

# Magic States Distillation Using Quantum LDPC Codes.

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## 1 Current Status.

1. Section 5 - Correct. In any CSS code, one can find a large subspace  $\Lambda \subset C_X$  with a dimension that is linear in  $n$  and this subspace also satisfies the required relation for distillation. Specifically, for any  $x \in \Lambda, y, z \in C_X$ , it holds that  $xy = 0$  and  $xyz = 0$ .
2. Sections 6 and 7 - Incorrect. Initially, I believed that assuming the code is LDPC, one could encode the state  $C_Z^\perp$  in constant depth. However, this idea turned out to be incorrect both in calculation and in contrast to the fact that synthesizing the ground state of the Toric code requires  $\Omega(\log n)$  depth.

## 2 Classic Codes With Few Checks.

**Claim 2.1.** *There is a family of classic binary codes, with positive rate,  $\Theta(n^{\frac{1}{3}})$  distance, and  $\gamma n^{\frac{1}{3}}$  checks.*

*Proof.* We are going to show the existences of bipartite expander, over  $n$  left vertices and  $\gamma n^{\frac{1}{3}}$  right vertices such that for any  $S \subset L$  at size at most  $\alpha n^{\frac{1}{3}}$ , the neighbors of  $S$  is at size at least  $\beta|S|$ . We use the standard probabilistic 'fusion construction', meaning that we are going to sample permutation from  $[n \times d_1]$  to  $[n^{\frac{1}{3}} \times d_2]$  and fuse together  $d_1$ 's left vertices subsets  $\{d_1 \cdot j, d_1 \cdot j + 1, d_1 \cdot j + 2, \dots, d_1 \cdot (j + 1) - 1\}$  and similarly fuse together  $d_2$  right vertices.

Now observes that the probability of neighbors  $S \subset L$  being contained in  $T \subset R$  is at most:

$$\Pr[X_{S,T}] \leq \frac{|T|d_2 \cdot (|T|d_2 - 1) \cdots (|T|d_2 - |S|d_1)}{nd_1 \cdot (nd_1 - 1) \cdots (nd_1 - |S|d_1)} \leq \left( \frac{|T|d_2}{nd_1 - |S|d_1} \right)^{|S|d_1}$$

And for the  $|T| < \beta|S|$  the above is lower than:

$$\Pr[X_{S,T}] \leq \left( \frac{2\beta|S|d_2}{nd_1} \right)^{|S|d_1}$$

By the union bound we get that the probability that there exist  $S$  at size  $|S| < \alpha n^{\frac{1}{3}}$  such that the

neighbors of  $S$  is at size less than  $\beta|S|$  is bounded by:

$$\begin{aligned}
\Pr \left[ \bigcup_{\substack{|S| < \alpha n^{\frac{1}{3}} \\ |T| < \beta|S|}} X_{S,T} \right] &\leq \sum_{\substack{|S| < \alpha n^{\frac{1}{3}} \\ |T| < \beta|S|}} \Pr [X_{S,T}] \\
&\leq \sum_{k \geq 1}^{\alpha n^{\frac{1}{3}}} \binom{n}{k} \binom{\gamma n^{\frac{1}{3}}}{\beta k} \cdot \left( \frac{2\beta k d_2}{n d_1} \right)^{k d_1} \\
&\leq \sum_{k \geq 1}^{\alpha n^{\frac{1}{3}}} \left( \frac{e^{2+\beta}}{k} \cdot \frac{n^{1+\beta/3}}{\beta^\beta k^\beta} \cdot \left( \frac{2\beta k d_2}{n d_1} \right)^{d_1} \right)^k \\
&= \sum_{k \geq 1}^{\alpha n^{\frac{1}{3}}} \left( \frac{e^{2+\beta}}{k} \cdot \frac{n^{1+\beta/3}}{\beta^\beta k^\beta} \cdot \left( \frac{2\beta k n^{2/3} d_1}{n d_1} \right)^{d_1} \right)^k \\
&= \sum_{k \geq 1}^{\alpha n^{\frac{1}{3}}} \left( \frac{e^{2+\beta} (2\beta)^{d_1}}{\beta^\beta} \cdot \frac{k^{d_1-\beta-1}}{n^{d_1/3-\beta/3-1}} \right)^k \\
&\leq \sum_{k \geq 1}^{\infty} \left( \frac{e^{2+\beta} (2\beta)^{d_1}}{\beta^\beta} \cdot \gamma^{d_1-\beta/3-\frac{1}{3}} \right)^k = \frac{1}{\varepsilon} - 1
\end{aligned}$$

So one can find parameters such that the probability is strictly less than 1 meaning that with positive probability we sample our desirable bipartite expander graph.  $\square$

