppose that $\varphi:=\forall y_{0} \cdots \forall y_{m} \theta$. We will first show the claim for quantifier-free formulas $\theta$, i.e., we will show that for all $\mathfrak{A}, X, \mathfrak{A}^{\prime}$, and $X^{\prime}$ such that $\mathfrak{A}^{\prime}:=\mathfrak{A} \upharpoonright B$ and $X^{\prime}:=X \upharpoonright B$ for some $B \subseteq A$, the following implication holds.

$$
\begin{equation*}
\mathfrak{A} \models_{X} \theta \Rightarrow \mathfrak{A}^{\prime} \models_{X^{\prime}} \theta \tag{7}
\end{equation*}
$$

The claim obviously holds if $\theta$ is a first-order literal. If $\theta$ is a generalized atom from $\mathcal{A}$, then the claim holds by assumption. The case $\theta:=\psi_{1} \wedge \psi_{2}$ follows immediately from the induction hypothesis. Let us then assume that $\theta:=\psi_{1} \vee \psi_{2}$. Since $\mathfrak{A} \models_{X} \theta$, there are sets $Y$ and $Z$ such that $Y \cup Z=X, \mathfrak{A} \models_{Y} \psi_{1}$ and $\mathfrak{A} \models_{Z} \psi_{2}$. By the induction hypothesis, we have $\mathfrak{A}^{\prime} \models_{Y^{\prime}} \psi_{1}$ and $\mathfrak{A}^{\prime} \models_{Z^{\prime}} \psi_{2}$, where $Y^{\prime}:=Y \upharpoonright B$ and $Z^{\prime}:=Z \upharpoonright B$. Since $Y^{\prime} \cup Z^{\prime}=X^{\prime}$, it follows that $\mathfrak{A}^{\prime} \models_{X^{\prime}} \theta$.

We will now show that the claim also holds for $\varphi$. Suppose that $\mathfrak{A} \models_{X} \varphi$. Then, by the truth definition, $\mathfrak{A} \models_{X\left[A / y_{0}\right] \cdots\left[A / y_{m}\right]} \theta$. Using (7), we have $\mathfrak{A}^{\prime} \models_{\left(X\left[A / y_{0}\right] \cdots\left[A / y_{m}\right]\right) \upharpoonright B} \theta$. It is easy to check that $\left(X\left[A / y_{0}\right] \cdots\left[A / y_{m}\right]\right) \upharpoonright B=(X \upharpoonright B)\left[B / y_{0}\right] \cdots\left[B / y_{m}\right]$. Hence we have $\mathfrak{A}^{\prime} \models_{X^{\prime}} \varphi$.

Let us then prove 2. Assume $\varphi$ is a sentence of the form $\exists x_{0} \cdots \exists x_{n} \forall y_{0} \cdots \forall y_{m} \theta$, where $\theta$ is quantifier-free, and that there is a structure $\mathfrak{A}$ such that $\mathfrak{A} \models \varphi$. Hence there exists functions $F_{i}$ such that $\mathfrak{A}=_{X} \forall y_{0} \cdots \forall y_{m} \theta$, where $X=\{\emptyset\}\left[F_{0} / x_{0}\right] \cdots\left[F_{n} / x_{n}\right]$. Let $s$ be some assignment in $X$. Let range $(s)$ denote the set of elements $b$ such that $s(x)=b$ for some variable $x$ in the domain of $s$. If range $(s) \neq \emptyset$ define $B:=\operatorname{range}(s)$, and if range $(s)=\emptyset$ (i.e., $s=\emptyset$ ), define $B=\{b\}$, where $b$ is an arbitrary element in $A$. By claim (1), the formula $\forall y_{0} \cdots \forall y_{m} \theta$ is closed under substructures. Thus $\mathfrak{A} \upharpoonright B \models_{X \upharpoonright B} \forall y_{0} \cdots \forall y_{m} \theta$. Thus it follows that $\mathfrak{A} \upharpoonright B \models \varphi$.

