

Tightening Faster Model Checking of First-Order Graph Properties, via Tarski's Calculus of Relations

YN: Tentative

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Abstract

This is a SUPER-STRONG-PAPER and brings a NEW ERA to the field of FO.

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1 Introduction

(Background)

- \mathbb{FO}^k モデル検査: in $\mathcal{O}(\|\varphi\| \cdot n^k)$ time (see, *e.g.*, [18, Proposition 3.1])
- Williams' algorithm: in $\mathcal{O}(2^{\|\varphi\|} \cdot n^\omega)$ time for \mathbb{FO}^3 [20, Corollary 1.3][9, Theorem 7].
(★: $\mathcal{O}(2^{\|\varphi\|} \cdot n^{k-3+\omega})$ time for \mathbb{FO}^k は正しい?)
- CoR モデル検査: in $\mathcal{O}(\|\varphi\| \cdot n^\omega)$ time.

(Contribution)

- Williams' algorithm (for \mathbb{FO}^3 and for \mathbb{FO}^k in PNF) \approx \mathbb{FO}^3 -to-CoR translation.
- For \mathbb{FO}^k , for any $\varepsilon > 0$, there is no algorithm in $\text{poly}(\|\varphi\|) \cdot n^{k-\varepsilon}$ time, under SETH.
 - There is no polynomial-time translation from \mathbb{FO}^3 to CoR, under SETH.
- $\text{poly}(\|\varphi\|) \cdot n^{3-\varepsilon}$ time algorithms for (parameterized) fragments of \mathbb{FO}^3 .
 - “**quantifier-width**”
 - A dichotomy *w.r.t.* signatures

...

YN: TODO: (Below is wip.)

The *model checking problem* is the following problem:

given both a sentence φ and a structure \mathfrak{A} , does $\mathfrak{A} \models \varphi$ hold?

Unfortunately, the *model checking problem* for \mathbb{FO} is complete in PSPACE. Nevertheless, when the number of variables is fixed, the model checking problem for \mathbb{FO}^k can be decided in $\mathcal{O}(\|\varphi\| \cdot n^k)$ time by a naive bottom-up evaluation algorithm (see, *e.g.*, [18, Proposition 3.1]), where $\|\varphi\|$ is the length of φ and n is the number of vertices in \mathfrak{A} . A natural question is

Can the complexity be improved from $\mathcal{O}(\|\varphi\| \cdot n^k)$ time?

Williams [20] gave a positive answer to this question. He showed that when φ is *fixed*, the following holds:



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- 25 ■ For every $k \geq 3$, the model checking problem for \mathbb{FO}^k sentences¹ can be decided in
 26 $\mathcal{O}(n^{k-3+\omega})$ time [20, Corollary 1.3][9, Theorem 7], where ω is the exponent of matrix
 27 multiplication.
- 28 ■ For some $k \geq 3$,² if the model checking problem for \mathbb{FO}^k sentences³ can be decided in
 29 $\mathcal{O}(n^{k-1-\varepsilon})$ time for some $\varepsilon > 0$, then the Strong Exponential Time Hypothesis (SETH)
 30 is false [20, Corollary 1.4].

31 However when φ is not fixed, Williams' algorithm [20] for \mathbb{FO}^3 sentences is $\mathcal{O}(2^{\|\varphi\|} \cdot n^\omega)$
 32 time. The exponential blowup *w.r.t.* $\|\varphi\|$ is due to that the algorithm requires a transformation
 33 to the disjunctive normal form (DNF).

34 ...

35 In this paper, we first revisit Williams' algorithm. We give a perspective from *Tarski's*
 36 *calculus of relations* (henceforth, *CoR*) [16]. ... We can translate \mathbb{FO}^3 sentences into CoR
 37 sentences in $\mathcal{O}(2^{\|\varphi\|})$ time. The original translation is given in [17]. An $\mathcal{O}(2^{\|\varphi\|})$ time
 38 translation is given in [14]; an implementation of this translation is given in [2].⁴ For CoR
 39 sentences, the model checking problem can be decided in $\mathcal{O}(\|\varphi\| \cdot n^\omega)$ time by a naive
 40 bottom-up evaluation algorithm, where we use the *matrix multiplication* algorithm for the
 41 relational composition ($;$). Combining them, we have that the model checking problem for
 42 \mathbb{FO}^3 sentences can be decided in $\mathcal{O}(2^{\|\varphi\|} \cdot n^\omega)$ time.

43 Now, ...

YN: TODO: Parameterized complexity (*e.g.*, [7, 6, 5, 8]).
 Parameterized complexity for space complexity [3]

45 2 Preliminaries

46 We write \mathbb{Z} for the set of all integers.

47 We write $[l, r]$ for the set $\{i \in \mathbb{Z} \mid l \leq i \leq r\}$.

48 A *relational signature* (henceforth, *signature*) σ is ...

49 A *structure* \mathfrak{A} over a signature σ is a tuple $\langle |\mathfrak{A}|, \{R^{\mathfrak{A}}\}_{R \in \sigma} \rangle$, where its *universe* $|\mathfrak{A}|$ is a
 50 finite⁵ set of *vertices* and each $R^{\mathfrak{A}}$ is a binary relation on $|\mathfrak{A}|$.

51 A *sentence* is a formula without free variables.

52 For a formula φ and a *valuation* \mathbf{v} , we write $\mathfrak{A}, \mathbf{v} \models \varphi$ to denote that φ is true at \mathfrak{A} under
 53 \mathbf{v} . For a sentence φ , we write $\mathfrak{A} \models \varphi$ to denote that φ is true at \mathfrak{A} .

54 We write \mathbb{FO}^k to denote the set of formulas of First-order predicate logic with k -variables.

¹ Precisely, Williams [20] shows for \mathbb{FO}^k sentences in *prenex normal form* (so, the number of occurrences of *quantifiers* is at most k) [20, Corollary 1.3]. Later, Gao and Impagliazzo [9, Theorem 7] claims that Williams' algorithm can be extended for general \mathbb{FO}^k sentences.

YN: ($\not\equiv 3$)

² When $k \leq 3$, it is almost clear that the model checking problem cannot be decided in $\mathcal{O}(n^{k-1-\varepsilon})$, as the input adjacent matrix is given by $\Theta(n^2)$ cells. See Proposition 14 for a more precise proof.

³ This claims holds even for \mathbb{FO}^k sentences in $\exists^*\forall$ -prenex normal form.

⁴ Williams' algorithm takes almost the same steps as in the translation in [14]... (TODO: revise the sentences)

⁵ In this paper, we are only interested in finite structures.