The idea form here, set  $C_Z$  to be the above code, and  $C_X^\perp = \emptyset$ , clearly  $C_X^\perp \subset C_Z$ . Now, construct  $\Lambda$  by the method in section 6. The additional phase that we get by applying the gate  $T^n$  corresponds to controlled-S, controlled-controlled-Z between the generators of  $C_Z^\perp$ . So one can fix them by applying at most  $\Theta((\gamma n^{\frac{1}{3}})^3)$  perfect  $T$  gates. In total, if we show a decoder that against noise  $p$  succeeds to correct with probability  $q$ , Then we got a distillation protocol that consume  $\gamma^3 n$  perfect magic states,  $n$  noisy magic states at error rate  $p$  and with probability  $q$  distillate  $pn$  magic states.

$$\langle \gamma^3 n, n, p \rangle \rightarrow \langle 0, pn, q \rangle$$

### 3 Bipartite Random Constructions, Collisions Number.

Let  $u, v \in R$  be checks vertices (right vertices) and let  $Y_{u,v}$  the indicator for a bit been checked by both  $u$  and  $v$  checks, Means that there is a left vertex  $w$  adjances to both  $u, v$ .

$$\begin{aligned}
\Pr [Y_{u,v} = 0 | N(v)] &\geq \frac{(nd_1 - d_2 d_1)(nd_1 - d_2 d_1 - 1) \cdots (nd_1 - d_2 d_1 - d_2)}{nd_1 \cdot (nd_1 - 1) \cdots (nd_1 - d_2)} \\
&= \prod_{i \in d_2} 1 - \frac{d_2 d_1}{nd_1 - i} \geq \left( 1 - \frac{d_2 d_1}{nd_1 - d_2} \right)^{d_2} \geq 1 - d_2 \frac{d_2 d_1}{nd_1 - d_2} \\
&\geq 1 - 2 \frac{d_2^2}{n}
\end{aligned}$$

Thus, the expectation for collision between  $u$  and  $v$  and expected total collisions number are less than:

$$\begin{aligned}\mathbf{E}[Y_{u,v}] &\leq 2 \frac{d_2^2}{n} \\ \mathbf{E}\left[\sum_{u,v \in R} Y_{u,v}\right] &\leq \binom{n \frac{d_1}{d_2}}{2} 2 \frac{d_2^2}{n} \leq n d_1^2 \\ &\sim n \cdot \binom{d_1}{2} \sim m d_2 d_1\end{aligned}$$

Another direction, assume that we choose the adjacency matrix by picking function such that the probability of overlapping is  $\sim \frac{\gamma}{n}$ . And then:

$$\mathbf{E}[Y] \sim m^2 \cdot \frac{\gamma}{n} = \frac{d_1}{d_2} m \gamma = \left(\frac{d_1}{d_2}\right)^2 \gamma n$$

So, for any 'subset' of  $m$  edges use another hash function, (first labaling  $R$  with random labels, and then draw  $h \sim \mathcal{H}$ ). The probability for collision, by the union bound is lower than  $d_2 \cdot \frac{1}{n}$ . Therefore if we take  $d_2$  to behave like  $\sim d_1^5$ , then:

$$\mathbf{E}[Y] \sim \left(\frac{d_1}{d_2}\right)^2 \gamma n \sim \left(\frac{d_1}{d_2}\right)^2 d_2 n = \frac{n}{d_1^3}$$

## 4 Decoding The Code.

Consider the simplest decoding procedure, any right vertex in  $R$  decode it's local view using decoder  $D$ . The probability of failure is lower than the probability that there is a single vertex which fails to decode.

$$\begin{aligned}\Pr[\text{failure}] &\leq \Pr[\exists v \in R \text{ which fails to decode}] \\ &\leq m \cdot \Pr[|e(\Gamma(v))| \geq \mu] \leq m \cdot \Pr[|e(\Gamma(v))| \geq (1 + \theta) p d_2] \\ &\leq m \cdot e^{-\frac{\theta^2}{2 + \theta} d_2 p}\end{aligned}$$

## 5 Candidate For Triorthogonal LDPC Code.

**Claim 5.1.** Consider the ring  $\mathbb{F}_q[x]$  where  $q$  is a prime number. Let  $\Delta = 4^c$  where  $c \geq 3$ . Then we have:

$$\sum_{x \in [\Delta]} x^i =_{\Delta} \in \{0, \Delta/2\}$$

*Proof.* By induction on  $c$ .

1. Base. For  $c = 3$  we computes the summation brutforcely.
2. Assumption. Assume the correctness of the claim for  $c - 1$ .
3. Step. Denote by  $B_j(\Delta)$  the bucket  $\Delta \cdot j + 1, \Delta \cdot j + 2, \dots, \Delta \cdot (j + 1) - 1$ . Observes that:

$$\sum_{x \in B_{j+1}(\Delta)} x^i =_{\Delta} \sum_{x \in B_{j+1}(\Delta)} (x - \Delta)^i =_{\Delta} \sum_{x \in B_j(\Delta)} x^i$$

On the other hand, by the induction assumption, there is some integer  $a$  for which:

$$\sum_{x \in B_1(\Delta/4)} x^i = \Delta/8 \cdot a$$

Thus the summation over  $\Delta$  elements equals to:

$$\sum_{x \in [\Delta]} x^i = \sum_{j \in [4]} \sum_{x \in B_j(\Delta/4)} x^i = \Delta/8 \cdot a \cdot 4 = \Delta \cdot a/2$$

□

**Definition 5.1.** Let  $G = (L, R, E)$  be a bipartite graph, and let  $\Delta$  be an integer. Define  $G'$  to be the graph:  $G' = (\Delta \times L, R, E')$  defined as follows:

$$E' = \{ \{ (i, v), u \} : i \in [\Delta], \{u, v\} \in E \}$$

In addition, we define the equivalence relation  $u \sim v$  for  $u, v \in \Delta \times L$  to hold if the first coordinates of  $u$  and  $v$  are equal.