A generalized atom $A_{Q}$ is said to be polynomial time computable if the question whether $\mathfrak{A} \models_{X} A_{Q}\left(\bar{y}_{1}, \ldots, \bar{y}_{n}\right)$ holds can be decided in time polynomial in the size of $\mathfrak{A}$ and $X$. A class of atoms $\mathcal{A}$ is said to be uniformly polynomial time computable if there exists a polynomial function $f: \mathbb{N} \rightarrow \mathbb{N}$ such that for every atom $A_{Q} \in \mathcal{A}$ it holds that the question whether $\mathfrak{A} \models_{X} A_{Q}\left(\bar{y}_{1}, \ldots, \bar{y}_{n}\right)$ holds can be decided in time $f\left(|\mathfrak{A}|+|X|+\left|A_{Q}\left(\bar{y}_{1}, \ldots, \bar{y}_{n}\right)\right|\right)$. Note that every finite class of polynomial time computable atoms is also uniformly polynomial time computable.

The following theorem now follows from Lemma 19. We will make use of the recent result of Grädel showing that for a uniformly polynomial time computable collection $\mathcal{A}$ of atoms, the model checking problem for $\mathrm{FO}(\mathcal{A})$-formulas is in NEXPTIME [8].

- Theorem 20. Let $A_{Q}$ be a generalized atom that is closed under substructures and polynomial time computable. Then $\operatorname{SAT}\left(\exists^{*} \forall^{*}\left[A_{Q}\right]\right)$ and $\operatorname{FinSAT}\left(\exists^{*} \forall^{*}\left[A_{Q}\right]\right)$ are in 2NEXPTIME. If $\tau$ is a vocabulary consisting of relation symbols of arity at most $k, k \in \mathbb{Z}_{+}$, then $\operatorname{Sat}\left(\exists^{*} \forall^{*}\left[A_{Q}\right](\tau)\right)$ and $\operatorname{FinSAT}\left(\exists^{*} \forall^{*}\left[A_{Q}\right](\tau)\right)$ are NEXPTIME-complete.

Proof. Note first that the lower bounds follow from the fact that both $\operatorname{Sat}\left(\exists^{*} \forall^{*}\right)$ and $\operatorname{FinSat}\left(\exists^{*} \forall^{*}\right)$ are already NEXPTIME-complete. It hence suffices to show containments in 2NEXPTIME and NEXPTIME, respectively.

Let $\varphi \in \exists^{*} \forall^{*}\left[A_{Q}\right]$. By Lemma $19, \varphi$ is satisfiable if and only if it has a model of cardinality at most $|\varphi|$. We can decide satisfiability of $\varphi$ as follows: non-deterministically guess a structure $\mathfrak{A}$ of cardinality at most $|\varphi|$ and accept iff $\mathfrak{A} \models \varphi$. By the result of Grädel in [8], the question whether $\mathfrak{A} \models \varphi$ can be checked non-deterministically in exponential time with input $\mathfrak{A}$ and $\varphi$. Assume first that the maximum arity of relation symbols that may occur in $\varphi$ is not a fixed constant. Relation symbols of arity at most $|\varphi|$ may occur in $\varphi$. Thus the size of the binary encoding of a model $\mathfrak{A}$ of $\varphi$ such that $A \leq|\varphi|$ is worst case exponential with respect to $|\varphi|$. If, on the other hand, the maximum arity of relation symbols that can occur in $\varphi$ is a fixed constant, then the size of the encoding of $\mathfrak{A}$ is just worst case polynomial with respect to $|\varphi|$. Therefore it follows that our algorithm for checking satisfiability of $\varphi$ is in NEXPTIME in the case of fixed arity vocabularies and in 2NEXPTIME in the general case. The corresponding results for the finite satisfiability problem follow by the observation that $\exists^{*} \forall^{*}\left[A_{Q}\right]$ has the finite model property, Lemma 19.

- Corollary 21. Let $\mathcal{A}$ be a uniformly polynomial time computable class of generalized atoms that are closed under substructures. Then $\operatorname{SAT}\left(\exists^{*} \forall^{*}[\mathcal{A}]\right)$ and $\operatorname{FinSat}\left(\exists^{*} \forall^{*}[\mathcal{A}]\right)$ are in 2NEXPTIME. If $\tau$ is a vocabulary consisting of relation symbols of arity at most $k, k \in \mathbb{Z}_{+}$, then $\operatorname{SAT}\left(\exists^{*} \forall^{*}[\mathcal{A}](\tau)\right)$ and $\operatorname{FinSAT}\left(\exists^{*} \forall^{*}[\mathcal{A}](\tau)\right)$ are NEXPTIME-complete.

