

# Magic States Distillation Using Quantum LDPC Codes.

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## 1 Good Codes With Large $\Lambda$ .

**Definition 1.1.** Let  $M \in \mathbb{F}_2^{k \times n}$  upper triangular matrix such that  $k < n$ . We say that  $M$  has the 1-stairs property if  $M_{ij} = 1$  any  $j < i$ .

**Claim 1.1.** Any  $M \in \mathbb{F}_2^{k \times n}$  upper triangular matrix can be turn into upper triangular matrix that has the 1-stairs property by elementary operation.

$$\begin{bmatrix} 1 & 1 & 1 & 1 & 1 & \cdot & \cdot & \cdot & \cdot & \cdot \\ 0 & 1 & 1 & 1 & 1 & \cdot & \cdot & \cdot & \cdot & \cdot \\ 0 & 0 & 1 & 1 & 1 & \cdot & \cdot & \cdot & \cdot & \cdot \\ 0 & 0 & 0 & 1 & 1 & \cdot & \cdot & \cdot & \cdot & \cdot \\ 0 & 0 & 0 & 0 & 1 & \cdot & \cdot & \cdot & \cdot & \cdot \end{bmatrix}$$

*Proof.* Consider the following algorithm: Let  $M$  be our initial matrix. We iterate over the rows from left to right. In the  $i$ th iteration, we check for any row  $j < i$  if  $M_{ji} = 1$ . If not, we set  $M$  to be the matrix obtained by adding the  $i$ th row to the  $j$ th row. Since  $M$  is an upper triangular matrix, adding the  $i$ th row does not change any entry  $M_{js}$  for  $s < i$ . Therefore, the obtained matrix is still an upper triangular matrix and the entries at  $M_{js}$  for  $j, s < i$  remain the same, namely 1 if and only if  $j \leq s$ .

Continuing with the process eventually yields, after  $k$  iterations, a matrix with the 1-stair property.  $\square$

**Claim 1.2.** Let  $C$  be a  $[n, k, d]$  binary linear code, and let  $\Lambda$  be subcode  $\Lambda \subset C$  at dimension  $k' > \alpha k$  for some  $\alpha \in (0, 1)$ . Then there exists a code  $C' = [\leq 2n, \geq (1 - \alpha + \frac{\alpha^3}{24})k, d]$  and a subcode of it  $\Lambda'$  in it at dimension  $\geq \frac{\alpha^3}{24}k$ , such:

1. For every  $x \in \Lambda'$  and  $y \in C'$   $x \cdot y = 0$
2. For every  $x \in \Lambda'$  and  $y, z \in C'$   $x \cdot y \cdot z = 0$

*Proof.* First, we can assume that the generator matrix of  $C$  is an upper triangular matrix, such that the first  $k'$  rows span  $\Lambda$ . Notice that after applying the algorithm from Claim 1.1 starting from the first row and stopping at the  $k'$ th row, the first  $k'$  rows are kept in  $\Lambda$ . So let's assume that is the form of the generator matrix.

Now, let's consider the following process: going uphill, from right to left, starting at the  $k'$  row. Initially, set  $j \leftarrow k'$  and in each iteration, advance it to be the index of the next row, namely  $j \leftarrow j - 1$ . In each iteration, ask how many rows  $G_m$ , such that  $m \leq j$ , satisfy  $G_m G_j = 0$  and how many pairs of rows  $G_m, G_{m'}$  such that  $m, m' \leq j$  satisfy  $G_m \cdot G_{m'} \cdot G_j = 0$ . Denote by  $p$  the probability to fall on unsatisfied equation from the above.

- If  $p \geq \frac{1}{2}$  then we move on to the next iteration.
- Otherwise, we encode the  $j$ th coordinate by  $C_0$ , which maps  $1 \rightarrow w$  such that  $w \cdot w = 0$ . This flips the value of  $G_m G_j$  for any pair and  $G_m G_{m'} G_j$  for any triple such that  $m, m' \leq j$ , so we get that the majority of the equations are satisfied. Also notice that the concatenation doesn't change the value of any multiplication at the form  $G_m G_{j'}$  for  $j' > j$ . Therefore, for any  $j < j' \leq k'$  the number of the satisfied equations relative to  $j'$  is not changed, meaning it is still the majority.

Set  $G$  to be the new matrix after the concatenation by  $C_0$ .