Let  $G'$  be a graph constructed as described above. Consider the code  $C$  over the  $\mathbb{F}_q$  alphabet, defined as all the assignments of symbols from  $\mathbb{F}_q$  to the  $\Delta \times L$  vertices. Such any vertex on the right side of  $G$  sees a polynomial of degree at most  $d$  on its local view, in addition the  $x$ 's value of bit in  $\Delta \times L$  is the same module  $\Delta$  for all the checks. To clarify, if one checks, treat  $u \in \Delta \times L$  as the value of the polynomial at coordinate  $z$ , and treat the other check as the value of the polynomial at coordinate  $z'$ , then  $z =_{\Delta} z'$ .

**Claim 5.2.**  $C$  is a good LDPC code. (If  $G$  is expander graph).

*Proof.* We obtain a lower bound on the code dimension by subtracting restrictions. So,

$$\dim C = \Delta \cdot |L| - |R| \cdot (1 - \rho) \cdot q$$

Now, assume trough contradiction that there is  $x \in C$  at weight  $|x| < \gamma n$  denote by  $S' \subset \Delta \times L$  the set of vertices setted to a non-trivial symbol. And observes that in the original graph  $G$ ,  $S'$  induce a set of vertices  $S$  by taking the delegations of the equivalence classes.

Since  $G$  is a  $(n, m, \gamma, \alpha)$  expander, and  $|S| < |S'| < \gamma n$ , it followees that  $|\Gamma(S)| > \alpha |S| \Rightarrow$

$$\begin{aligned} |S|/|\Gamma(S)| &< \frac{1}{\alpha} \\ \Rightarrow |S'|/|\Gamma(S)| &< \frac{\Delta}{\alpha} \end{aligned}$$

So there is a check that sees a local view at weight less than  $\frac{\Delta}{\alpha}$  bits. (Otherwise,  $|S'| > |\Gamma(S)| \cdot \frac{\Delta}{\alpha}$ ). So, if  $\frac{\Delta}{\alpha}$  is lower than  $C_0$  distance we get a contradiction. □

**Claim 5.3.** Let  $h_1, h_2, h_3$  be arbitrary checks of  $C$ , not necessarily different. Then:

$$\begin{aligned} h_1 h_2 &=_{\Delta} 0 \\ h_1 h_2 h_3 &=_{\Delta} 0 \end{aligned}$$

*Proof.* Complete it. □

Consider the Tanner **Graph**, such that the graph  $G$  is bipartite, and every two checks overlap on the  $i$ th bucket,  $\Delta$ -size, bits. So for any two checks, we have that

$$\begin{aligned} \sum_{x=i \cdot \Delta}^{(i+1)\Delta} x^j &=_{\Delta} \sum_{x'=(i-1) \cdot \Delta}^{i\Delta} (x' + \Delta)^j \\ &=_{\Delta} \sum_{x=(i-1) \cdot \Delta}^{i\Delta} x'^j = \sum_{x \in \mathbb{F}_{\Delta}} x^j \\ &= \sum_{x \in \mathbb{F}_{\Delta}} (x + a\Delta)^i (x + b\Delta)^j = \sum_{x \in \mathbb{F}_{\Delta}} x^{i+j} \end{aligned}$$

So it's left to show that if we take the bipartite graph to be an expander graph then we have a good code.

Let  $G$  be a bipartite graph  $G = (L, R, E)$  that is a  $(n, m, \gamma, \alpha)$  expander. This means that for any subset  $S \subset V(G)$  with  $|S| < \gamma n$ , the size of the group of neighbors of  $S$  is at least  $\Gamma(|S|) > \alpha|S|$ . Consider the graph  $G' = (\Delta \times L, R, E')$  defined as follows:

$$E' = \{\{(i, v), u\} : i \in [\Delta], \{u, v\} \in E\}$$

Thus for any  $S \subset \Delta \times L$  if  $|S|/\Delta < \gamma n$  we have that:  $\Gamma'(S) < \Gamma(|S|/\Delta)$ .

Therefore, if  $S$  is the set of vertices associated with the non-trivial symbols induced by the assignment of a codeword on the vertices, then if  $|S| < \gamma n$ , we have:

$$\frac{|S|}{\Gamma'(|S|)} \leq \frac{|S|}{\Gamma(|S|/\Delta)} \leq \frac{\Delta}{\alpha}$$

So there is a check that sees on his local view less than  $\Delta/\alpha$  non-trivial bits  $< d(C_0)$ .