55 Important Remark (Unary and Binary Predicates)

56 We allow unary and binary predicates in our \mathbb{FO}^3 . Therefore, we do not allow terms involving
 57 ternary predicates such as $\forall x. \exists y. \forall z. P(x, y, z)$.⁶

58 On this remark, we can identify each model \mathcal{M} as color graphs as follows:

- 59 ■ Each element corresponds a node (vertex).
- 60 ■ Each unary predicate U_i means “color- i ” node.
- 61 ■ Each binary predicate P_j is “color- j ” edge.
- 62 ■ For example, on two nodes x, y , a term $(U_i(x) \wedge U_j(y) \wedge P_k(x, y))$ means that x has the
 63 color i , y has the color j , and there is a color k edge from x to y .

64 Quantifier width

Quantifier widthという謎の指標を導入します。
 与えられた式の部分項 $\forall/\exists x.(\dots)$ について \dots 部分に「直接」出てくるquantifierの
 個数を、この部分項のquantifier widthと呼ぶことにします。
 たとえば、次の部分項

$$\forall x. \left(\left(\underline{\forall y. P(y)} \right) \vee \left(\underline{\forall z. (\forall x. P(z, x))} \wedge Q(z) \right) \right)$$

については、内部に3つのquantifierの出現（下線）があるのですが、直接の出現は赤
 くした2つだけなので、この部分項のquantifier widthは「2」です。
 そして、式のquantifier widthというのは、最大のquantifier widthを持つ部分項におけ
 る、その数として定義します。

66 Tarski's Calculus of Relations (CoR)

67 CoR **terms** are generated by the following grammar:

$$\begin{aligned} 68 \quad t, s &::= a \mid t + s \mid t \cap s \mid t^- \\ 69 \quad &\mid t ; s \mid t \uparrow s \mid t^\pi. \end{aligned}$$

70 ► **Proposition 1.** *The model checking problem for **atomic** CoR formulas is in $\mathcal{O}(\|\varphi\| \cdot n^\omega)$*
 71 *time.*

72 **Proof Sketch.** By a naive bottom-up evaluation algorithm. Here, we use the **matrix multi-**
 73 **plication** algorithm for the **relational composition** (;). ◀

74 SETH

75 ► **Conjecture 2** (The **Strong Exponential Time Hypothesis** (**SETH**) [11, 12]). *For every $\delta < 1$,*
 76 *there exists an integer k such that **k-CNF-SAT** with n variables cannot be solved in $\mathcal{O}(2^{\delta n})$*
 77 *time.*

78 **SETH** implies the following hypothesis.

⁶ For each unary predicate U , we can simulate it by a diagonal binary predicate $B_U(x, y) := x = y \wedge U(x)$.
 However, in our later construction, we use some unary predicates; thus, we also allow unary predicates
 explicitly.

YN: たとえば [19] では、 $\mathcal{O}(2^{\delta n})$ ではなく $2^{\delta n} \text{poly}(n)$ を用いているが、予想としては同値（のはず。 $\text{poly}(n) < \mathcal{O}(2^{\delta n})$ なの）。 $\mathcal{O}(2^n \text{poly}(n))$ のアルゴリズムが存在するという fact に合わせている？

YN: However, it is not known whether the two conjectures are equivalent. [4, Previous Work]

79 ► **Conjecture 3** (e.g., in [15]). *For every $\delta < 1$, CNF-SAT with n variables and m clauses*
 80 *cannot be solved in $2^{\delta n} \cdot \text{poly}(m)$ time.*

81 3 Faster FO3 model checking and the FO3-to-CoR translation

82 YN: TODO: Give an outline of the translation [14].

83 For instance, when φ is the following \mathbb{FO}^3 sentence

$$84 \quad \exists y.((P(x, y) \vee Q(y, z)) \wedge R(x, y)) \wedge S(x, z),$$

85 we can translate φ into the CoR formula as follows:

$$\begin{aligned} 86 \quad & \exists y.((P(x, y) \vee Q(y, z)) \wedge R(x, y)) \wedge S(x, z) \\ 87 \quad & \rightsquigarrow \exists y.(P(x, y) \wedge R(x, y) \wedge S(x, z)) \vee (Q(y, z) \wedge R(x, y) \wedge S(x, z)) \quad (\text{taking the DNF}) \\ 88 \quad & \rightsquigarrow ((\exists y.P(x, y) \wedge R(x, y)) \wedge S(x, z)) \vee ((\exists y.Q(y, z) \wedge R(x, y)) \wedge S(x, z)) \\ & \quad \quad \quad (\text{pushing quantifiers}) \\ 89 \quad & \rightsquigarrow (((P \cap R) ; \top)(x, z) \wedge S(x, z)) \vee ((R ; Q)(x, z) \wedge S(x, z)) \\ & \quad \quad \quad (\text{translating into CoR (intermediate)}) \\ 90 \quad & \rightsquigarrow (((P \cap R) ; \top) \cap S) + ((R ; Q) \cap S)(x, z). \quad (\text{translating into CoR}) \end{aligned}$$

91 ► **Proposition 4** ([14]). *There is an $\mathcal{O}(2^{\|\varphi\|})$ -time translation from an \mathbb{FO}^3 formula φ with*
 92 *at most two free variables into an **atomic** CoR formula **semantically equivalent** to φ .*

93 Hence, combining with Proposition 1, we have obtained Williams' result for $k = 3$, via
 94 the \mathbb{FO}^3 -to-CoR translation.

95 ► **Theorem 5.** *The model checking problem for \mathbb{FO}^3 sentences can be decided in $\mathcal{O}(2^{\|\varphi\|} \cdot n^\omega)$*
 96 *time.*

97 **Proof.** Given an \mathbb{FO}^3 sentence φ , we can translate φ into an **atomic** CoR formula φ' in
 98 $\mathcal{O}(2^{\|\varphi\|})$ time (Proposition 4). The **size** of φ' is also $\mathcal{O}(2^{\|\varphi\|})$. Hence, by Proposition 1, we
 99 have obtained an $\mathcal{O}(2^{\|\varphi\|} \cdot n^\omega)$ time algorithm. ◀

100 By Proposition 1 and Theorem 5, when the input formula is given by CoR instead of \mathbb{FO}^3 ,
 101 the model checking problem is exponentially more efficiently solvable.

102 4 Complexity gap between FO3 and CoR

103 A natural question arising from Theorem 5 is whether there is a faster algorithm without
 104 the exponential blowup with respect to the length of the input sentence. Namely,

105 Can we give an $\text{poly}(\|\varphi\|) \cdot n^{3-\varepsilon}$ time algorithm for the model checking problem for \mathbb{FO}^3 ?

106 In this section, we answer to this question negatively, assuming SETH, as follows.

107 ► **Theorem 6.** *Under SETH, for any $k \geq 1$ and any $\varepsilon > 0$,⁷ the model checking problem for*
 108 *\mathbb{FO}^k sentences cannot be decided in $\text{poly}(\|\varphi\|) \cdot n^{k-\varepsilon}$ time.*

⁷ When $k \leq 2$, the model checking problem cannot be decided in $\text{poly}(\|\varphi\|) \cdot n^{k-\varepsilon}$ time (even when φ is fixed), as the input adjacent matrix is given by $\Theta(n^2)$ cells (Proposition 14, for a more precise proof).