In the end of the process  $G$  is going to be the generator matrix of  $C'$ . It's left to construct  $\Lambda'$ , we are going to do so by taking from the  $k'$  rows a subset that satisfies the desired property in Claim 1.2.

Let  $S$  be the set of rows among the first  $k'$  rows for which there is at least one unsatisfied equation. We will now prove that if  $k'$  is large enough, specifically linear in  $k$ , then  $|S|$  is small enough to obtain  $\Lambda'$  by removing the rows in  $S$ .

Observe that the number of satisfied equations is at least:

$$\begin{aligned} & \frac{1}{2} (k' - 1 + (k' - 1)^2) + \frac{1}{2} (k' - 1 + (k' - 1)^2) + \frac{1}{2} (k' - 2 + (k' - 2)^2) + \dots + \frac{1}{2} (1 + (1)^2) \\ &= \frac{1}{2} \left( \binom{k' + 1}{2} + \frac{k'(k' + 1)(2k' + 1)}{6} \right) \end{aligned}$$

So

$$\begin{aligned} |S| \cdot k + |S| \cdot k^2 &\leq k' (k + k^2) - \frac{1}{2} \left( \binom{k' + 1}{2} + \frac{k'(k' + 1)(2k' + 1)}{6} \right) \\ \Rightarrow |S| &< k' - \frac{1}{2} \left( \frac{1}{k^2 + k} \binom{k' + 1}{2} + \frac{1}{k^2 + k} \frac{k'(k' + 1)(2k' + 1)}{6} \right) \\ \Rightarrow |S| &< k' - \frac{k'^3}{24k^2} \end{aligned}$$

Therefore, if  $k' \geq \alpha k$  we have that  $|S| < (\alpha - \frac{\alpha^3}{24})k$  implies that  $\dim \Lambda' \geq \frac{\alpha^3}{24}k$ .

□

**Claim 1.3.** Consider  $C, \Lambda$  and  $C', \Lambda'$  defined in Claim 1.2. Denote by  $\bar{\Lambda}$  the subspace  $C/\Lambda$ . Then

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

**Claim 1.4** (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  $x \in \Lambda'$ ,  $|x|_4 = 1$ , and for any  $x \in C'/\Lambda'$ , we have  $|x|_4 = 0$ .

**[COMMENT]** The argument above that the distance  $d'$  remain the same is not correct. Yet, if we are defining the distance of any codeword in  $C/(C/\Lambda)$  to be greater than  $d'$  then we win. (The problem was that gauss elimination might change the weight of rows associate with  $\Lambda$  generators.

## 2 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  $\hat{X}_g$ , and 0 otherwise. Observe that applying  $T^\otimes$  on  $|f\rangle$  yields the state:

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

So in our case:

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

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

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

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.

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

Denote by  $M_1, M_2$  the gates:

$$\begin{aligned}
M_1 &= \prod_{g \in \Lambda, h} \{CZ_{g,h}|I\rangle \\
M_2 &= \prod_{g, h \in C_Z^\perp} \{CS_{g,h}|CZ_{g,h}|I\rangle \prod_{g, h, l \in C_Z^\perp} \{CCZ_{g,h,l}|I\rangle
\end{aligned}$$

And then we get that

$$\begin{aligned}
\prod_{g \in \Lambda} T_g |f\rangle &= M_1^\dagger T^n M_2^\dagger |f\rangle \\
\prod_{g \in \Lambda} T_g |f\rangle &= M_1^\dagger T^n E_{L[M_2^\dagger]} |L[f]\rangle
\end{aligned}$$

**Claim 2.1.** *The state  $(M_2^\dagger \otimes I) |C_Z^\perp + \text{span } \Lambda\rangle |0\rangle$  can be computed, such that the light cone depth of any non-clifford gate is bounded by constant.*

*Proof.*

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

□

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 2.2.** *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*

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

2. *Otherwise, the circuit computes the state*

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

, 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 2.1. □

**Claim 2.3.** *For any  $\alpha \in (0, 1)$  the probability that  $|e| > (1 + \alpha)p\zeta_\Delta$  is less than:*

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

*Proof.* By the Chebyshev bound, 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 2.1.** *We will said that a decoder  $\mathcal{D}$  for the good quantum 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 2.4.** *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

- [BH12] Sergey Bravyi and Jeongwan Haah. “Magic-state distillation with low overhead”. In: *Physical Review A* 86.5 (2012), p. 052329.