## 6 Hyperproduct Code of two Triorthogonal Codes.

Suppose that  $H$  is a parity check matrix such that  $h_i h_j = \Delta \in \{\Delta, \Delta/2\}$  for any two rows. Is that true that the same property holds for the following check matrix?

$$H' \leftarrow [H \otimes I | I \otimes H]$$

$$H'_i H'_j = \overbrace{(H \otimes I)_i (H \otimes I)_j}^A + (I \otimes H)_i (I \otimes H)_j$$

Denote  $i = (i_1, i_2)$  and  $j = (j_1, j_2)$ . So:

$$(H \otimes I)_i (H \otimes I)_j = \delta_{i_2, j_2} H_{i_1} H_{j_1}$$

and

$$(I \otimes H)_i (I \otimes H)_j = \delta_{i_1, j_1} H_{i_2} H_{j_2}$$

## 7 The problem with the above.

The code that is obtained by the polynomial tanner is (almost) self dual code, module  $\Delta$  the multiplication  $x \cdot x$  belongs to  $\{0, \Delta/2\}$ . While what we actually want to have is  $x \cdot x =_4 1$ . An idea how to correct that, sets the checks such only two of them don't commute. After taking the Hyperproduct code, they will turned to  $\Theta(\sqrt{n})$  that don't commute. So if we have a perfect  $\Theta(\sqrt{n})$  T states, we can cancel their phase before the encoding.

Let  $B$  be the bucket which matches  $\{2, 3, \dots, \Delta - 1\}$ . On that bucket, the multiplication of the checks corresponds to  $\sum_{x \in \mathbb{F}_\Delta} x^i - 1^i$ , which is  $\in \{-1, \Delta/2 - 1\}$ . On the otherhand, the codeword  $\xi$  that corresponds to the constant function  $f(x) = 1$  in every bucket gives  $\xi \cdot \xi =_\Delta -1$ .

So  $\xi' = \xi \otimes I$  padding with zeros, is a codeword of the Hyperproduct code, such that  $\xi' \cdot \xi' = 1$ .

## 8 Good Codes With Large $\Delta$ .

**Claim 8.1.** *Let  $v_1, v_2, \dots, v_k$  vectors in  $\mathbb{F}_2^n$ , then there are  $u_1, u_2, \dots, u_{k'}$  for  $k' > k/2$ . Such  $\text{span}\{u_1, u_2, \dots, u_{k'}\} \subset \text{span}\{v_1, v_2, \dots, v_k\}$  and for any  $i, j$  it holds that  $u_i u_j = 0$ .*

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1 Let  $J \leftarrow \emptyset$ 
2 for  $i \in [k/2]$  do
3    $J \leftarrow J \cup \{v_{2i-1}, v_{2i}\}$ 
4   for  $S \subset J$  do
5     Compute the vector  $m_S$ 
6     define as  $m_{S,j} = u_j \sum_{w \in S} w$ 
7   end
8   Pick  $S$  such  $m_S = 0$  and set
9      $u_i \leftarrow \sum_{w \in S} w$ 
10  Choose randomly  $w \in S$  and set
11     $J \leftarrow J/w$ 
12 end
13 : Find commuted vectors  $u_1, u_2, \dots, u_{k'}$ 

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1 Let  $J \leftarrow \emptyset$ 
2 for  $i \in [k/3]$  do
3    $J \leftarrow J \cup \{v_{3i-2}, v_{3i-1}, v_{3i}\}$ 
4   for  $S \subset J$  do
5     Compute the vector  $m_S$ 
6     define as
7        $m_{S,j,j'} = u_{j'} u_j \sum_{w \in S} w$ 
8   end
9   Pick  $S$  such  $m_S = 0$  and set
10     $u_i \leftarrow \sum_{w \in S} w$ 
11  Choose randomly  $w \in S$  and set
12     $J \leftarrow J/w$ 
13 end
14 : Find commuted vectors  $u_1, u_2, \dots, u_{k'}$ 

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*Proof.* Consider Algorithm 1a, We are going to prove that at line number (8) the alg always finds a subset  $S$  that satisfies the equality. Assume not. On one hand, the number of possible values that  $m_S$  can have is  $2^i - 1$ . On the other hand, since  $J$  contains  $i + 1$  vectors on the  $i$ th iteration, it follows that the number of subsets is  $2^{i+1} - 1 \geq 2^i$ .