109 Theorem 6 is a direct consequence of the following lemma.

110 ► **Lemma 7.** *Suppose that there are an integer $k \geq 1$, a function f , and an $\varepsilon > 0$ such that*
 111 *the model checking problem for \mathbb{FO}^k sentences is solvable in $\mathcal{O}(f(\|\varphi\|) \cdot n^{k-\varepsilon})$ time. Then*
 112 *CNF-SAT is in $\mathcal{O}(f(\|\varphi\|) \cdot 2^{n(1-\varepsilon/k)})$ time, where $\|\varphi\|$ is the length of the CNF φ and n is*
 113 *the number of variables in φ .*

114 **Proof.** Let φ be a CNF with n variables. *W.l.o.g.*, assume that k divides n . Let p_1, \dots, p_n
 115 be variables in φ . We define \mathfrak{A} as the structure with $2^{n/k}$ vertices and unary relation symbols
 116 $P_1, \dots, P_{n/k}$ such that for every distinct pair $\langle v, w \rangle$ of vertices in \mathfrak{A} , there is some unary
 117 relation symbol P_i such that $v \in P_i^{\mathfrak{A}}$ and $w \notin P_i^{\mathfrak{A}}$ hold or $v \notin P_i^{\mathfrak{A}}$ and $w \in P_i^{\mathfrak{A}}$ hold. We also
 118 define ψ as the following \mathbb{FO}^k sentence:

$$119 \quad \exists z_1 \dots \exists z_k. \varphi',$$

120 where φ' is the φ in which each variable p_{kq+r} (where $0 \leq q < n/k$ and $1 \leq r \leq k$) has been
 121 replaced with the atomic formula $P_q(z_r)$. For instance, when φ is the following CNF with 6
 122 variables $(p_1 \vee p_3) \wedge (p_5 \vee p_2) \wedge (p_4 \vee p_6)$, the \mathbb{FO}^3 sentence ψ is given as follows:

$$123 \quad \exists z_1. \exists z_2. \exists z_3. (P_1(z_1) \vee P_1(z_3)) \wedge (P_2(z_2) \vee P_1(z_2)) \wedge (P_2(z_1) \vee P_2(z_3)).$$

124 For each valuation \mathbf{v} of φ , let \mathbf{v}' be the (unique) valuation so that for all $0 \leq q < n/k$ and
 125 $1 \leq r \leq k$, $\mathbf{v}(p_{kq+r})$ is true iff $\mathbf{v}'(z_r) \in P_q^{\mathfrak{A}}$. Then, φ is true at \mathbf{v} iff $\mathfrak{A}, \mathbf{v}' \models \varphi'$. As the
 126 map above is a bijection, we have that φ is satisfiable iff $\mathfrak{A} \models \psi$. Thus, by applying the
 127 assumption, CNF-SAT is in $\mathcal{O}(f(\|\varphi\|) \cdot (2^{n/k})^{k-\varepsilon})$ time. ◀

128 YN: TODO: Add a note on succinctness.

129 Moreover, the result above still holds even over the signature with exactly one binary
 130 relation symbol.

YN: Theorem 8 uses non-PNF formulas.

131 ► **Theorem 8.** *Under SETH, for any $k \geq 1$ and any $\varepsilon > 0$,⁸ the model checking problem for*
 132 *\mathbb{FO}^k sentences with exactly one binary relation symbol E cannot be decided in $\text{poly}(\|\varphi\|) \cdot n^{k-\varepsilon}$*
 133 *time.*

134 Theorem 8 is a direct consequence of the following lemma.

135 ► **Lemma 9.** *Suppose that there is an integer $k \geq 2$ and an $\varepsilon > 0$ such that the model*
 136 *checking problem for \mathbb{FO}^k sentences with exactly one binary relation symbol E is solvable in*
 137 *$\text{poly}(\|\varphi\|) \cdot n^{k-\varepsilon}$ time. Then CNF-SAT is in $\text{poly}(\|\varphi\|) \cdot 2^{n(1-\varepsilon/k)}$ time, where $\|\varphi\|$ is the*
 138 *length of the CNF φ and n is the number of variables in φ .*

139 **Proof.** Let φ be a CNF with n variables. *W.l.o.g.*, assume that k divides n . Let p_1, \dots, p_n
 140 be variables in φ . We define the structure \mathfrak{A} as follows:

- 141 ■ $|\mathfrak{A}| := (\{V\} \times [0, 2^{n/k} - 1]) \uplus (\{E\} \times [1, n/k]);$ we write v^V for $\langle V, v \rangle$ (similarly for v^E).
- 142 ■ $E^{\mathfrak{A}} := \{ \langle v^V, w^E \rangle \mid v \in [0, 2^{n/k} - 1], w \in [1, n/k], \text{ and the } w\text{-th bit of } v \text{ is } 1 \} \cup \{ \langle w^E, (w +$
 143 $1)^E \rangle \mid w \in [1, n - 1] \} \cup \{ \langle n^E, n^E \rangle \}.$

⁸ When $k \leq 2$, the model checking problem cannot be decided in $\text{poly}(\|\varphi\|) \cdot n^{k-\varepsilon}$ time (even when φ is fixed), as the input adjacent matrix is given by $\Theta(n^2)$ cells (Proposition 14, for a more precise proof).

144 We also define ψ as the following \mathbb{FO}^k sentence:

$$145 \quad \exists z_1 \dots \exists z_k. \varphi',$$

146 where φ' is the φ in which each variable p_{kq+r} (where $0 \leq q < n/k$ and $1 \leq r \leq k$) has been
147 replaced with the \mathbb{FO}^2 formula $X_q(z_r)$ defined as follows:

$$148 \quad X_q(z) := \begin{cases} E(z, z) & \text{if } q = 0, \\ \exists z'. E(z, z') \wedge X_{q-1}(z') & \text{if } q > 0 \text{ where } z' \text{ is a variable not equal to } z. \end{cases}$$

149 Note that $X_q(z_r)$ is **semantically equivalent** to the CoR formula $(E^q(E \cap I))^d(z_r, z_r)$.

150 For instance when $k = 3$, the variable $p_{2 \cdot 1 + 3}$ is transformed to the following formula:

$$151 \quad \exists z'. E(z_3, z') \wedge (\exists z''. E(z', z'') \wedge E(z'', z'')).$$

152 For each **valuation** \mathbf{v} of φ , we can take a **valuation** \mathbf{v}' of φ' so that for all $0 \leq q < n/k$ and
153 $1 \leq r \leq k$, $\mathbf{v}(p_{kq+r})$ is true iff $\mathbf{v}' \models X_q(z_r)$. Then, φ is true at \mathbf{v} iff $\mathbf{v}' \models \varphi'$. Conversely,
154 for each **valuation** \mathbf{v}' of φ' , we can take a **valuation** \mathbf{v} of φ so that φ is true at \mathbf{v} iff $\mathbf{v}' \models \varphi'$.
155 Thus, we have that φ is **satisfiable** iff $\mathbf{v}' \models \psi$. Thus, by applying the assumption, CNF-SAT is
156 in $\text{poly}(n/k \cdot \|\varphi\|) \cdot (2^{n/k} + n/k)^{k-\varepsilon}$ time. As $n \leq \|\varphi\|$, this implies $\text{poly}(\|\varphi\|) \cdot 2^{n(1-\varepsilon/k)}$. ◀

157 **YN: TODO: Add a dichotomy ($\times \in 3$) in some section.**

158 5 FPT on F03

159 We write $\mathbb{FO}_{k,w}^3$ to denote the set of formulas of \mathbb{FO}^3 with k -predicates and w -quantifier-width.