In the following sense Theorem 20 is optimal: there exists a polynomial time computable generalized atom $A_{Q}$ such that $\operatorname{Sat}\left(\exists^{3} \forall\left[A_{Q}\right]\right)$ and $\operatorname{FinSAT}\left(\exists^{3} \forall\left[A_{Q}\right]\right)$ are undecidable. This already holds for vocabularies with at least one binary relation symbol and a countably infinite set of unary relation symbols. Let $\varphi_{5-i n c}:=\forall x_{1} \ldots \forall x_{5}\left(R\left(x_{1}, \ldots, x_{5}\right) \rightarrow S\left(x_{1}, \ldots x_{5}\right)\right)$, and let $A_{5-\text { inc }}$ be the related generalized atom of the type (5,5), i.e., $A_{5-\text { inc }}$ is the 5 -ary inclusion atom interpreted as a generalized atom. Clearly $A_{5-i n c}$ is computable in polynomial time.

- Theorem 22. Let $\tau$ be a vocabulary consisting of one binary relation symbol and a countably infinite set of unary relation symbols. Then both $\operatorname{SAT}\left(\exists^{3} \forall\left[A_{5-\text { inc }}\right](\tau)\right)$ and $\operatorname{FinSAt}\left(\exists^{3} \forall\left[A_{5-\text { inc }}\right](\tau)\right)$ are undecidable.

Proof. It well known that for the Kahr class (i.e., the prefix class $\forall \exists \forall$ of FO with vocabulary $\tau$ ) the satisfiability and the finite satisfiability problems are undecidable (see, e.g., [3]). From the proof of [12, Theorem 5] it follows that there exists a polynomial time translation $\varphi \mapsto \varphi^{*}$ from the Kahr class into $\exists^{3} \forall\left[A_{5-\text { inc }}\right](\tau)$ such that $\mathfrak{A} \models_{X} \varphi \Leftrightarrow \mathfrak{A} \models_{X} \varphi^{*}$ holds for every model $\mathfrak{A}$ and team $X$ with codomain $A$. Thus $\operatorname{Sat}\left(\exists^{3} \forall\left[A_{5-\text { inc }}\right](\tau)\right)$ and $\operatorname{FinSat}\left(\exists \exists^{3} \forall\left[A_{5-\text { inc }}\right](\tau)\right)$ are undecidable.

It is easy to see that dependence atoms viewed as generalized atoms are closed under substructures since they are downwards closed and universe independent. Likewise, it is
straightforward to check that the class of dependence atoms is uniformly polynomial time computable. Hence we obtain:

- Corollary 23. Both the satisfiability and the finite satisfiability problems for the $\exists^{*} \forall^{*}$ sentences of dependence logic are in 2NEXPTIME. If $\tau$ is a vocabulary consisting of relation symbols of arity at most $k$, then the satisfiability and the finite satisfiability problems for the $\exists^{*} \forall^{*}$-sentences of dependence logic over the vocabulary $\tau$ are NEXPTIME-complete.


## 7 Conclusion

We have tied some loose ends concerning the complexity of predicate logics based on team semantics. Using a general approach, we have shown that the satisfiability and the finite satisfiability problems of the two-variable fragments of inclusion logic, exclusion logic, inclusion/exclusion logic, and independence logic are all NEXPTIME-complete. Additionally, we have shown that the satisfiability and the finite satisfiability problems of the prefix class $\exists^{*} \forall^{*}$ of dependence logic are NEXPTIME-complete for any vocabulary of bounded arity, and in 2NEXPTIME in the general case. The general approach we have employed of course also implies a range of other results on team-semantics-based logics. Finally, we have proved that the validity problem of two-variable dependence logic is undecidable, thereby answering an open problem from the literature on team semantics.

This article clears path to a more comprehensive classification of the decidability and complexity of different fragments of logics with generalized atoms and team semantics. In the future, we aim to identify further interesting related systems with a decidable satisfiability problem.

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