Therefore, there must be at least two different subsets  $S$  and  $S'$  such that  $u_S = u_{S'}$ . However, this means that

$$\begin{aligned}
m_{S \Delta S', j} &= u_j \sum_{w \in S \Delta S'} w = u_j \left( \sum_{w \in S \Delta S'} w + 2 \sum_{w \in S \cap S'} w \right) \\
&= m_{S,j} + m_{S',j} = 0
\end{aligned}$$

Thus,  $m_{S \Delta S'} = 0$ . Additionally, it is clear that the rank does not decrease, as for  $u_i$ , there exists one  $v_j$  such that only  $u_i$  is supported by  $v_j$ .  $\square$

**Claim 8.2.** Let  $v_1, v_2, \dots, v_k$  vectors in  $\mathbb{F}_2^n$  and  $m$  be an integer  $m < k$ , then there are  $u_1, u_2, \dots, u_{k'}$  for  $k' > k/2 - m$ . Such  $\text{span}\{u_1, u_2, \dots, u_{k'}\} \subset \text{span}\{v_{m+1}, v_{m+2}, \dots, v_k\}$ , for any  $i, j$  it holds that  $u_i u_j = 0$  and for any  $i \in [k']$ ,  $j \leq m$  it holds that  $u_i v_j = 0$ .

*Proof.* Modify the Algorithm 1a as follows, Initialize  $u_1, \dots, u_m$  to be  $v_1, \dots, v_m$  and  $J = \{v_{m+1}, \dots, v_{2m+2}\}$ . Notice that in the  $i$ th iteration, for the counting argument to work in the proof of Claim 8.1, we have to ensure that:

$$\begin{aligned}
|J| &\geq m + i + 1, \text{ So } m + i + 1 \leq k - m - i \\
\Rightarrow i &\leq k/2 - m - \frac{1}{2}
\end{aligned}$$

In the end,  $u_{m+1}, u_{m+2}, \dots, u_{k'}$  will satisfy the equations.  $\square$

**Claim 8.3.** Let  $v_1, v_2, \dots, v_k$  vectors in  $\mathbb{F}_2^n$ , then there are  $u_1, u_2, \dots, u_{k'}$  for  $k' > k/4$ . Such  $\text{span}\{u_1, u_2, \dots, u_{k'}\} \subset \text{span}\{v_1, v_2, \dots, v_k\}$ . And for any  $i, j$   $\sum u_{i,k} u_{j,k} = 4 \cdot 0$ .

*Proof.* Use the Algorithm 1a twice. However, in the second iteration, define  $m_{S,j}$  to be the product of module 4. Note that  $m_{S,j}$  must be either  $4n$  or  $4n + 2$ . Thus, we can follow the proof of Claim 8.1.  $\square$

**Claim 8.4.** [COMMENT] Complete for the above the version, which handle triples. number of options is  $(2^i)^2 = 2^{2i}$  and therefore we have the correctness if  $|J| > 2i + 1$ .

**Claim 8.5.** Consider the Left-Right  $(\Delta, n)$ -Complex  $\Gamma$ .  $\dim C_X / C_Z^\perp \cap C_Z / C_X^\perp$  is linear in  $n$ .

*Proof.* The rates of both  $C_X / C_Z^\perp$  and  $C_Z / C_X^\perp$  are  $(2\rho - 1)^2$ , where  $\rho$  can be any number in the range  $(0, 1)$  [LZ22]. Consider choosing  $\rho$  such that the rates of the quotient spaces are strictly greater than  $\frac{1}{2} + \alpha$ . This implies that the rate of their intersection is greater than  $2\alpha$ .  $\square$

**Corollary 8.1.** Fix the rate of the small codes  $C_A$  and  $C_B$  to  $\rho = \frac{1}{2} + \alpha$ . There is a subspace  $\Lambda \subset C_X/C_Z^\perp$  at rate  $\frac{1}{4} \cdot 2\alpha$  such that for any  $x \in \Lambda$  and  $y, z \in C_Z^\perp \cup \Lambda$  it holds that:

1.  $xy =_4 0$
2.  $xyz =_4 \sum_i x_i y_i z_i =_4 0$

**Claim 8.6.** Consider  $C, \Lambda$  and  $C', \Lambda'$  defined in ?? . Denote by  $\bar{\Lambda}$  the subspace  $C/\Lambda$ . Then:

$$d(C'/\bar{\Lambda}') \geq d(C/\bar{\Lambda})$$

*Proof.* The way we perform Guess elimination is critical. We want to make sure that we do not add an  $\Lambda$  row to a  $\bar{\Lambda}$  row. **[COMMENT]** Continue, Easy. Just need to perform the row reduction when rows of  $\Lambda$  at bottom, and then rotate the matrix  $\curvearrowright$

$$\begin{bmatrix} A & B \\ C & D \end{bmatrix} \curvearrowright \begin{bmatrix} D & C \\ B & A \end{bmatrix}$$

□

**Claim 8.7** (Not Formal). It is easy to see that by using concatenation again, one can obtain the code  $\dim \Lambda' \leftarrow \frac{1}{2} \dim \Lambda'$ . For any  $x \in \text{gen } \Lambda'$ ,  $|x|_4 = 1$ , and for any  $x \in C'/\Lambda'$ , we have  $|x|_4 = 0$ .