160 ► **Theorem 10** ($\mathbb{FO}_{k,w}^3$ model checkig is FPT).

161 *Let $\varphi : \mathbb{FO}_{k,w}^3$.*

162 *For a given model (graph) \mathcal{G} , we can decide if $\mathcal{G} \models \varphi$ in $O(2^{k+w} \cdot |\varphi| \cdot N^\omega)$ where*

163 **■** *N is the number of vertices in \mathcal{G} ; and*

164 **■** *ω is a complexity constant for the boolean matrix multiplication (BMM): i.e., $O(N^\omega)$ -time
165 is the time complexity for BMM of two (N, N) matrices.⁹*

166 ► **Corollary 11.** *If we fix the number of predicates k and width w , the model checking problem
167 $\mathcal{G} \models \varphi$ in $O(|\varphi| \cdot N^\omega)$ time-complexity where N is the number of nodes in \mathcal{G} .*

k をfixしない、素朴な $\{(x, y, z)\}$ 構築アルゴリズムだと $O(|\varphi| \cdot N^3)$ になって、その
設定だと Theorem 10 は 係数 が大きくなりすぎて遅いということを、意味のある形
で言及した方が良い。

168
169 以降、Theorem 10 の証明をやっていきます。

170 5.1 Proof of Theorem 10

171 Here we use $P, Q, R \dots$ to denote binary predicates and U, \dots for unary predicates.

172 For a given graph (model) \mathcal{G} , we write $N_{\mathcal{G}}$ or simply N to denote the number of vertices
173 of \mathcal{G} .

⁹ 何らかの論文をciteして、今の記録がどれくらいか書いておくと分かりやすい

174 Our idea: Quantifier elimination (Simple case)

175 Our query evaluation strategy is evaluating the innermost (quantifier-free formula) to the
176 outermost along with quantifier eliminating.

Let us consider the following situation:

$$Qx \cdots (Qy \cdots (Qz \cdots (Qx \cdots (Qy \cdots (\exists x.(P(y, x) \wedge R(x, z))))))))$$

To eliminate $\exists z.P(y, x) \wedge R(x, z)$, we introduce a new predicate defined as follows:

$$S(x, y) := \exists z.P(x, z) \wedge Q(z, y).$$

177 It is worth noting that S is just the composed predicate of P and Q . So, we write $P \odot Q$
178 instead of S .

Using this predicate, we can rewrite the above one to the following one:

$$Qx \cdots (Qy \cdots (Qz \cdots (Qx \cdots (Qy \cdots (P \odot R)(y, z))))))$$

179

180 Continuing this process, if we eventually enter $\forall x.\exists y.T(x, y)$:

- 181 1. building a unary predicate $T_y(x) := \exists y.T(x, y)$ in $O(N^2)$ -time, we rewrite it to $\forall x.T_y(x)$.
- 182 2. we then simply evaluate $\forall x.T_y(x)$ as $\bigwedge_{1 \leq i \leq N} T_y(i)$ in $O(N)$ -time.

183 Time-complexity of basic operations

184 Here we enumerate the time-complexity of our basic operations, which will be used later:

- 185 ■ Building a new predicate $Q(x, y) := P(x, y) \bowtie R(x, y)$ ($\bowtie \in \{\wedge, \vee\}$) needs $O(N^2)$ -time.
- 186 ■ Transposing a predicate $P^T(x, y) := P(y, x)$ needs $O(N^2)$ -time.
- 187 ■ Evaluating unary predicates $Qx.P(x)$ ($Q \in \{\forall, \exists\}$) needs $O(N)$ -time.
- 188 ■ Building a new predicate $Q(x) := Qy.P(x, y)$ ($Q \in \{\exists, \forall\}$) needs $O(N^2)$ -time.

189 So, the remaining operator is predicate composing $(P \odot Q)(x, y) := \exists z.P(x, z) \wedge Q(z, y)$
190 from predicates P and Q . Although a very simple matrix multiplication requires $O(N^3)$ -time,
191 we can build such $P \odot Q$ using fast matrix multiplication algorithms.

192 ► **Proposition 12 (Fact).** *From two predicates P and Q , we can build the composited predicate*
193 *$P \odot Q$ in $O(N^\omega)$ -time.*

194 5.2 Quantifier elimination: General case

195 If we encounter terms of the very restricted form like $\exists z.P(x, z) \wedge Q(z, y)$, we can replace it
196 by $(P \odot Q)(x, y)$ with eliminating the quantifier occurrence $\exists z$.

In other words, if we encounter more general form of $\exists z \cdots$, we need to translate it to some adequate form for quantifier elimination. Let us consider the following term:

$$\exists z. \left((P(x, z) \vee P(y, z) \vee P(x, y)) \wedge (Q(x, z) \vee Q(x, y)) \wedge (R(y, z) \vee R(x, y)) \right).$$

197 We cannot directly apply our quantifier elimination strategy because \exists -quantifiers do not
198 distribute on \wedge : i.e., $(\exists z.\varphi_1 \wedge \varphi_2) \neq (\exists z.\varphi_1) \wedge (\exists z.\varphi_2)$.

To apply our quantifier elimination procedure, DNF is an adequate form. However, as well-known, converting-to-DNF translation generates an exponential-large expression in the worst case. For our example, we first translate the first $(\dots) \wedge (\dots)$ subterm as follows:

$$\begin{aligned} & (P(x, z) \vee P(y, z) \vee P(x, y)) \wedge (Q(x, z) \vee Q(x, y)) \\ \Rightarrow & (P(x, z) \wedge Q(x, z)) \vee (P(y, z) \wedge Q(x, z)) \vee (P(x, y) \wedge Q(x, z)) \\ & \vee (P(x, z) \wedge Q(x, y)) \vee (P(y, z) \wedge Q(x, y)) \vee (P(x, y) \wedge Q(x, y)). \end{aligned}$$

199 We continue to normalize this term with $R(y, z) \vee R(x, y)$ and it generates DNF with 12
200 bases of the form $P(_, _) \wedge Q(_, _)$.

201 Since our goal is to design an $O(F(k, w) \cdot |\varphi| \cdot N^\omega)$ -time algorithm, we cannot adopt
202 converting-to-DNF transformations.

203 5.3 Tseytin-like (?) transformation for General Cases

204 As we have seen above, explicit translations to DNF does not work well since it may explode
205 given expression to an exponential size.

206 We here develop a translation inspired by Tseytin's transformation.¹⁰

実際はインスパイアされたという程でもないんですが、とはいえよく知られた構成
な気もするので、何かciteできるものがあればしたいという感じ。

207

Let us revisit the above example:

$$\exists z. \left((P(x, z) \vee P(y, z) \vee P(x, y)) \wedge (Q(x, z) \vee Q(x, y)) \wedge (R(y, z) \vee R(x, y)) \right).$$

First, we generate all possible *assignments* to predicates. For our example, we have the following assignments:

	$P(x, z)$	$P(y, z)$	$P(x, y)$	$Q(x, z)$	$Q(x, y)$	$R(y, z)$	$R(x, y)$
α_1	0	0	0	0	0	0	0
α_2	0	0	0	0	0	0	1
α_3	0	0	0	0	0	1	0
\vdots							
α	1	0	1	0	1	0	0
\vdots							
α_\star	1	1	1	1	1	1	1

208 Let \mathcal{A} be the set of all assignments.

209 Let $\llbracket \Psi \rrbracket_\alpha$ be a boolean value obtained by evaluationg Ψ under an assignment α .

210 For our example expression ψ , $\llbracket \psi \rrbracket_\alpha = 0$ using α appearing in the above table.

211 If we change α as $\alpha(R(y, z)) \leftarrow 1$, then $\llbracket \psi \rrbracket_\alpha = 1$.

212 Introducing assignment leads to the following useful proposition.

► **Proposition 13.**

$$\Psi(x, y, z) \iff \bigvee_{\alpha \in \mathcal{A}} (\alpha \wedge \llbracket \Psi \rrbracket_\alpha).$$

¹⁰https://en.wikipedia.org/wiki/Tseytin_transformation

Epecially, let $\mathcal{V} = \{\alpha \in \mathcal{A} : \llbracket \Psi \rrbracket_\alpha\}$, then we have:

$$\Psi(x, y, z) \iff \bigvee_{\alpha \in \mathcal{V}} \alpha(x, y, z). \quad (\star)$$

By the definition of assignments, α takes the following form:

$$\alpha(x, y, z) \equiv P_1(x, y) \wedge P_1(x, z) \wedge P_2(y, z) \wedge \cdots \wedge P_k(x, z).$$

It means that the right-hand side of (\star) takes DNF. Therefore,

$$\exists z. \Psi(x, y, z) \iff (\exists z. \alpha_1(x, y, z)) \vee (\exists z. \alpha_2(x, y, z)) \vee \cdots \vee (\exists z. \alpha_{|\mathcal{V}|}(x, y, z)).$$

213 It suffices to eliminating an \exists -quantifier from $\exists z. \alpha(x, y, z)$ for an assignment α .

We first divide $\alpha(x, y, z)$ into two parts α_z and α_{-z} as follows:

$$\alpha(x, y, z) \equiv \underbrace{(P_1(x, z) \wedge P_2(z, x) \wedge P_1(y, z) \wedge \cdots)}_{\alpha_z: z\text{-involving-terms}} \wedge \underbrace{(P_1(x, y) \wedge P_1(y, x) \wedge P_2 \wedge \cdots)}_{\alpha_{-z}: z\text{-non-involving-terms}}$$

We now have the following:

$$(\exists z. \alpha(x, y, z)) \iff (\alpha_{-z}(x, y) \wedge (\exists z. \alpha_z(x, y, z))).$$

214 Finally, we transform α_z in the following steps (order):

- 215 1. If α_z has a term $P(z, x)$, we replace it with $\tilde{P}(x, z)$ where $\tilde{P}(a, b) = P(b, a)$.
- 216 2. Similarly, we change terms of the form $P(y, z)$ to $\tilde{P}(z, y)$.
3. At this point,

$$\alpha_z(x, y, z) \equiv \underbrace{(P_1(x, z) \wedge P_3(x, z) \wedge P_4(x, z) \wedge \cdots)}_{\Psi_x} \wedge \underbrace{(P_2(z, y) \wedge P_3(z, y) \wedge P_5(z, y) \wedge \cdots)}_{\Psi_y}.$$

- 217 4. Let Ψ_x (resp. Ψ_y) be the set of predicates appearing in the above former (resp. latter) part.
- 218 5. Let us introduce two new predicates:

$$\pi(x, z) := \bigwedge_{P \in \Psi_x} P(x, z), \quad \lambda(z, y) := \bigwedge_{Q \in \Psi_y} Q(z, y).$$

6. We reach our goal:

$$\alpha_z(x, y, z) = \pi(x, z) \wedge \lambda(z, y).$$

Using the final form, we obtain the following:

$$\begin{aligned} (\exists z. \alpha(x, y, z)) &\iff \left(\alpha_{-z}(x, y) \wedge (\exists z. \pi(x, z) \wedge \lambda(z, y)) \right) \\ &\iff \left(\alpha_{-z}(x, y) \wedge (\pi \odot \lambda)(x, y) \right) \\ &\iff (\alpha_{-z} \wedge (\pi \odot \lambda))(x, y) \end{aligned}$$

219 **5.4 Considering Complexity**

220 To eliminate an \exists -quantifier from an assignment α , we introduce a new single predicate
 221 $\alpha_{-z} \wedge (\pi \odot \lambda)$.

Now we have a question: To eliminate the \exists -quantifier from $\exists z. \Psi(x, y, z)$, from the following characterization

$$\left(\exists z. \Psi(x, y, z) \right) \iff \left(\exists z. \bigvee_{\alpha \in \mathcal{V}} \alpha(x, y, z) \right) \iff \bigvee_{\alpha \in \mathcal{V}} \left(\exists z. \alpha(x, y, z) \right) \iff \bigvee_{\alpha \in \mathcal{V}} (\alpha_{-z} \wedge (\pi \odot \lambda))(x, y)$$

can we avoid to introduce $O(2^k)$ predicates if we have k -predicates?? Please recall that each assignment is a choice from $9k$ -literals from \mathcal{L} where

$$\mathcal{L} = \{Q(a, b) : a, b \in \{x, y, z\}, Q \in \{P_1, P_2, \dots, P_k\}\}.$$

Yes. We can avoid introducing a large number of predicates because we just need to introduce the following single (but large) predicate:

$$(\alpha_{-z}^1 \wedge (\pi^1 \odot \lambda^1)) \vee (\alpha_{-z}^2 \wedge (\pi^2 \odot \lambda^2)) \vee \dots \vee (\alpha_{-z}^n \wedge (\pi^n \odot \lambda^n))$$

222 for $\mathcal{V} = \{\alpha^1, \alpha^2, \dots, \alpha^n\}$.