*Proof.* **[COMMENT]** We will do it by iterating the generators of  $C$  after performing rows reduction to the generator matrix. Now we will concatenate the  $i$  coordinate to complete the weight of the  $i$ th row to satisfy the requirements. □

## 9 Compute $|C_Z^\perp\rangle$ In Constant Depth. **[COMMENT]** Wrong Section.

Let  $C_0$  be a  $\Delta$ -length error linear binary code,  $\Gamma$  a  $\Delta$ -regular bipartite graph, and let  $C_Z$  be the Tanner code defined by  $C_0$  and  $\Gamma$ . We are about to prove that the uniform superposition over  $C_Z^\perp$  codewords can be computed with constant probability at a depth dependent only on  $\Delta$ , in particular independent of the  $C_Z^\perp$ -length. For this, we are going to use Proposition 10 in [MN98], which states that both the encoder and the decoder of any stabilizer  $m$ -length code can be implemented by a circuit at depth  $\Theta(\log m)$  with  $\Theta(m^2)$  ancillae.

**Claim 9.1.** Let  $G$  be a  $\Delta$ -regular bipartite graph, and denote by  $C_Z^\perp$  the dual-tanner code  $\mathcal{T}(G, C_0^\perp)^\perp$ . Then there is a circuit that with constant probability computes the state  $|C_Z^\perp\rangle$  at  $\Theta(\log \Delta)$  depth, and  $\Theta(\Delta^2)n$  ancillary qubits.

*Proof.* Let  $E_v$  and  $D_v$  be the encoder and the decoder of  $C_0$  over the local view of vertex  $v$ . By [MN98] we have that both have depth  $\Theta(\log \Delta)$  and require  $\Delta^2$  ancillae. Since  $\Gamma$  is bipartite, we can decompose  $V$  into  $V^-$  and  $V^+$  such that the local views of any two vertices in  $V^\pm$  are disjoint. Therefore, for any two different vertices  $v, u \in V^\pm$ , the encoders  $E_v$  and  $E_u$  act on disjoint subsets of qubits, each corresponding to the local view of either  $v$  or  $u$ . Consider the following algorithm:

- 1 Initialize  $2n$  qubits.
- 2 Call the left and right segments  $L$  and  $R$ .
- 3 Apply  $E_v$  in parallel on  $L$  for any  $v \in V^+$ .
- 4 Apply  $E_v$  in parallel on  $R$  for any  $v \in V^-$ .
- 5 XOR  $R$  into  $L$  by applying CNOT from the  $i$ th bit of  $R$  to the  $i$ th bit of  $L$ .
- 6 Apply  $D_v$  in parallel on  $R$  for any  $v \in V^-$ .
- 7 Apply  $H^k$  on  $L$ . And measure.
- 8 Accept if the result in  $C_Z$

**Algorithm 1:** Compute  $|C_Z^\perp\rangle$

For any  $v \in V$ , let  $|z_v\rangle$  be the superposition of codewords in  $C_0$  supported by the local view of  $v$ . Similarly, for any subset of vertices  $W \subset V$ , let  $|z_W\rangle$  be the uniform superposition over the subspace spanned by the generators supported by the vertices in  $W$ . In other words:

$$|z_W\rangle = \left| \sum_{v \in W} z_v \right\rangle$$

Using the notation, applying the encoders  $E_v, E_u$  for any pair of vertices with disjoint local view become:

$$\begin{aligned} E_v \cup E_u |0\rangle^n &= E_v |0 + z_u\rangle = E_v |0_{/u's \text{ view}}\rangle \otimes |z_u\rangle \\ &= |z_v\rangle |z_u\rangle = |z_u + z_v\rangle = |z_{\{u,v\}}\rangle \end{aligned}$$

So applying all the encoders  $E_v$  at once over the positive vertices results in:

$$(\cup_{v \in V^+} E_v) |0\rangle^n = (\cup_{v \in V^+ / v_0} E_v) |z_{v_0} + 0\rangle = |z_{V^+}\rangle$$

Thus the whole computation sum up into:

$$\begin{aligned} (\cup_{v \in V^+} E_v) \otimes (\cup_{v \in V^+} E_v) & |0\rangle^n \otimes |0\rangle^n \mapsto \\ \text{CNOT} \sum_{z \in A} \sum_{z' \in B} & |z_{V^+}\rangle |z_{V^-}\rangle \mapsto \\ I \otimes H^k \sum_{z \in A} \sum_{z' \in B} & |z + z'\rangle |z'\rangle \mapsto \\ \sum_{z \in A} \sum_{z' \in B} & |z + z'\rangle (-1)^{wz'} |w\rangle \mapsto \end{aligned}$$