Let us consider what happens when repeatedly eliminating quantifiers from the innermost to the outermost:

$$\begin{aligned} & \exists y. \left((\exists z. \dots) \bowtie (\exists z. \dots) \bowtie \dots \bowtie (\exists z. \dots) \right) \\ & \Rightarrow \\ & \exists y. \left(\begin{array}{l} \text{Q. how many predicates appear in this scope?} \\ \text{A. } k + w \text{ where } w \text{ is the quantifier width of } \psi \end{array} \right) \\ & \Rightarrow \\ & \text{Eliminating } (\exists y.) \text{ introduces a single new predicate.} \end{aligned}$$

223 Since the required number of quantifier elimination (= the quantifier depth) is bounded
 224 by the size of a given expression $|\psi|$, we obtain the following our theorem in the end.

225 ► **Theorem 10 (Restatement).** Let $\varphi : \mathbb{FO}_{k,w}^3$.

226 For a given model (graph) \mathcal{G} , we can decide if $\mathcal{G} \models \varphi$ in $O(2^{k+w} \cdot |\varphi| \cdot N^w)$.

227 **6 FO2 model checking**

\mathbb{FO}^2 の論理式 ψ と、グラフ \mathcal{G} について $\mathcal{G} \models \psi$ のモデル検査問題は $O(|\psi| \cdot |\mathcal{G}|)$ で解ける気がします。ただし、 $|\mathcal{G}|$ は頂点数ではなく、述語（つまり辺）の定義の記述長も合わせたものです。面白そうなら書いても良いかと思うのですが、 \mathbb{FO}^3 の話だけにしようが、話がぼやけないならなくても良いかなという気がします。

Dense graphについては (x, y) の値の集合を構築する方法で $O(|\psi| \cdot N^2)$ になります。この N は \mathcal{G} の頂点数の方です。ただ、sparse graphの時には N^2 は $|\mathcal{G}|$ について線形ではないので、そこを改良したいという感じです。

アイディアは、 \mathbb{FO}^3 の時と比べると、 $\exists y. P(x, y) \wedge Q(y, x) \simeq (P \odot Q)$ の計算について $(P \odot Q)(x, x)$ の diagonal な形しか要求しないので、行列積を全部やる必要が多分なさそうとかそういう感じです。

Sparse graphの話なので、もしかすると Impagliazzo たちがやっているかもしれません。

YN: [10] Lemma 9.1 Base Case は $O(|\psi| \cdot |\mathcal{G}|)$??

7 Conclusion

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A Tips

A.1 Trivial Results

► **Proposition 14** (Almost trivial fact). *Let φ_0 be the following FO^k sentence:*

$$\exists x_1 \dots \exists x_k. R(x_1, \dots, x_k).$$

Then for any $\varepsilon > 0$, the model checking problem where the sentence is fixed to φ_0 cannot be decided in $\mathcal{O}(n^{k-\varepsilon})$ time (without any assumptions).

Proof. (Below, we suppose that we can access to at most one cell in each step.) Towards a contradiction, assume that for some M , there is an $Mn^{k-\varepsilon}$ time RAM model \mathcal{M} solving the model checking problem (for sufficiently large n). Then, a sufficiently large $n_0 > n$ satisfies $Mn_0^{k-\varepsilon} < n_0^k$.

Let \mathfrak{A} be the structure with n_0 vertices such that $R^{\mathfrak{A}} = \emptyset$. We consider the run ρ of \mathcal{M} when the input is \mathfrak{A} . In the run ρ (of time at most $Mn^{k-\varepsilon}$), as $Mn_0^{k-\varepsilon} < n_0^k$, there is an input cell c for asserting $\langle v_1, \dots, v_k \rangle \in R^{\mathfrak{A}}$ such that ρ does not access c . We then define \mathfrak{B} as the structure \mathfrak{A} in which $R^{\mathfrak{B}}$ has been replaced with the singleton set $\{\langle v_1, \dots, v_k \rangle\}$. When \mathfrak{B} is input, the run is the same as ρ , as the cell c is not accessed in ρ . Hence, in \mathcal{M} , the output of \mathfrak{B} is also the same as that of \mathfrak{A} , reaching a contradiction. (When \mathfrak{A} is input, the output should be “no”. When \mathfrak{B} is input, the output should be “yes”.) ◀

YN: 同様の議論がある論文/本からの引用などで簡単に済ませたい。
とくに 3 項関係記号をもつ FO^3 は、 $\mathcal{O}(n^{3-\varepsilon})$ で解けない (行列積で解けるのは 2 項関係記号の時)。

► **Proposition 15.** *The model checking problem for FO^k with only unary relation symbols (without equality) is solvable in $\mathcal{O}(n)$ time.*

Proof Sketch. In this case, by taking the DNF, we can transform each subformula of the form $\exists x. \varphi$ into the following form: $\exists x. \bigwedge_i P_i(x)$. Thus, each subformula has at most one free variable. Hence, by the naive evaluation algorithm, we have obtained $\mathcal{O}(n)$ time. ◀

A.2 Summary

Problem / max. arity k	1	2	3
PNF FO^3 (data)	$\mathcal{O}(n)$ (Proposition 15)	$\mathcal{O}(n^\omega)$ [20]	$\mathcal{O}(n^3)$ (trivial)
FO^3 (data)	no $\mathcal{O}(n^{k-\varepsilon})$ (Proposition 14)		
PNF FO^3 (combined)	$\mathcal{O}(\ \varphi\ \cdot n^3)$ (trivial)	$\mathcal{O}(\ \varphi\ \cdot n^3)$ (trivial)	$\mathcal{O}(\ \varphi\ \cdot n^3)$ (trivial)
FO^3 (combined)	$\mathcal{O}(2^{\ \varphi\ } \cdot n)$ (Proposition 15)	$\mathcal{O}(2^{\ \varphi\ } \cdot n^\omega)$ [20, 14]	
	no poly $(\ \varphi\) \cdot n^{3-\varepsilon}$, under SETH (Theorem 6)		

(以下おまけ)

Problem / max. arity k	1	2	3	4
PNF FO^4 (data)	$\mathcal{O}(n)$ (Proposition 15)	$\mathcal{O}(n^{1+\omega})$ [20, Cor. 1.3] (?? for FO^4)	$\mathcal{O}(n^{\omega_3})$	$\mathcal{O}(n^4)$ (trivial)
FO^4 (data)	no $\mathcal{O}(n^{1-\varepsilon})$ (Proposition 14)	$\mathcal{O}(n^{\omega_3})$	no $\mathcal{O}(n^{k-\varepsilon})$ (Proposition 14)	
PNF FO^4 (combined)	$\mathcal{O}(\ \varphi\ \cdot n^4)$ (trivial)	$\mathcal{O}(\ \varphi\ \cdot n^4)$ (trivial)	$\mathcal{O}(\ \varphi\ \cdot n^4)$ (trivial)	$\mathcal{O}(\ \varphi\ \cdot n^4)$ (trivial)
FO^4 (combined)	$\mathcal{O}(2^{\ \varphi\ } \cdot n)$ (Proposition 15)	$\mathcal{O}(2^{\ \varphi\ } \cdot n^{1+\omega})$ [20, Cor. 1.3] (?? for FO^4)	$\mathcal{O}(2^{\ \varphi\ } \cdot n^{\omega_3})$	
		$\mathcal{O}(2^{\ \varphi\ } \cdot n^{\omega_3})$		
	no poly $(\ \varphi\) \cdot n^{4-\varepsilon}$, under SETH (Theorem 6)			

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Problem / max. arity k	1	2	3	4	5
PNF \mathbb{FO}^5 (data)	$\mathcal{O}(n)$ (Proposition 15)	$\mathcal{O}(n^{2+\omega})$ [20, Cor. 1.3] (?? for \mathbb{FO}^5) $\mathcal{O}(n^{1+\omega_3})$ (?? for \mathbb{FO}^5) $\mathcal{O}(n^{\omega_4})$	$\mathcal{O}(n^{1+\omega_3})$ [20, Cor. 1.3] (?? for \mathbb{FO}^5) $\mathcal{O}(n^{\omega_4})$	$\mathcal{O}(n^{\omega_4})$	$\mathcal{O}(n^5)$ (trivial)
\mathbb{FO}^5 (data)	no $\mathcal{O}(n^{1-\varepsilon})$ (Proposition 14)	no $\mathcal{O}(n^{1-\varepsilon})$ under SETH [20, Cor. 1.4]		no $\mathcal{O}(n^{1-\varepsilon})$ (Proposition 14)	
PNF \mathbb{FO}^5 (combined)	$\mathcal{O}(\ \varphi\ \cdot n^5)$ (trivial)	$\mathcal{O}(\ \varphi\ \cdot n^5)$ (trivial) $\mathcal{O}(2^{\ \varphi\ } \cdot n^{2+\omega})$ [20, Cor. 1.3] (?? for \mathbb{FO}^5) $\mathcal{O}(2^{\ \varphi\ } \cdot n^{1+\omega_3})$ (?? for \mathbb{FO}^5) $\mathcal{O}(2^{\ \varphi\ } \cdot n^{\omega_4})$	$\mathcal{O}(\ \varphi\ \cdot n^5)$ (trivial) $\mathcal{O}(n^{1+\omega_3})$ (?? for \mathbb{FO}^5) $\mathcal{O}(2^{\ \varphi\ } \cdot n^{\omega_4})$	$\mathcal{O}(\ \varphi\ \cdot n^5)$ (trivial) $\mathcal{O}(2^{\ \varphi\ } \cdot n^{\omega_4})$	$\mathcal{O}(\ \varphi\ \cdot n^5)$ (trivial)
\mathbb{FO}^5 (combined)	$\mathcal{O}(2^{\ \varphi\ } \cdot n)$ (Proposition 15)	no poly $(\ \varphi\) \cdot n^{\omega-\varepsilon}$, under SETH (Theorem 6)			

YN:

■ ?? は主張が不明/怪しい箇所（メモ2参照）

- 主張が成り立たなそうであることを言うには、
Binary \mathbb{FO}^4 is in $\mathcal{O}(n^{4-\varepsilon})$ time $\implies \omega_3 < 4$ を言えればよさそう。

► **Observation 16.** * $\omega_2 = \omega < 2.38 (< 3)$.

* $\omega_3 < 4$ is wished [20] but open.

4-clique in a 3-uniform hypergraph を含む難しい問題のようです。Cf. HyperClique Conjecture

- もしくは、もう少し精密化して今回の設定に合った行列積の問題に一般化する。
入力を $[n]$ 上の任意有限個の m 項関係として、 k 引数関係 A_1, \dots, A_k は入力の m 項関係から \mathbb{FO}^{k+1} -definable であるとする。この時の“一般化行列積”

$$D[i_1, \dots, i_k] = \bigvee_{\ell=1}^n \bigwedge_{j=1}^k A_j[i_1, \dots, i_{j-1}, i_{j+1}, \dots, i_k, \ell]$$

の時間計算量を $\mathcal{O}(n^{\omega_{k,m}})$ とする。

* Fact: $\omega_{k,k} = \omega_k$ 。とくに $\omega_{2,2} = \omega$ 。

($\omega_{k,k} \leq \omega_k$: \mathbb{FO}^{k+1} model checking が $\mathcal{O}(n^{\omega_k})$ より。 $\omega_{k,k} \geq \omega_k$: A_1, \dots, A_k が入力になる場合を考える)

* Fact: $\omega_{k,1} = k$ (成分ごと独立なので)

* Fact: $k = \omega_{k,1} \leq \omega_{k,2} \leq \dots \leq \omega_{k,k} = \omega_k \leq k+1$

* Fact: $\omega_k = k+1$, under k -uniform HyperClique Hypothesis [1], for $k \geq 3$

\rightsquigarrow (??) $\omega_{k,2} = k+1$, under “Strong” k -uniform HyperClique Hypothesis, for $k \geq 3$

みたいなことを言っておいても許される??

We then have:

► **Proposition 17.** Binary \mathbb{FO}^4 is in $\mathcal{O}(n^c)$ time $\iff \omega_{3,2} \leq c$.

Proof. \Leftarrow : By an analog of the FO3-to-CoR translation. \Rightarrow : Every matrix multiplication problem w.r.t. $\omega_{3,2}$ can be expressed as a model checking problem where the sentence is fixed to a binary \mathbb{FO}^4 sentence. ◀

► **Corollary 18.** If binary \mathbb{FO}^4 is in $\mathcal{O}(n^{1+\omega})$ time, then $\omega_{3,2} \leq 1 + \omega$. Hence, $\omega_{3,2} < 4$.

\rightsquigarrow “Strong” k -uniform HyperClique Hypothesis に反する

$\rightsquigarrow \omega_{3,2}$ の問題の形で書けるが $\mathcal{O}(n^{4-\varepsilon})$ が open な問題はあるでしょうか?