So if  $w \in C_Z$  then clearly  $z'w = 0$ . The probability for that to occur is

$$\Pr[w \in C_Z] = \frac{|C_Z|}{\mathbb{F}_2^n} = 2^{(\rho-1)n}$$

□

## 10 Distillate $|\Lambda + C_Z^\perp\rangle$ Into Magic.

Let  $|f\rangle$  be a codeword in  $C_X$ , and let  $\hat{X}_g$  be the indicator that equals 1 if  $f$  has support on generator  $g$ , and 0 otherwise. Observe that applying  $T^\otimes$  on  $|f\rangle$  yields the state:

$$\begin{aligned} T^{\otimes n} |f\rangle &= T^{\otimes n} \left| \sum_g \hat{X}_g g \right\rangle = \exp \left( i\pi/4 \sum_g \hat{X}_g |g| - 2 \cdot i\pi/4 \sum_{g,h} \hat{X}_g \hat{X}_h |g \cdot h| \right. \\ &\quad \left. + 4 \cdot i\pi/4 \sum_{g,h} \hat{X}_g \hat{X}_h \hat{X}_l |g \cdot h \cdot l| - 8 \cdot i\pi/4 \cdot \text{integers} \right) |f\rangle \\ &= \exp \left( i\pi/4 \sum_g \hat{X}_g |g| - 2 \cdot \pi/4 \sum_{g,h} \hat{X}_g \hat{X}_h |g \cdot h| + 4 \cdot i\pi/4 \sum_{g,h} \hat{X}_g \hat{X}_h \hat{X}_l |g \cdot h \cdot l| \right) |f\rangle \end{aligned}$$

So in our case:

$$\begin{aligned} T^{\otimes n} |f\rangle &= \\ &= \exp \left( i\pi/4 \sum_{g \in \text{gen } \Lambda} \hat{X}_g \right. \\ &\quad \left. - 2 \cdot \pi/4 \sum_{g,h \in \text{gen } C_Z^\perp} \hat{X}_g \hat{X}_h |g \cdot h| \right. \\ &\quad \left. + 4 \cdot i\pi/4 \sum_{g,h \in \text{gen } C_Z^\perp} \hat{X}_g \hat{X}_h \hat{X}_l |g \cdot h \cdot l| \right) |f\rangle \end{aligned}$$



So eventually, we have a product of gates when non-Clifford gates are applied on only on generators of  $C_Z^\perp$ .

$$T^n |f\rangle = \prod_{g \in \text{gen } \Lambda} T_g \prod_{g,h \in \text{gen } C_Z^\perp} \{CS_{g,h}|CZ_{g,h}|I\rangle \prod_{g,h,l \in \text{gen } C_Z^\perp} \{CCZ_{g,h,l}|I\rangle |f\rangle$$

Decompose  $f = f_1 + f_2$ , where  $f_1$  is supported only on  $C_X/C_Z^\perp$  and  $f_2$  is supported only on  $C_Z^\perp$ . By using commuting relations, the above can be turned into.

$$T^n |f\rangle = \prod_{g \in \text{gen } \Lambda} T_g X_{f_1} \prod_{g,h \in \text{gen } C_Z^\perp} \{CS_{g,h}|CZ_{g,h}|I\rangle \prod_{g,h,l \in \text{gen } C_Z^\perp} \{CCZ_{g,h,l}|I\rangle |f_2\rangle$$

Denote by  $M_1, M_2$  the gates:

$$M_1 = \prod_{g \in \text{gen } \Lambda, h} \{CZ_{g,h}|I\rangle$$

$$M_2 = \prod_{g,h \in \text{gen } C_Z^\perp} \{CS_{g,h}|CZ_{g,h}|I\rangle \prod_{g,h,l \in \text{gen } C_Z^\perp} \{CCZ_{g,h,l}|I\rangle$$

And then we get that

$$\prod_{g \in \text{gen } \Lambda} T_g |f\rangle = M_1^\dagger T^n M_2^\dagger |f\rangle$$

$$\prod_{g \in \text{gen } \Lambda} T_g |f\rangle = M_1^\dagger T^n E L[M_2^\dagger] |L[f]\rangle$$

**Claim 10.1.** Let  $v \in V^-$ , and let  $g_1$  be the generator supported by  $v$ , which matches an assignment of a codeword in  $C_A \otimes C_B$  on the local view of  $v$ . Denote by  $U_{v,g_1}$  the control-gate which, depending on the control bit  $(v, 1)$ , turns on  $g_1$  over the edges associated with the local view of  $v$  in the graph  $G$ . Then, the depth of  $U_{v,g_1}$  depend only on  $\Delta$ .

**Claim 10.2.** Let  $(v, g_1)$  and  $(u, g_2)$  be control wires for two different generators in the graph  $G$ . Then  $U_{v,g_1}$  and  $U_{u,g_2}$  **[COMMENT]** There must be a claim about the relationship between two different generators intersection, But I don't sure exactly why.

**Definition 10.1.** We say that a quantum circuit  $\mathcal{C}$  is well error spreading if the light cone define by any  $T$ .