■ 赤の箇所は未証明

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B memo

1. 2025/06/25 メールより

あとは、FO3 が $O(2^{\delta \|\varphi\|} n^\omega)$ time ($\delta < 1$) で解けないことを示せるか？という問題も考えられそうです。（これが示せれば $O(2^{\delta \|\varphi\|})$ time の FO3-to-CoR の変換がないこともいえます）

(Open Problem)

2. 2025/07/04 メールより

YN:

[20, Theorem 1.3] の大枠の方針は、CoR変換のコアの部分と似た感じで、 $\exists y.P(\vec{x}, y) \wedge Q(y, \vec{z})$ の形に持って行って、*最も内側の* \exists に対して行列積するというものです。ここで \vec{x} と \vec{z} には以下のような条件を課しておきます。

- 共通の変数を持たない
- 次元はおおよそ $(k-1)/2$ ずつに分割されている

この方針を一般に拡張しようとすると、たとえば

$$\exists y.P(x_1, x_2, y) \wedge Q(x_2, x_3, y) \wedge R(x_3, x_2, y)$$

（ただし、各 $P(x_1, x_2, y)$ はたとえば $\exists z.P'(x_1, z) \wedge P'(x_2, z) \wedge P'(z, y)$ のような式）
 みたいな式をどうやって行列積にしますか？という問題が発生しそうです。
 自由変数が高々2の場合（FO3の場合）には、外側の量子子の自由変数を除くと自由変数が高々1つなので、綺麗な分割ができますが、そうでなかった場合には、自由変数間の依存が生じるので、行列積に持っていけないのでは？という気がしています。# 逆に言うとこの辺りの計算量をもう少し厳密に考える価値がありそう？

★ $k \geq 4$ の時に以下は本当に成り立つか？ (cf. [9, Theorem 7])

- the model checking problem for \mathbb{FO}^k sentences ~~in prenex normal form~~ can be decided in $n^{k-1+o(1)}$ time for $k \geq 9$.
- the model checking problem for \mathbb{FO}^k sentences ~~in prenex normal form~~ can be decided in $O(n^{k-3+\omega})$ time for $k \geq 3$.
 - 定数置き換えの方針で、たとえば上の論理式で x_3 を定数としても $P(x_1, x_2, y)$ によって行列積にならない (cf. Proposition 14)
 - 下の“関連”の方針から高々 $k-1$ 項関係であれば $O(n^{\omega_{k-1}})$ time は言えるはず。
 - * “ k 次元行列積”をもつ “ k -ary CoR” を考えると同じ議論。
 - * したがって、 $\omega_{k-1} < k$ の仮説 (?) のもと $O(n^{k-\varepsilon})$ time for some $\varepsilon > 0$.

関連 (\mathbb{FO}^4 が3項関係をもつ場合、 $O(2^{\|\varphi\|} n^{\omega_3})$ time。ただし ω_k は以下の“ k 次元行列積”の時間計算量 ($k-1 \leq \omega_k \leq k$) ($\omega_2 = \omega$ 。 $\omega_k < 1$ を満たすかどうかは $k \geq 3$ で open (?))

- [20, Open Problem 2 (p.5)]

“What about for vocabularies with ternary relations? ... We wish to compute the following 3D matrix in $n^{4-\varepsilon}$ time, for some $\varepsilon > 0$:

$$D[i, j, k] = \sum_{\ell=1}^n A[i, j, \ell] \cdot B[j, k, \ell] \cdot C[k, i, \ell].$$

... would allow us to more efficiently find (for example) a 4-clique in a 3-uniform hypergraph, which can be expressed as a first-order sentence over a ternary relation.”

- [13]

“... no efficient classical algorithm for 4-clique finding has been discovered so far.”

3. 2025/6/27 メールより

“

(補足) 結果を組み合わせると、以下の dichotomy を示せそうです (関係記号の個数, 定数記号の個数, 等号の有無を考えています)

【冠頭標準形の FO3 の場合】

- 一般 (*) : $\text{poly}(|\psi|)N^{3-\varepsilon}$ だと SETH に反する ← ② と その修正 (*)
厳密には、1 引数関係記号を ω 個含む場合 もしくは 定数記号を ω 個含む場合 (等号あり) もしくは 2 引数関係記号 1 個以上含むかつ定数記号を ω 個含む場合
- 関係記号の個数 k 固定 : $O(|\psi|N^\omega)$ ← 上里さんの方針
- 1 引数関係記号 k 個 かつ 定数記号 ω 個 (等号なし) : $O(|\psi|N^\omega)$ ← 上里さんの方針 (+ 定数記号を先に除去)

【一般の FO3 の場合】

- 一般 (*) : $\text{poly}(|\psi|)N^{3-\varepsilon}$ だと SETH に反する ← ② と その修正 (*)
厳密には、1 引数関係記号を ω 個含む場合 もしくは 定数記号を ω 個含む場合 (等号あり) もしくは 2 引数関係記号 1 個以上含む場合
- 1 引数関係記号 k 個 かつ 定数記号 ω 個 (等号なし) : $O(|\psi|N^\omega)$ ← 上里さんの方針 (+ 定数記号を先に除去 + DNF 変換して量子子内側入れの後の論理式 $\exists x \bigwedge_i \alpha_i$ (各 α_i は $P(x)$ or $\neg P(x)$) が有限通り)
- 1 引数関係記号 k 個 かつ 定数記号 k' 個 (等号あり) : $O(|\psi|N^\omega)$ ← 上里さんの方針 (+ DNF 変換して量子子内側入れの後の論理式 $\exists x \bigwedge_i \alpha_i$ (各 α_i は $P(x), \neg P(x), \neg x = y, \neg x = z, \text{ or } \neg x = c$) が有限通り)

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~~ 「定数記号を ω 個含む場合 (等号あり)」は結論自体誤り。¹¹

~~ ひとまず定数記号を考えず、関係記号の個数と等号の有無を考えるなら以下

【冠頭標準形の FO3 の場合】

- 一般 (*) : $\text{poly}(|\psi|)N^{3-\varepsilon}$ だと SETH に反する ← ② と その修正 (*)
厳密には、1 引数関係記号を ω 個含む場合 (等号なし)
- 関係記号の個数 k 固定 (等号あり) : $O(|\psi|N^\omega)$ ← 上里さんの方針

【一般の FO3 の場合】

- 一般 (*) : $\text{poly}(|\psi|)N^{3-\varepsilon}$ だと SETH に反する ← ② と その修正 (*)
厳密には、1 引数関係記号を ω 個含む場合 (等号なし) もしくは 2 引数関係記号 1 個以上含む場合 (等号なし)
- 1 引数関係記号 k 個 (等号あり) : $O(|\psi|N^\omega)$ ← 上里さんの方針 (+ DNF 変換して量子子内側入れの後の論理式 $\exists x \bigwedge_i \alpha_i$ (各 α_i は $P(x), \neg P(x), \neg x = y, \neg x = z$) が有限通り)

¹¹ この時、各変数の割当は、定数が指す箇所 $\|\psi\|$ 通り + 定数が指さない場合 1 通りのみを考えれば十分なので $\text{poly}(\|\psi\|)$ time。 (構造のサイズに依存しない)