**Claim 10.3.** The state:

$$\sum_{z \in C_Z^\perp} \exp\left(-2 \cdot \pi/4 \sum_{g,h \in \text{gen } C_Z^\perp} \hat{X}_g \hat{X}_h |g \cdot h| + 4 \cdot i\pi/4 \sum_{g,h \in \text{gen } C_Z^\perp} \hat{X}_g \hat{X}_h \hat{X}_l |g \cdot h \cdot l|\right) |z\rangle$$

Can be computed such that any

*Proof.* Denote by  $U_v$  the gate which turn on all the generators supported on  $v$ . As any of them is just of a code word of  $C_A \otimes C_B$ , namely turning on generator require touching at most constant number of qubits combing  $\square$

**Claim 10.4.** *The state  $(M_2^\dagger \otimes I) |C_Z^\perp + \Lambda\rangle |0\rangle$  can be computed, such that the light cone depth of any non-clifford gate is bounded by constant.*

*Proof.*

$$\begin{aligned}
(I \otimes H_X) CX_{n \rightarrow n} (E \otimes E) I \otimes L[M_2^\dagger] & \prod_{\substack{J \in \{\text{gen } \Lambda, \text{gen } C_Z^\perp\} \\ g \in J}} \prod (I + X_{L[g]}) & |0\rangle |0\rangle \\
= (I \otimes H_X) CX_{n \rightarrow n} \sum_{\substack{z \in C_Z^\perp \\ x \in \Lambda}} e^{\varphi(z)} & |x\rangle |z\rangle \\
= \sum_{\substack{z \in C_Z^\perp \\ x \in \Lambda}} e^{\varphi(z)} & |x+z\rangle |0\rangle \\
= \sum_{\substack{z \in C_Z^\perp \\ x \in \Lambda}} (M_2^\dagger \otimes I) & |x+z\rangle |0\rangle \\
= (M_2^\dagger \otimes I) & |C_Z^\perp + \Lambda\rangle |0\rangle
\end{aligned}$$

□

Denote by  $p \in [0, 1]$  the error rate of input magic states, and let  $|A\rangle$  be an ancilla initialized to a one-qubit magic state. This  $|A\rangle$  can be used to compute the  $T$  gate, with a probability of  $Z$  error occurring with a probability of  $p$  [BH12].

**Claim 10.5.** *There are constant numbers  $\zeta_\Delta, \xi_\Delta$ , and a circuit  $\mathcal{C}$  such that:*

1. *In the no-noise setting, The circuit compute the state*

$$\mathcal{C} |0\rangle^{\Theta(n)} \otimes |A\rangle^{\Theta(n)} \rightarrow \prod_{g \in \text{gen } \Lambda} T_g |C_Z^\perp + \Lambda\rangle$$

2. *Otherwise, the circuit computes the state*

$$\mathcal{C} |0\rangle^{\Theta(n)} \otimes |A\rangle^{\Theta(n)} \rightarrow Z^e \prod_{g \in \text{gen } \Lambda} T_g |C_Z^\perp + \Lambda\rangle$$

, where the probability that  $e_i = 1$  is less than  $\zeta_\Delta \cdot p$ . Additionally, for any  $i$ , there are at most  $\xi_\Delta$  indices  $j$  such that  $e_i$  and  $e_j$  are dependent.

*Proof.* Concatenate the  $T^n \otimes I$  with the gate in Claim 10.4. □

**Claim 10.6.** *For any  $\alpha \in (0, 1)$  the probability that  $|e| > (1 + \alpha)p\zeta_\Delta$  is less than:*

$$\Pr[|e| > (1 + \alpha)\mathbf{E}[|e|]] < \frac{1 \cdot \xi_\Delta n}{\alpha^2 \zeta_\Delta^2 p^2 n^2} = o(1/n)$$

*Proof.* By the Chebyshev inequality, notice that the number for which  $\mathbf{E}[e_i e_j] - \mathbf{E}[e_i] \mathbf{E}[e_j] \neq 0$  is less than  $\xi_\Delta n$ . □

**Definition 10.2.** *We will said that a decoder  $\mathcal{D}$  for the good qunatum LDPC code is an good-local decoder if*

1. *There is a treashold  $\mu n$  such that if the error size is less than  $|e| < \mu n$  then  $\mathcal{D}$  correct  $e$  in constant number of rounds. With probability  $1 - o(1/n)$ .*
2. *In any rounds  $\mathcal{D}$  performs at most  $O(n)$  work (depth  $\times$  width).*

3. The above is true in operation-noisy settings, where there is a probability of  $p$  for an error to occur after acting on a qubit. (★)

★ The motivation for this is that if the decoder does not act on the qubit, then it also does not apply a  $T$  gate on it. Therefore, in the distillation setting, there is zero chance for an error to occur.

**Claim 10.7.** Suppose there is a good local decoder  $\mathcal{D}$  for the good qLDPC code. Then, there exists  $p_0$  such that for any sufficiently large  $n$ , there is a distillation protocol that, given  $\Theta(n)$  magic states at an error rate  $p < p_0$ , successfully distills  $\Theta(n)$  perfect magic states with a probability of  $1 - o(1/n)$ . Furthermore, the protocol's space and time complexity (both quantum and classical) are  $\Theta(n)$  and  $\Theta(n^2)$ , respectively.

## References

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